

# Adaptively Secure MPC with Sublinear Communication Complexity

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## Abstract

A central challenge in the study of MPC is to balance between security guarantees, hardness assumptions, and resources required for the protocol. In this work, we study the cost of tolerating adaptive corruptions in MPC protocols under various corruption thresholds.

In the strongest setting, we consider adaptive corruptions of an arbitrary number of parties (potentially all) and achieve the following results:

- A two-round secure function evaluation (SFE) protocol in the CRS model, assuming  $\text{LWE}$  and  $\text{iO}$  for circuits. The communication, the CRS size, and the online-computation are independent of the *size* of the function. The  $\text{iO}$  assumption can be replaced by secure erasures. Previous results required either the communication or the CRS size to be polynomial in the function size.
- Under the same assumptions, we construct a “Bob-optimized” 2PC (where Alice talks first, Bob second, and Alice learns the output). That is, the communication complexity and total computation of Bob are sublinear in the function size and in Alice’s input size. We prove impossibility of “Alice-optimized” protocols.
- Assuming  $\text{LWE}$ , we bootstrap adaptively secure  $\text{NIZK}$  arguments to achieve proof size only depending on the witness size, and independent of the  $\text{NP}$ -relation.

On a technical level, our results are based on *laconic function evaluation (LFE)* (Quach, Wee, and Wichs, FOCS’18) and shed light on an interesting duality between  $\text{LFE}$  and  $\text{FHE}$ .

Next, we analyze adaptive corruptions of all-but-one of the parties, and show a two-round SFE protocol in the threshold PKI model, assuming  $\text{LWE}$  and  $\text{NIZK}$ , with communication complexity only depending on the input and output of the parties. Finally, we consider the honest-majority setting, and show a two-round SFE protocol with guaranteed output delivery under the same constraints.

Our results highlight that the asymptotic cost of adaptive security can be reduced to be comparable to, and in many settings almost match, that of static security, with only a little sacrifice to the concrete round complexity and asymptotic communication complexity.

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# 1 Introduction

After establishing feasibility in the 1980’s [91, 56, 11, 29, 85], the rich literature of multi-party computation (MPC) has focused on several performance aspects of the problem. These aspects include (a) studying the resources required in terms of communication rounds, total amount of communication, and total amount of computation, (b) minimizing the complexity assumptions that are required under the various notions, and most importantly, (c) enhancing the notion of security, starting from the simplest notion of static corruptions with semi-honest adversaries in a stand-alone model, to sequential, and concurrent composition, to adaptive corruption of parties with a fully malicious adversary.

Recent results have considered a few of these questions simultaneously. Despite several decades of progress, many basic questions about feasibility and asymptotic optimality of MPC protocols remain. The focus of this paper is to study *the price of adaptive security* in light of recent round-optimal and low-communication protocols for the static security setting.

Recall that *adaptive security* [10, 21] for an MPC protocol models the realistic threat in which the adversary can corrupt a party during the execution of a protocol—in particular, after seeing some of the transcript of a protocol. In contrast, with *static corruptions*, the adversary must choose which parties to corrupt *before* the protocol begins. In this simpler static case, the security argument relies on the fact that the inputs of the corrupted parties are known, and thus the simulator can “work around” these parties to generate a reasonable, and consistent transcript for the remaining parties. Indeed, adaptive security is known to be strictly stronger than static security [21, 23].

While the idea of allowing an adversary to corrupt parties at anytime during protocol executions seems natural, its technical formulation is captured by obliging the simulator in the security definition to support some specific tasks. In particular, the technical difficulty in achieving adaptive security is that the simulator must produce a transcript for the execution *before* knowing which parties are corrupted. In an extreme case, the protocol can already be completed, and the adversary can then *begin* to corrupt all of the parties, one by one.

Two main models are considered for adaptive corruptions. In the first and simpler one, it is assumed that parties can *securely erase* certain parts (and even all) of their random tapes.<sup>1</sup> In this setting, when simulating a party who gets corrupted, the simulator may not be required to provide random coins explaining all the messages previously sent by that party. In the second, *erasures-free* model, there are no assumptions about the ability to erase local information. When a party is corrupted in this adaptive security notion, the adversary can learn all of that party’s inputs and internal random coins. In this case, a secure protocol requires a simulator that, after producing the transcript, can “explain” the transcript by generating the coins and inputs for a given party after they are corrupted. In particular, the simulator only learns the input of that party after the corruption (e.g., after the entire execution), and then must “explain” the transcript it produced beforehand in a way that is consistent with the given input.

As a result of this difficulty, most of the literature shows that achieving adaptive security is notoriously harder than achieving static security; in some cases there are outright impossibility results such as the case of fully homomorphic encryption [74], public-key encryption which cannot exist for arbitrary messages [80], constant-round MPC in the plain model (under black-box simulation) [48],

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<sup>1</sup>We note that in certain cases it is reasonable to erase the random coins, e.g., when encrypting a message it is normally fine not to store the encryption randomness; however, in some cases one cannot erase all of its random tape, e.g., when sending a public encryption key it is normally essential to store the decryption key. We refer the reader to [21, 19] for further discussion on secure erasures.

MPC protocols with non-expander communication graphs [18], and UC-secure broadcast protocols without an honest majority [68]. All of these lower bounds, with the exception of [48], hold also for the weaker adaptive setting with secure erasures.

## 1.1 Full Adaptivity: Adaptive Corruptions of All the Parties

We start by considering the strongest adversary that can adaptively corrupt, and arbitrarily control, any subset of the participating parties. We will focus on the resources required for securely evaluating a function, balancing between the number of rounds, the communication complexity, and the online-computational complexity (the work performed between the first and last messages).

The feasibility of *adaptively secure* MPC was established in the seminal CLOS protocol [22] in a resoundingly strong manner in the UC framework [20]. This paper established the notion of fully adaptive security as described above, in the stronger, erasures-free setting, when the adversary can corrupt *all* protocol parties after execution. They then achieved this notion with a brilliant, yet complicated protocol that worked in the *common random string* model.<sup>2</sup> However, that protocol’s round complexity depended on the circuit *depth*, and its communication was polynomially larger than the *size* of the circuit being computed. Roughly 15 years later, Canetti et al. [28] constructed a constant-round protocol under standard assumptions, and recently Benhamouda et al. [13] constructed a 2-round protocol assuming 2-round adaptively secure oblivious transfer (OT). But again, both of these recent results require communication that is larger than the circuit size, and thus come at a larger cost than recent protocols for static corruptions that require two rounds with communication that is only proportional to the size of the inputs and outputs [78].

Another line of recent work overcomes the communication bottleneck, but at the cost of stronger assumptions and a large *common reference string*. Constant-round [38, 25] and 2-round [47, 27] protocols for adaptively secure MPC are known assuming indistinguishability obfuscation (iO) for circuits and one-way functions (OWF). These protocols have sublinear communication ([38, 25] in the semi-honest model, [47, 27] in the malicious setting<sup>3</sup>), but require a large CRS (at least linear in the circuit size). In particular, the approach of these results is to place an obfuscated universal circuit into the common reference string which can compute any function of a given size. Thus, these results are more aptly described as *bounded-circuit-size* adaptively secure MPC. In contrast, we aim to study a setup model in which the reference string is smaller (preferably independent) of the size of the evaluated function.

Lastly, recent advances in the static setting [84, 3] presented protocols with online-computation that only depends on the function’s *depth* but not on its *size*. In the adaptive setting, on the contrary, all known protocols require computing the function during the online part of the computation.

We now present three results in the fully adaptive setting: in Section 1.1.1, a resource-efficient MPC protocol; in Section 1.1.2, feasibility and infeasibility results regarding one-sided-optimized two-party protocols; and in Section 1.1.3, NIZK protocols with a compact proof.

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<sup>2</sup>In the *common random string* model, all parties receive a uniformly random string generated in a trusted setup phase. In the *common reference string* model, the common string is sampled according to some pre-defined distribution.

<sup>3</sup>The protocols in [38, 25] use the CLOS compiler [22] to get malicious security. Since the communication of previously known adaptively secure ZK protocols depends on the NP relation (see [76, 62, 44] and references therein), the communication of the maliciously secure protocols depended on the CRS. Our results in Section 3 can be used to reduce the communication of [38, 25] in the malicious setting as well.

### 1.1.1 Two-Round MPC with Low Communication and Online-Computation

Thus, the first major result of this paper is to present the first 2-round fully adaptively secure MPC that requires only sublinear communication (i.e., depends only on the inputs, outputs, and depth of the function), sublinear online-computation, and that uses a sublinear common reference string. Our result requires carefully combining the techniques from the recent work on *Laconic Function Evaluation* (LFE) [84] and *explainability compilers* [38]. In this sense, our answer to the main question regarding the cost of adaptive security versus static security shows a minimal cost to the communication complexity in the secure-erasures model, and the addition of complexity assumptions in the erasures-free setting: namely the need for iO in order to implement the explainability compiler. In addition, in the latter setting there is a price to pay in the size of the CRS, which depends on the depth of the function. Table 1 summarizes the performance characteristics of prior work in comparison to our new result.

To explain the key bottleneck in achieving our result, note that almost all known methods for succinct MPC in the static setting rely on *fully homomorphic encryption* [51].<sup>4</sup> The general template is for parties to encrypt and broadcast their inputs, independently evaluate the function on said inputs, and then jointly decrypt the output. The problem in the case of adaptive security is that the simulator must produce a transcript for such a protocol, consisting of the input ciphertexts and the output ciphertext, without knowing the inputs of any parties; later after corruption, the simulator would need to provide a decryption key that explains the ciphertexts for any given input and for the final output. Unfortunately, Katz et al. [74] showed that this exact task is not possible for all functions, since the existence of such a simulator would imply a compact circuit that can be used to compute the function.

To get around the impossibility of adaptively secure FHE, the key insight of our approach is to instead use a recent technique of laconic function evaluation (LFE) [84], itself an extension of the idea of laconic OT [30]. At a high level, LFE allows a party to publish a short *digest* of a function; later any party can encrypt an input to that function such that the resulting ciphertext is still small with respect to the *size* of the function. In particular, both the digest and the ciphertext size are proportional to the *depth* of the function. Because the computational cost of the decryption algorithm is proportional to evaluating the function, LFE avoids the impossibility argument for adaptive security from [74], while preserving the succinct communication pattern. Quach et al. [84] instantiated LFE under a natural variant of the learning with errors assumption, called *adaptive LWE* (*ALWE*).<sup>5</sup> LFE is in some sense a dual notion to FHE. We extend on this duality in the discussion on the two-party case below.

Our starting point follows the statically secure protocol from [84]. The idea is for the parties to each locally compute a digest of the function  $f$  (this is done deterministically, using a CRS for LFE parameters), and then use an adaptively secure MPC protocol (possibly not communication efficient) to jointly compute the encryption of the inputs  $(x_1, \dots, x_n)$ . The communication and online-computation required are naturally proportional only to the encryption algorithm, which depends on the depth of the original function but not on its size. Finally, each of the parties can then locally decrypt the ciphertext with respect to the digest to recover the output.

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<sup>4</sup>Another approach for compact MPC is using *functional secret sharing* (*FSS*) [16, 17]. This approach does not seem to support adaptive corruptions.

<sup>5</sup>The basic construction in [84] holds under the standard LWE assumption, however, for the purpose of (semi-)malicious MPC, in which the inputs to the protocol can be chosen adaptively, after the CRS is published, we require the stronger variant.

Nonetheless, for adaptive security, it is unclear how to simulate the output ciphertext when possibly all  $n$  parties can be corrupted. To circumvent this barrier, we first show fully adaptive security of this protocol in the *erasures* model, without any additional assumptions, and then carefully remove the erasures using the explainability compiler technique from [38].

Overall, this framework achieves all of the round, communication, and online-computation complexity goals, but in the erasures-free setting, it still requires a common reference string whose size is related to the *depth* of the function being computed, and further it relies on iO. In contrast, in the static corruption setting, only LWE is required.

Protocol	Security (erasures)	Rounds	Communication	Online Computation	Setup size	Setup type	Assumption
MW [78]	static	2	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}(\kappa)$	CRS	LWE, NIZK
QWW [84] ABJMS [3]	static	2	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, d, \kappa, n)$	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, d, \kappa, n)$	$\text{poly}(\kappa)$	CRS	ALWE LWE
CLOS [22]	adaptive(no)	$O(d)$	$ C  \cdot \text{poly}(\kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}(\kappa)$	CRS	TDP, NCE dense-crypto
GS [48]*	adaptive(no)	$O(d)$	$ C  \cdot \text{poly}(\kappa, n)$	$\text{poly}( C , \kappa)$	-	-	CRH TDP, NCE dense-crypto
DKR [38] CGP [25]	adaptive(no)	$O(1)$	$ C  \cdot \text{poly}(\kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}( C , \kappa)$	Ref	OWF, iO
GP [47]	adaptive(no)	2	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}( C , \kappa)$	Ref	OWF, iO
CPV [28]	adaptive(no)	$O(1)$	$ C  \cdot \text{poly}(\kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}(\kappa)$	CRS	NCE dense-crypto
BLPV [13]	adaptive(no)	2	$ C  \cdot \text{poly}(\kappa, n)$	$\text{poly}( C , \kappa)$	$\text{poly}(\kappa)$	Ref	adaptive 2-round OT
This work	adaptive(yes)	2	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, d, \kappa, n)$	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, d, \kappa, n)$	$\text{poly}(\kappa)$	CRS	ALWE
	adaptive(no)					Ref	ALWE, iO

Table 1: Round, communication, and online-computation of MPC tolerating any number of corruptions, for  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  represented by a circuit  $C$  of depth  $d$ . *CRS* refers to a common random string, whereas *Ref* refers to a common *reference* string whose sampling coins are secret. (\*) The results in [48] only hold in the stand-alone setting.

### 1.1.2 Alice/Bob-Optimized protocols.

Consider a two-message protocol for two parties, where Alice sends the first message, Bob replies with the second, and only Alice learns the output. In this setting, it is possible for one party’s total *computation* (and thus also total communication) to be proportional to the size of their input and output, while the other party “does all of the work” of securely evaluating the function. These protocol variants are designated as “optimized for Alice” or “optimized for Bob,” depending on which party saves the work.

In the static-corruption setting, Alice-optimized protocols can be constructed assuming FHE, where Alice encrypts her input, Bob homomorphically evaluates the circuit and returns the encrypted result. Quach et al. [84] showed that Bob-optimized protocols can be constructed from LFE, where Alice compresses the function with her input hard-wired, sends the digest to Bob who replies with the encryption of his input. Therefore, in the static setting, FHE and LFE are dual notions with respect to the work-load of the computation. We next show that in the adaptive

setting this duality breaks. On the one hand, we extend the impossibility result of FHE [74] to rule out adaptively secure 2-round Alice-optimized protocols (even assuming secure erasures). On the other hand, we construct an adaptively secure, semi-malicious,<sup>6</sup> Bob-optimized protocol from LFE and iO (alternatively, just from LFE assuming secure erasures). We note that any 2-round Bob-optimized protocol can be converted into a 3-round Alice-optimized protocol, which is the best one could hope for. Table 2 summarizes our results vis a vis prior work.

The key idea behind our Bob-optimized protocol is to use the same LFE approach put forth in [84] for static security, and strengthen it to tolerate adaptive corruptions. To support an adaptive corruption of Alice, the simulator will need to produce an *equivocal* first message, i.e., to simulate the digest without knowing the input value of Alice, and upon a later corruption of Alice generate appropriate random coins explaining the message. Our first technical contribution is to create an equivocal version of the LFE scheme of [84]. Similarly, to support an adaptive corruption of Bob, the simulator should be able to generate an equivocal second message, i.e., generate the ciphertext without knowing the input of Bob, and upon a later corruption of Bob provide appropriate random coins. This can be handled either assuming secure erasures, or using explainability compilers.

Approach	Security (erasures)	CRS	Communication		Computation		Assumptions
			Alice	Bob	Alice	Bob	
GC [91]	static	-	$\ell_A$	$ f $	$ f $	$ f $	static OT
LOT [30]	static	$O(1)$	$O(1)$	$ f $	$ f $	$ f $	DDH, etc.
FHE [51]	static	-	$\ell_A$	$\ell_{\text{out}}$	$\ell_A + \ell_{\text{out}}$	$ f $	LWE
LFE [84]	static	$O(1)$	$O(1)$	$\ell_B + \ell_{\text{out}}$	$ f $	$\ell_B + \ell_{\text{out}}$	ALWE
equivocal GC [28]	adaptive (no)	-	$\ell_A$	$ f $	$ f $	$ f $	adaptive OT
This work	adaptive (yes)	$O(1)$	$O(1)$	$\ell_B + \ell_{\text{out}}$	$ f $	$\ell_B + \ell_{\text{out}}$	ALWE
	adaptive (no)	$\ell_B$	$O(1)$	$\ell_B + \ell_{\text{out}}$	$ f $	$\ell_B + \ell_{\text{out}}$	ALWE and iO
	adaptive (yes)	$ f $	$ f $	$\ell_{\text{out}} + o(\ell_B)$	$ f $	$ f $	impossible

Table 2: Comparison of two-message semi-honest protocols for  $f : \{0, 1\}^{\ell_A} \times \{0, 1\}^{\ell_B} \rightarrow \{0, 1\}^{\ell_{\text{out}}}$ . Alice sends the first message, Bob sends the second, and only Alice learns the output. For simplicity, multiplicative factors that are polynomial in the security parameter  $\kappa$  or the circuit depth  $d$  are suppressed.

### 1.1.3 Succinct Adaptively Secure NIZK

Next, we consider the problem of constructing an adaptively secure non-interactive zero-knowledge protocol (NIZK) that is “succinct,” i.e., the size of the proof and of the common reference string should be smaller than the size of the circuit relation. The best we can hope for is for the proof to be the size of the witness (as otherwise, the lower-bound of Gentry and Wichs [52] requires a non-standard complexity assumption). The first adaptively secure NIZK was constructed by Groth et al. [60], however it was not succinct. Gentry [51] and later Gentry et al. [55] combined fully homomorphic encryption (FHE) with a standard NIZK system in order to construct such schemes that are secure against static corruptions, and as observed in [55] also against adaptive corruptions

<sup>6</sup>In the semi-malicious setting, the adversary follows the protocol as in the semi-honest case, but he can choose arbitrary random coins for corrupted parties.



in the secure-erasures setting. However, these schemes are not secure against adaptive corruptions in the erasure-free setting. In particular, they run into the FHE bottleneck for adaptive security by Katz et al. [74] described above.

Our main technique to overcome this lower bound is to use *homomorphic trapdoor functions (HTDF)* [57]. HTDF schemes are a primitive that conceptually unites homomorphic encryption and homomorphic signatures. In our usage, HTDF can be thought of as fully homomorphic commitment schemes which are equivocal (hence, statistically hiding), where a trapdoor can be used to open any commitment to any desired value. Using HTDF, the prover can commit to the witness (instead of encrypting it), evaluate the circuit over the commitments, and use adaptive but non-succinct NIZK (e.g., from [60]) to prove knowledge of the witness and that the result commits to 1. The verifier evaluates the circuits over the committed witness, and verifies the NIZK to ensure that the result is a commitment to 1. These succinct NIZK will turn out useful for our protocols in Section 1.2, and we believe they are of independent interest. A summary of our results in comparison to prior work appears in Table 3.

Protocol	Security (erasures)	CRS size	Proof size	Assumptions
Groth [59]	static	$ C  \cdot \text{poly}(\kappa)$	$ C  \cdot \text{poly}(\kappa)$	TDP
Groth [59]	static	$ C  \cdot \text{polylog}(\kappa) + \text{poly}(\kappa)$	$ C  \cdot \text{poly}(\kappa)$	Naccache-Stern
GOS [60]	adaptive (no)	$\text{poly}(\kappa)$	$ C  \cdot \text{poly}(\kappa)$	pairing based
Gentry [51]	adaptive (yes)	$\text{poly}(\kappa)$	$ w  \cdot \text{poly}(\kappa)$	FHE, NIZK
GGIPSS [55]	adaptive (yes)	$\text{poly}(\kappa)$	$ w  + \text{poly}(\kappa)$	FHE, NIZK
This work	adaptive (no)	$\text{poly}(\kappa)$	$ w  \cdot \text{poly}(\kappa)$	HTDF, NIZK

Table 3: NIZK proofs with security parameter  $\kappa$ , circuit size  $|C|$ , and witness size  $|w|$ .

## 1.2 Other Notions of Adaptive Security

Recall that the notion of fully adaptive security allows the adversary to corrupt *all* of the parties in the execution—in which case the protocol offers no privacy of inputs. A criticism of this notion is that it may be too strong for certain applications. In fact, the motivation behind this strong notion arises mainly from its application to *composition* of protocols. Namely, in a larger protocol that involves more parties, participants of a sub-protocol may eventually *all* become corrupted, and thus security of the larger protocol will depend on the fully adaptive security of the subprotocol.

It is equally justifiable, however, to consider other protocol-design tasks in which the protocol need only withstand a weaker adversary who can corrupt either all-but-one of the participants, or—weaker still—only a minority of the players. We next consider adaptive security in these two settings.

### 1.2.1 All-but-one Corruptions

When considering adaptive security for all-but-one corruptions, Ishai et al. [69] constructed a constant-round, information-theoretically secure protocol in the OT-hybrid model. Garg and Sahai [48] showed an elegant way to instantiate the trusted setup required for [69] using non-black-box



techniques and thus constructed a constant-round MPC protocol in the plain model, under standard cryptographic assumptions. The communication in both of these protocols is super-linear in the circuit size.

In contrast, for the weaker notion of static security, Asharov et al. [5] presented a 2-round protocol with succinct communication, albeit in the *threshold-PKI model*. The threshold-PKI model is a setup in which all of the participants in the protocol are privately given individualized key shares corresponding to a public key. A single-round protocol for threshold PKI was also given in [5], yielding a 3-round protocol in a standard CRS setup. Mukherjee and Wichs [78] removed the need for this extra round, thereby presenting a 2-round succinct MPC in the common random string model with static corruptions.

We can thus pose our main question regarding the *cost* of adaptive security for communication-optimal protocols. Recently, Damgård et al. [43] constructed an adaptively secure 3-round MPC protocol with succinct communication complexity in the threshold-PKI model assuming LWE. Their main idea is to use a special threshold FHE scheme that enables equivocating encryptions of 0 to encryptions of 1. Initially, the parties broadcast encryptions of their inputs. Next, each party locally evaluates the circuit, and the parties re-randomize the evaluated ciphertext in the second round by broadcasting (special) encryptions of 0. The third round is a single-round threshold decryption protocol.

To simulate this protocol, the simulator uses the equivocal mode of the public key. This way, all ciphertexts in the first round are simulated as encryptions of 0. After extracting corrupted parties’ inputs, and obtaining the output value, the simulator uses the re-randomizing round to carefully add non-zero encryptions, and force the joint ciphertext to be an encryption of the output. Finally, the threshold decryption protocol is simulated. We note that using the approach of [43] (which is based on [41]), the re-randomization round seems to be inherent, and so it is unclear how to obtain optimal two rounds using this technique.

Our second major result is to construct an adaptively secure 2-round MPC assuming non-committing encryption (NCE) and threshold equivocal FHE in the threshold-PKI setup model. The setup assumption can be instantiated using the recent 2-round protocol of [13], assuming 2-round adaptively secure OT, resulting in a 4-round variant in the CRS model. All of the necessary primitives can be instantiated from LWE in the semi-malicious setting, and security in the malicious case follows using the NIZK from Section 1.1.3.

Our protocol follows the template of [5], where every party encrypts its input in the first round, locally evaluates the circuit over the ciphertexts, uses its key-share to partially decrypt the result, and broadcasts the decrypted share (some additional “smudging” noise is sometimes required to protect the decryption share). The technical challenges are: (1) the ciphertexts in the first round must be created in an equivocal way, and (2) the simulation strategy used for the threshold decryption in [5] (and similarly in [78]) is inherently static, and does not translate in a straightforward way to the adaptive setting.

We overcome the first challenge by constructing a novel *threshold equivocal FHE* scheme. The scheme is equipped with an equivocal key-generation algorithm. All ciphertexts encrypted in this mode are “meaningless” and carry no information about the plaintext, and a trapdoor can be used to equivocate any ciphertext to any message. We instantiate this FHE scheme using the *dual-mode* HTDF scheme of Gorbunov et al. [57] that can generate the homomorphic trapdoor functions in an extractable mode, corresponding to the standard (meaningful) mode of the FHE, and an equivocal mode, corresponding to the meaningless mode.

We now explain the second challenge. As observed in [5, 78], the threshold decryption protocol may leak some information about the shares of the secret key, and the simulator for the decryption protocol can be used to protect *exactly* one party. In the static setting, when the set of corrupted parties is known ahead of time, the simulator can choose one of the honest parties  $P_h$  as a special party for simulating the threshold decryption. This approach does not work in the adaptive setting since the party  $P_h$  may get corrupted after simulating the decryption protocol. The simulator cannot know in advance which party will be the last to remain honest. For this reason, we use a different simulation strategy which allows the simulator to “correct” its choice of the party that is simulated as honest for the decryption protocol. Technically, this is done by having each party send shares of zero to each other party over a secure channel (that can be instantiated via NCE). These shares are used to hide the partial decryptions without changing their value. Since shares exchanged between pairs of honest parties remain hidden from the eyes of the adversary, the simulator has more freedom to replace the special party  $P_h$  upon corruption, by another honest party, even after simulating the decryption protocol.

Thus, as it stands, the cost of adaptive security with respect to the best statically secure protocols is either the threshold-PKI setup assumption, or the requirement of 2 additional rounds. Removing either of these costs remains an interesting open question. Table 4 summarizes the prior work and our contribution in this model.

Protocol	Security	Rounds	Communication	Assumptions	Setup
AJLTVW [5]	static	2 3	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI CRS
MW [78]	static	2	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	CRS
IPS [69]	adaptive	$O(1)$	$ C  + \text{poly}(d, \log  C , \kappa, n)$	OT-hybrid	-
GS [48]*	adaptive	$O(1)$	$ C  + \text{poly}(d, \log  C , \kappa, n)$	CRH, TDP, NCE dense crypto	-
DPR [43]	adaptive	3	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI
This work	adaptive	2 4	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI CRS

Table 4: Comparison of maliciously secure MPC for  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  represented by a circuit  $C$  of depth  $d$ , tolerating  $n - 1$  corruptions. (\*) The results in [48] only hold in the stand-alone setting.

### 1.2.2 Honest-Majority Setting

In the honest-majority setting, it is possible to guarantee output delivery to all honest parties. Damgård and Ishai [39] demonstrated the feasibility of constructing adaptively secure protocols that use a constant number of rounds and only require one-way functions. However, the communication of their protocol is super-linear in the circuit size.

Asharov et al. [5] constructed the first succinct protocol in this security setting using threshold FHE; their protocol requires 4 rounds in the threshold-PKI model and 5 rounds in the CRS model. However, their protocol is only secure against static corruptions. Gordon et al. [58] reduced the round complexity to 2 in the threshold-PKI model or 3 in the CRS model. Recently, Ananth et al. [4] showed a 3-round protocol in the plain model with communication proportional to the *size* of the

circuit, and Badrinarayanan et al. [6] showed a similar result with communication proportional to the *depth*. Moreover, this round complexity is tight because it is known that 2-round fair protocols are impossible in the CRS model [50, 58, 83].

Our third result is to construct an adaptively secure analogue of these latest results. In particular, we construct a 2-round adaptively secure MPC with guaranteed output delivery and the same communication complexity as in the static case, assuming NCE and threshold equivocal FHE in the threshold-PKI model (all assumptions can be based on LWE) in the semi-malicious setting. Security in the malicious case follows using the NIZK from Section 1.1.3. We can compile our 2-round protocol into a constant-round protocol in the plain with the same communication complexity using the protocol of Damgård and Ishai [39].

The 2-round protocol is based on the protocol from the all-but-one case, described in Section 1.2.1. The challenge lies in overcoming aborting parties and guarantee output delivery. We combine techniques from the threshold FHE of [58] that required  $n/2$  decryption shares to reconstruct the output into our threshold equivocal FHE. The main idea is to share the decryption key using Shamir’s secret sharing instead of additive secret sharing. Both Shamir’s reconstruction and the decryption algorithm consist of linear operations, which make them compatible with each other. As observed by Gordon et al. [58] (see also [15]), the problem with a naïve use of this technique is that the “smudging noise” (used to protect partial decryptions from leakage) is multiplied by the Lagrange coefficients, which may cause an incorrect decryption. Following [58], we have each party secret share his smudging noise using Shamir’s scheme, in a way that is compatible with the reconstruction procedure. We show that this technique can support adaptive corruptions.

To conclude, in the threshold-PKI model, the price of adaptive security is *the same* as of static security in terms of assumptions, number of rounds, and communication complexity. In the plain model, the cost is an additional constant number of rounds. Table 5 summarizes prior work and our results.

Protocol	Security	Rounds	Communication	Assumptions	Setup
AJLTVW [5]	static	4 5	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI CRS
GLS [58]	static	2 3	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI CRS
ACGJ [4]	static	3	$ C  \cdot \text{poly}(\kappa, n)$	PKE and zaps	-
BJMS [6]	static	2 3	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, d, \kappa, n)$	LWE, zaps, dense crypto	threshold PKI -
DI [39]	adaptive	$O(1)$	$ C  \cdot \text{poly}(\kappa, n)$	OWF	-
This work	adaptive	2 $O(1)$	$\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$	LWE, NIZK	threshold PKI -

Table 5: Comparison of maliciously secure MPC for  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  represented by a circuit  $C$  of depth  $d$ , in the honest-majority setting.

### 1.3 Additional Related Work

Adaptive security tolerating an arbitrary number of corruptions has been considered in various models, including protocols in the CRS model [22, 28, 13], the sunspot model [24], the key-registration

model [8], the temper-proof hardware model [65], the super-polynomial simulation model [7, 63], and more generally, based on UC-puzzles [37, 90]. All of these protocols require super-linear communication complexity.

Adaptive security in the secure-erasures model was considered in [10, 75, 70, 12, 81, 42, 64], and in the erasures-free model tolerating all-but-one corruptions in [72, 69, 61, 43] as well as in the honest-majority setting [36, 41, 39]. With the exception of [43], all of these protocols also require super-linear communication complexity.

Garay et al. [46] considered information-theoretic MPC in the client-server setting, where a constant number of clients uses  $n$  servers that assist with the computation, and studied sublinear communication in the number of *servers*. They gave a complete characterization for semi-honest security with static corruptions and adaptive corruptions with or without erasures.

In the static setting, MPC with sublinear communication complexity over eventual-delivery asynchronous channels was constructed in [33]. We conjecture that our techniques can also be applied in the asynchronous setting to obtain adaptive security with low communication.

We note that since the protocol in [47] has low communication complexity, and its CRS size depends on the circuit size, it is possible to use a more compact representation of the function, e.g., by a Turing machine (TM), and obfuscate it using iO for Turing machines; however, the solution provided in this paper is different in several qualitative aspects. First, to make the CRS independent of the computation at hand, it is preferred to obfuscate a *universal* TM, which receives the description of the concrete TM on its input tape; while iO for TM with *bounded* inputs exists under the same assumptions as iO for circuits [14, 26], iO for TM with *unbounded* inputs is only known under the stronger assumptions of public-coin differing-inputs obfuscation [71]. Second, it is not clear how to replace the iO for TM assumption by secure erasures. Third, the computation may require a large auxiliary information, e.g., access to a large database, whose description is independent of the TM; this may result with a large description of the function. In our solution, the obfuscated circuit is sublinear in the computation size even when a large auxiliary information is used.

## 1.4 Open Questions

Our main question is to study the price of adaptive security. Dramatic improvements in the answer to this question have emerged over the past 15 years, and this paper is able to establish almost zero cost in terms of round or communication. Our results, however, leave the following questions as future work.

- **Reducing setup assumptions.** Our major results for fully adaptive, 2-round, succinct protocols without erasures require a *common reference string*. Are there fully adaptive secure protocols with sublinear communication complexity in the common *random* string (even with super-constant number of rounds)?
- **Reducing hardness assumptions.** Are there fully adaptive secure protocols with sublinear communication without assuming secure erasures or explainability compilers and iO?
- **Improving setup assumptions/round complexity for all-but-one.** Our optimal-round protocol requires a pre-distribution of the FHE key. We show a 4-round protocol in the CRS model (equivalently, in the plain model for semi-honest). Are there 2 or 3 round protocols with sublinear communication in the CRS model to match results for static adversaries?

## Paper Organization

The preliminaries are presented in §2. In §3, we present our adaptively secure NIZK construction. In §4, we present our main results on fully adaptive security, and in §5.1 and §5.2, we present our results on Bob- and Alice-optimized protocols. In §6, we consider the all-but-one corruption case, and in §7, the honest-majority case. Some of the preliminaries (Appendix A), the model definition (Appendix B), and the construction of threshold equivocal FHE from LWE (Appendix C) are deferred to the appendix.

## 2 Preliminaries

**Basic notations.** In the following we introduce some notation and terminology. For  $n \in \mathbb{N}$  let  $[n] = \{1, \dots, n\}$ . We denote by  $\kappa$  the security parameter. Let  $\text{poly}$  denote the set all positive polynomials and let PPT denote a probabilistic algorithm that runs in *strictly* polynomial time. A function  $\nu: \mathbb{N} \rightarrow \mathbb{R}$  is *negligible* if  $\nu(\kappa) < 1/p(\kappa)$  for every  $p \in \text{poly}$  and sufficiently large  $\kappa$ . Given a random variable  $X$ , we write  $x \leftarrow X$  to indicate that  $x$  is selected according to  $X$ . The statistical distance between two random variables  $X$  and  $Y$  over a finite set  $\mathcal{U}$ , denoted  $\text{SD}(X, Y)$ , is defined as  $\frac{1}{2} \cdot \sum_{u \in \mathcal{U}} |\Pr[X = u] - \Pr[Y = u]|$ . Two distribution ensembles  $X = \{X(a, \kappa)\}_{a \in \{0,1\}^*, \kappa \in \mathbb{N}}$  and  $Y = \{Y(a, \kappa)\}_{a \in \{0,1\}^*, \kappa \in \mathbb{N}}$  are **computationally indistinguishable** (denoted  $X \stackrel{c}{\equiv} Y$ ) if for every non-uniform polynomial-time distinguisher  $\mathcal{A}$  there exists a function  $\nu(\kappa) = \text{negl}(\kappa)$ , such that for every  $a \in \{0, 1\}^*$  and every  $\kappa$ ,

$$|\Pr[\mathcal{A}(X(a, \kappa), 1^\kappa) = 1] - \Pr[\mathcal{A}(Y(a, \kappa), 1^\kappa) = 1]| \leq \nu(\kappa).$$

The distribution ensembles  $X$  and  $Y$  are **statistically close** (denoted  $X \stackrel{s}{\equiv} Y$ ) if for every  $a \in \{0, 1\}^*$  and every  $\kappa$  it holds that  $\text{SD}(X(a, \kappa), Y(a, \kappa)) \leq \nu(\kappa)$ . We represent elements in  $\mathbb{Z}_q$  as integers in the range  $(-q/2, q/2]$ .

**Cryptographic primitives.** In this work, we consider secure protocol in various security settings that require different cryptographic primitives. We present formal definitions for all primitives in Appendix A. An informal description of every primitive is given before it is used in the main body.

**Security model.** We present our results in the UC framework. In Appendix B, we provide a high-level description of the framework, and refer the reader to [20] for a more detailed description.

In our secure function evaluation protocols, we will consider two security notions. In the honest-majority setting, we will consider security with *guaranteed output delivery*, informally meaning that all honest parties will receive the correct output from the computation. In general, when an honest majority is not assumed this cannot be achieved [32], and the standard requirement is for *security with abort*, informally meaning that the adversary has the capability to first learn the output from the computation and later force all honest parties to output  $\perp$ .

Guaranteed output delivery and security with abort are not to be confused with *guaranteed termination*, which means that the honest parties actually finish the protocol. We emphasize that UC protocols cannot provide guaranteed termination since the adversary has full control over the communication channels, and he can simply “hang” the computation. Therefore, following the convention of Canetti et al. [22], we exclude trivial protocols, and require that the properties of guaranteed output delivery or security with abort will hold when the environment provides

sufficiently many activations to the parties, and the adversary delivers all messages.<sup>7</sup> In particular, unlike the stand-alone model, in the UC model even when a protocol guarantees output delivery, we allow the adversary to learn the output from the computation while the honest parties do not; however, if an honest party terminates it is guaranteed to receive the output. An alternative, is to consider the framework of [73] which ensures guaranteed termination regardless of the adversary’s actions (but, still, as long as the environment provides sufficiently many activations to the parties).

### 3 Adaptively Secure NIZK

Informally speaking, universally composable non-interactive zero-knowledge arguments (UC-NIZK) are single-message zero-knowledge protocols, that remain secure independently of the environment in which they are run, i.e., regardless of the protocol that is calling them. Groth et al. [60] formalized UC-NIZK as a protocol that realizes the NIZK ideal functionality, and constructed adaptively secure UC-NIZK protocols (that remain secure even if all the participants are dynamically corrupted) under hardness assumptions in bilinear groups. The size of the proof in [60] is  $|C| \cdot \text{poly}(\kappa)$ , where  $C$  is the circuit computing the NP-relation  $R$ , i.e.,  $C(x, w) = 1$  if and only if  $(x, w) \in R$ .

Gentry [51] showed how to use FHE and reduce the communication complexity to depend only on the witness size  $|w| \cdot \text{poly}(\kappa)$ . Gentry et al. [55] further reduced the proof size to  $|w| + \text{poly}(\kappa)$  and proved adaptive security of the scheme assuming secure erasures. The main underlying idea in [51, 55] is for the prover to encrypt the witness using an FHE scheme and prove knowledge of the plaintexts. In addition, the prover homomorphically evaluates the circuit  $C$  and proves that the resulting ciphertext decrypts to 1. The verifier homomorphically evaluates the circuit  $C$  over the encrypted witness, makes sure that the result matches the one sent by the prover, and validates the proof sent by the prover (that he knows the witness and the result decrypts to 1). Adaptive security is obtained by having the prover erase all of his coins before sending the proof.

We next present a new adaptively secure UC-NIZK with communication complexity that only depends on the witness size and does *not* rely on secure erasures. The main obstacles in translating the ideas from [51, 55] to the erasures-free adaptive setting are: (1) the encryption of the witness must be first simulated (without knowing the witness) and later, upon a corruption of the prover, the encryption coins must be properly explained, and (2) the impossibility result of Katz et al. [74] rule-out a standard use of FHE schemes in the protocol. We get around these barriers by using *homomorphic trapdoor functions (HTDF)* [57] (see Appendix A.2.3).

#### 3.1 The UC-NIZK Ideal Functionality

We start by presenting the NIZK ideal functionality from [60, 55]. The NIZK functionality is based on the zero-knowledge functionality from [20], and is adjusted to capture the special properties of a non-interactive proof. In particular, as opposed to interactive ZK protocols, we do not always know in advance who is going to receive the NIZK argument. Therefore, unlike the ZK functionality, there is no need to define the verifiers as part of the interface to the NIZK functionality. Furthermore, since the NIZK argument is simply a bit-string, anybody can use this string to verify the validity of the statement and can use it to convince others.

The NIZK ideal functionality (formally described in Figure 1) is parametrized by an NP-relation  $R$  and a set of parties  $P_1, \dots, P_n$ . Every party  $P_i$  can send a **prove** request consists of a pair  $(x, w)$ ,

---

<sup>7</sup>Other properties such as *privacy* and *independence of inputs* are always required to hold.



the functionality verifies that  $(x, w) \in R$ , and asks the adversary to generate a proof  $\pi$  for the statement  $x$ . The functionality stores  $(x, \pi)$  and returns the proof to  $P_i$ . Similarly, every party  $P_j$  can send a verify request, consisting of a statement  $x$  and a proof  $\pi$ . In case the pair  $(x, \pi)$  is stored, the functionality returns 1. Otherwise, the functionality asks the adversary to come up with a witness  $w$ . If  $(x, w) \in R$  the functionality returns 1 to  $P_j$ ; otherwise, return 0.

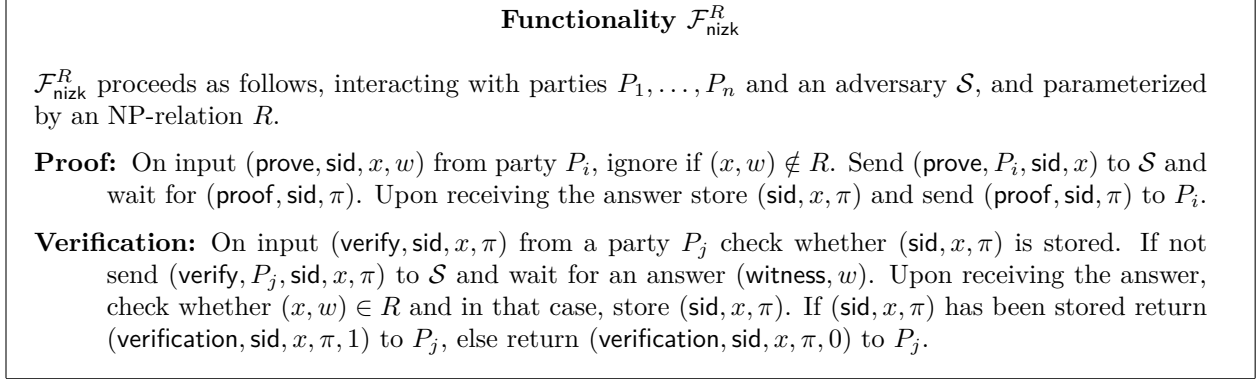


Figure 1: The ideal NIZK proof functionality

## 3.2 Cryptographic Primitives used in the Protocol

We informally describe the cryptographic primitives used in the construction. Formal definitions appear in the appendix.

### 3.2.1 Homomorphic Trapdoor Functions (HTDF)

We briefly recall the notion of homomorphic trapdoor functions, formally described in Appendix A.2.3. The key-generation algorithm outputs a public key and a secret key  $(\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d)$ . The public key along with some bit  $x \in \{0, 1\}$  define an efficiently computable function  $f_{\text{pk}, x} : \mathcal{U} \rightarrow \mathcal{V}$ , and the secret key acts as a trapdoor that enables inverting the function, i.e., defines a function  $\text{HTDF.Inv}_{\text{sk}, x} : \mathcal{V} \rightarrow \mathcal{U}$ . Two homomorphic valuation algorithms are defined given a function  $g : \{0, 1\}^\ell \rightarrow \{0, 1\}$  of depth  $d$ : the *inner* evaluation  $u^* = \text{HTDF.Eval}^{\text{in}}(g, (x_1, u_1), \dots, (x_\ell, u_\ell))$  and the *outer* evaluation  $v^* = \text{HTDF.Eval}^{\text{out}}(g, v_1, \dots, v_\ell)$ .

Informally, we require the following properties from an HTDF scheme:

- **Correctness.** Let  $x_1, \dots, x_\ell \in \{0, 1\}$  and  $v_i = f_{\text{pk}, x_i}(u_i)$  for  $i \in [\ell]$ . Then, for  $u^* = \text{HTDF.Eval}^{\text{in}}(g, (x_1, u_1), \dots, (x_\ell, u_\ell))$  and  $v^* = \text{HTDF.Eval}^{\text{out}}(g, v_1, \dots, v_\ell)$  it holds that  $f_{\text{pk}, y}(u^*) = v^*$ , where  $y = g(x_1, \dots, x_\ell)$ .
- **Distributional equivalence of inversion.** For a bit  $x \in \{0, 1\}$ , the tuple  $(\text{pk}, x, u, v)$  computed as  $v = f_{\text{pk}, x}(u)$  for a random  $u \leftarrow \mathcal{U}$  is statistically close to sampling  $v \leftarrow \mathcal{V}$  at random and computing  $u = \text{HTDF.Inv}_{\text{sk}, x}(v)$ .
- **Claw-free security.** Given the public key, no efficient adversary can come up with  $u$  and  $u'$  such that  $f_{\text{pk}, 0}(u) = f_{\text{pk}, 1}(u')$  with more than a negligible probability.

For the remaining of the section it will be useful to think of an HTDF scheme as a fully homomorphic, statistically-hiding commitment scheme with a trapdoor  $\text{sk}$ .



### 3.2.2 Strong One-Time Signatures

Strong one-time signatures are formally defined in Appendix A.2.4. Loosely speaking, the scheme consists of three algorithms ( $\text{Sig.Gen}, \text{Sign}, \text{Vrfy}$ ). The key generation algorithm outputs a public verification key and a secret signing key  $(\text{vk}, \text{sigk}) \leftarrow \text{Sig.Gen}(1^\kappa)$ . The signing algorithm uses the signing key to sign an arbitrary message  $m$  as  $\sigma \leftarrow \text{Sign}_{\text{sigk}}(m)$ . The verification algorithm uses the verification key to accept or reject the signature as  $b = \text{Vrfy}_{\text{vk}}(\sigma, m)$ . We require that the scheme be correct and strongly existentially unforgeable under a single chosen message attack, meaning that except for negligible probability, no efficient adversary that chooses a message  $m$  and receives a signature  $\sigma$  can come up with  $(m', \sigma') \neq (m, \sigma)$  such that  $1 = \text{Vrfy}_{\text{vk}}(\sigma', m')$ .

### 3.3 The UC-NIZK Protocol

We proceed to present a non-interactive protocol that securely realizes  $\mathcal{F}_{\text{nizk}}^R$  for an arbitrary NP-relation  $R$ . Denote by  $d$  the depth of the circuit computing the relation. A proof for a statement  $x$  with witnesses of size  $|w|$  consists of  $|w| \cdot \text{poly}(\kappa)$  bits. We make two assumptions: the existence of an HTDF scheme and that  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$  can be securely realized for the relation  $R_{\text{htdf}}$  defined as follows:

$$R_{\text{htdf}} = \left\{ ((\text{pk}, v_1, \dots, v_\ell, v^*), (w_1, \dots, w_\ell, u_1, \dots, u_\ell, u^*)) \mid \forall i : v_i = f_{\text{pk}, w_i}(u_i) \text{ and } v^* = f_{\text{pk}, 1}(u^*) \right\}.$$

We define the protocol in the common reference string model, with the distribution  $D_{\text{nizk}}$  that is parametrized by an HTDF scheme  $\Pi$ . The distribution  $D_{\text{nizk}}$  samples a key-pair  $(\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d)$ , and outputs  $\text{pk}$ . Note that since the public key in the construction of [57] is a random string, we can consider  $\mathcal{F}_{\text{crs}}^{D_{\text{nizk}}}$  as a common random string functionality.

The main idea of the protocol is to use the HTDF scheme to commit to the witness, bit by bit, and use the homomorphic properties to evaluate the relation on the committed witness. Now the prover can use a NIZK for the relation  $R_{\text{htdf}}$  to prove knowledge of the witness, and that the result of the homomorphic evaluation commits to 1. The simulator can use the equivocal properties of the HTDF scheme to generate an equivocal commitment that can be opened to any witness.

**Theorem 3.1.** *Assume the existence of HTDF schemes and let  $R$  be an NP-relation. The functionality  $\mathcal{F}_{\text{nizk}}^R$  can be UC-realized in the  $(\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}, \mathcal{F}_{\text{crs}}^{D_{\text{nizk}}})$ -hybrid model tolerating an adaptive, malicious adversary.*

*Proof.* Let  $\mathcal{A}$  be an adaptive, malicious adversary attacking  $\pi_{\text{nizk}}$  in the  $(\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}, \mathcal{F}_{\text{crs}})$ -hybrid model. We will construct an ideal-process adversary  $\mathcal{S}$ , interacting with the ideal functionality  $\mathcal{F}_{\text{nizk}}^R$  and with ideal (dummy) parties  $\tilde{P}_1, \dots, \tilde{P}_n$ , such that no environment can distinguish between  $\mathcal{S}$  and  $\mathcal{A}$ . Let  $\mathcal{Z}$  be an environment. Following the lines of [55], we will construct  $\mathcal{S}$  gradually, starting with a simulator with “extended capabilities” that can perfectly simulate an execution of  $\pi_{\text{nizk}}$  with  $\mathcal{A}$  and adjusting it in steps, such that the environment will have a negligible probability of distinguishing between consecutive steps.

**Protocol  $\pi_{\text{nizk}}$**

- **Common Input:** An HTDF scheme, a one-time signature scheme, and an NP-relation  $R$ .
- **Notation:** Assume that every witness to the relation  $R$  is  $\ell$ -bits long, i.e.,  $w = (w_1, \dots, w_\ell) \in \{0, 1\}^\ell$ . For a statement  $x$  denote by  $C_x$  the circuit that given input  $y \in \{0, 1\}^\ell$  computes  $C_x(y) = R(x, y)$ .
- **Hybrid model:** The parties have access to the NIZK functionality  $\mathcal{F}_{\text{nizk}}^{\text{Rhtdf}}$  and to the CRS functionality  $\mathcal{F}_{\text{crs}}^{\text{Dnizk}}$  that outputs a public key  $\text{pk}$  for the HTDF scheme.
- **The Protocol:**
  1. Upon receiving  $(\text{prove}, \text{sid}, x, w)$ , party  $P_i$  proceeds as follows:
    - (a) If  $(x, w) \notin R$  ignore the message. otherwise, denote  $w = (w_1, \dots, w_\ell)$ .
    - (b) For every  $i \in [\ell]$  compute  $v_i = f_{\text{pk}, w_i}(u_i)$  for a uniformly distributed  $u_i \leftarrow \mathcal{U}$ .<sup>a</sup>
    - (c) Compute  $u^* = \text{HTDF.Eval}^{\text{in}}(C_x, (w_1, u_1), \dots, (w_\ell, u_\ell))$ .
    - (d) Compute  $v^* = \text{HTDF.Eval}^{\text{out}}(C_x, v_1, \dots, v_\ell)$ .
    - (e) Send  $(\text{prove}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), (w_1, \dots, w_\ell, u_1, \dots, u_\ell, u^*))$  to  $\mathcal{F}_{\text{nizk}}^{\text{Rhtdf}}$ .
    - (f) Wait for the answer  $(\text{proof}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  from  $\mathcal{F}_{\text{nizk}}^{\text{Rhtdf}}$ .
    - (g) Generate signature keys  $(\text{vk}, \text{sigk}) \leftarrow \text{Sig.Gen}(1^\kappa)$ .
    - (h) Compute  $\sigma \leftarrow \text{Sign}_{\text{sigk}}((x, \text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk}))$ .
    - (i) Return  $(\text{proof}, \text{sid}, x, (\text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk}, \sigma))$ .
  2. Upon receiving  $(\text{verify}, \text{sid}, x, \Pi)$ , party  $P_j$  proceeds as follows:
    - (a) Parse  $\Pi$  as  $(\text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk}, \sigma)$ .
    - (b) Verify that  $\text{Vrfy}_{\text{vk}}(\sigma, (x, \text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk})) = 1$ . If not return  $(\text{verification}, \text{sid}, x, \Pi, 0)$ .
    - (c) Verify that  $v^* = \text{HTDF.Eval}^{\text{out}}(C_x, v_1, \dots, v_\ell)$ . If not return  $(\text{verification}, \text{sid}, x, \Pi, 0)$ .
    - (d) Send  $(\text{verify}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  to  $\mathcal{F}_{\text{nizk}}^{\text{Rhtdf}}$ .
    - (e) Wait for the answer  $(\text{verification}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi, b)$  from  $\mathcal{F}_{\text{nizk}}^{\text{Rhtdf}}$ .
    - (f) Return  $(\text{verification}, \text{sid}, x, \Pi, b)$ .

<sup>a</sup>Recall that an HTDF scheme defines a function  $f_{\text{pk}, x} : \mathcal{U} \rightarrow \mathcal{V}$ , see Definition A.7.

Figure 2: adaptively secure UC-NIZK protocol

**Simulator  $\mathcal{S}_{\text{REAL}}$ .** In the first mental experiment, the simulator  $\mathcal{S}_{\text{REAL}}$  can learn the inputs of the ideal functionality  $\mathcal{F}_{\text{nizk}}^{\text{R}}$  and control its output. It can therefore run a perfect simulation of the parties  $P_1, \dots, P_n$  and the adversary  $\mathcal{A}$  running  $\pi_{\text{nizk}}$ . The simulator  $\mathcal{S}_{\text{REAL}}$  starts by running the adversary  $\mathcal{A}$ .

- To simulate the communication with  $\mathcal{Z}$ , every input value that  $\mathcal{S}_{\text{REAL}}$  receives from  $\mathcal{Z}$  is written on  $\mathcal{A}$ 's input tape. Likewise, every output value written by  $\mathcal{A}$  on its output tape is copied to  $\mathcal{S}_{\text{REAL}}$ 's own output tape.
- To simulate the common reference string, the simulator  $\mathcal{S}$  first computes  $(\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d)$ , sets the reference string to be  $\text{pk}$ , and stores  $\text{sk}$ .

- When  $\mathcal{S}_{\text{REAL}}$  receives  $(\text{prove}, P_i, \text{sid}, x)$  from  $\mathcal{F}_{\text{nizk}}^R$ , it is because an honest party  $\tilde{P}_i$  has input  $(\text{prove}, \text{sid}, x, w)$  with  $(x, w) \in R$ . The simulator  $\mathcal{S}_{\text{REAL}}$  has access to the input of  $\mathcal{F}_{\text{nizk}}^R$ , and so it can simulate  $P_i$  running  $\pi_{\text{nizk}}$ . In particular, for simulating  $\mathcal{F}_{\text{nizk}}^{\text{htdf}}$  the simulator may send  $(\text{prove}, P_i, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*))$  to  $\mathcal{A}$  and get back the answer  $(\text{proof}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$ . Finally,  $\mathcal{S}_{\text{REAL}}$  computes a signature to prepare the proof  $\Pi = (\text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk}, \sigma)$  and answers  $(\text{proof}, \text{sid}, x, \Pi)$  to  $\mathcal{F}_{\text{nizk}}^R$ .
- When  $\mathcal{S}_{\text{REAL}}$  receives  $(\text{verify}, P_j, \text{sid}, x, \Pi)$  from  $\mathcal{F}_{\text{nizk}}^R$ , it is because an honest party  $\tilde{P}_j$  has input  $(\text{verify}, \text{sid}, x, \Pi)$  and  $(\text{sid}, x, \Pi)$  was not stored in  $\mathcal{F}_{\text{nizk}}^R$ , hence has not been created by an honest party. The simulator simulates  $P_j$  running the verification protocol on input  $(\text{verify}, \text{sid}, x, \Pi)$  until obtaining output  $(\text{verification}, \text{sid}, x, \Pi, b)$ . Next,  $\mathcal{S}_{\text{REAL}}$  instructs  $\mathcal{F}_{\text{nizk}}^R$  to return the message  $(\text{verification}, \text{sid}, x, \Pi, b)$  to  $\tilde{P}_j$ , and to store  $(\text{sid}, x, \Pi)$  if  $b = 1$ . (Note that  $\mathcal{S}_{\text{REAL}}$  does not send  $(\text{witness}, w)$  to  $\mathcal{F}_{\text{nizk}}^R$ .)
- Upon a corruption request for a party  $P_i$ , the simulator corrupts the dummy party  $\tilde{P}_i$ . For every input command of the form  $(\text{prove}, \text{sid}, x, w)$  that  $\tilde{P}_i$  received, set the random coins corresponding to the proof as  $(u_1, \dots, u_\ell)$  that were used for computing  $(v_1, \dots, v_\ell)$  along with the random coins that were used to generate the signature keys and the coins that were used to sign the message during the simulation of the proving protocol of  $\pi_{\text{nizk}}$ .

**Claim 3.2.**  $\text{REAL}_{\pi_{\text{nizk}}, \mathcal{A}, \mathcal{Z}} \equiv \text{IDEAL}_{\mathcal{S}_{\text{REAL}}, \mathcal{F}_{\text{nizk}}^R, \mathcal{Z}}$ .

*Proof.* The only difference between an execution of  $\pi_{\text{nizk}}$  in the  $(\mathcal{F}_{\text{nizk}}^{\text{htdf}}, \mathcal{F}_{\text{crs}}^{\text{D}_{\text{nizk}}})$ -hybrid model and the simulation by  $\mathcal{S}_{\text{REAL}}$ , lies in the proving protocol. In the simulation, it is guaranteed that a proof that is generated by an honest party will always be accepting (as it is generated, and stored by  $\mathcal{F}_{\text{nizk}}^R$ , when invoked with  $(\text{proof}, \text{sid}, x, w)$ ) and that accepted proofs will always be accepted again. This holds also in the execution of  $\pi_{\text{nizk}}$  by the correctness of the signature scheme and of the HTDF scheme, and by the properties of  $\mathcal{F}_{\text{nizk}}^{\text{htdf}}$ .

It follows that to the environment, a real execution of  $\pi_{\text{nizk}}$  in the  $(\mathcal{F}_{\text{nizk}}^{\text{htdf}}, \mathcal{F}_{\text{crs}}^{\text{D}_{\text{nizk}}})$ -hybrid model with adversary  $\mathcal{A}$  is perfectly indistinguishable from the simulation of  $\mathcal{S}_{\text{REAL}}$  running with  $\mathcal{F}_{\text{nizk}}^R$ .  $\square$

**Simulator  $\mathcal{S}_{\text{EXT}}$ .** The simulator  $\mathcal{S}_{\text{EXT}}$  runs like  $\mathcal{S}_{\text{REAL}}$  when the proofs are constructed, but changes the way proofs are verified. When  $\mathcal{S}_{\text{EXT}}$  receives  $(\text{verify}, P_j, \text{sid}, x, \Pi)$  from  $\mathcal{F}_{\text{nizk}}^R$ , it simulates  $P_j$  running the verification protocol on input  $(\text{verify}, \text{sid}, x, \Pi)$ . If  $P_j$  outputs  $(\text{verification}, \text{sid}, x, \Pi, 0)$ , then return  $(\text{witness}, \perp)$  to  $\mathcal{F}_{\text{nizk}}^R$ . If  $P_j$  outputs  $(\text{verification}, \text{sid}, x, \Pi, 1)$ , then  $\mathcal{S}_{\text{EXT}}$  will try to extract a witness  $w$  such that  $(x, w) \in R$ , return  $(\text{witness}, w)$  to  $\mathcal{F}_{\text{nizk}}^R$ , and send the resulting message  $(\text{verification}, \text{sid}, x, \Pi, 1)$  to  $\tilde{P}_j$  (in case the extraction fails, the simulator aborts).

To extract the witness,  $\mathcal{S}_{\text{EXT}}$  parses  $\Pi = (\text{pk}, v_1, \dots, v_\ell, v^*, \pi, \text{vk}, \sigma)$ . We can assume that the signature is valid and that  $v^* = \text{HTDF.Eval}^{\text{out}}(C_x, v_1, \dots, v_\ell)$ , since  $\pi_{\text{nizk}}$  will reject the proof otherwise. As part of the verification protocol, the functionality  $\mathcal{F}_{\text{nizk}}^{\text{htdf}}$  may send  $(\text{verify}, P_j, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  to  $\mathcal{A}$  (if  $(\text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  is not stored), who responds with  $(\text{witness}, w_1, \dots, w_\ell, u_1, \dots, u_\ell, u^*)$ . In that case, the simulator sets  $w = (w_1, \dots, w_\ell)$  and sends  $(\text{witness}, w)$  back to  $\mathcal{F}_{\text{nizk}}^R$ , and aborts otherwise.

**Claim 3.3.**  $\text{IDEAL}_{\mathcal{S}_{\text{REAL}}, \mathcal{F}_{\text{nizk}}^R, \mathcal{Z}} \stackrel{c}{\equiv} \text{IDEAL}_{\mathcal{S}_{\text{EXT}}, \mathcal{F}_{\text{nizk}}^R, \mathcal{Z}}$ .

*Proof.* The simulated  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$  will only return  $(\text{verification}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi, 1)$  if an honest party has created a proof  $\pi$  or if the adversary  $\mathcal{A}$  supplies a witness  $(w_1, \dots, w_\ell, u_1, \dots, u_\ell, u^*)$  such that  $v_i = f_{\text{pk}, w_i}(u_i)$  and  $v^* = f_{\text{pk}, 1}(u^*)$ . In the latter case, the simulator sets  $w = (w_1, \dots, w_\ell)$ . By the correctness of the HTDF scheme it holds that  $(x, w) \in R$ , therefore  $\mathcal{S}_{\text{EXT}}$  can return  $(\text{witness}, w)$  to  $\mathcal{F}_{\text{nizk}}^R$ . In case an honest party has previously created the proof  $\pi$  on  $(\text{pk}, v_1, \dots, v_\ell, v^*)$  (and so it was stored in the simulated  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$ ), then the strong existential unforgeability of the one-time signature scheme implies that the adversary is able to produce a different valid signature  $\sigma$  using  $\text{vk}$  only with a negligible probability. Therefore,  $\mathcal{S}_{\text{EXT}}$  fails to produce a witness only with a negligible probability.  $\square$

**Simulator  $\mathcal{S}_{\text{SIM}}$ .** The simulator  $\mathcal{S}_{\text{SIM}}$  runs exactly like  $\mathcal{S}_{\text{EXT}}$ , with the following modification. Instead of running a perfect simulation of  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$ , the simulator  $\mathcal{S}_{\text{SIM}}$  allows simulated honest parties to submit  $(\text{prove}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \perp)$ , even if  $(x, w) \notin R$ . This means that the simulated ideal functionality  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$  may ask  $\mathcal{A}$  for a proof  $\pi$  for a statement  $(\text{pk}, v_1, \dots, v_\ell, v^*)$ , store  $(\text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  as being a valid proof, and return  $(\text{proof}, \text{sid}, (\text{pk}, v_1, \dots, v_\ell, v^*), \pi)$  to the calling party  $P_i$ .

When  $\mathcal{S}_{\text{SIM}}$  receives  $(\text{prove}, P_i, \text{sid}, x)$  from  $\mathcal{F}_{\text{nizk}}^R$ , it changes the way the proof is simulated. Instead of computing the values  $v_1, \dots, v_\ell$  as  $v_i = f_{\text{pk}, w_i}(u_i)$  for  $u_i \leftarrow \mathcal{U}$ , the simulator samples uniformly distributed values  $v_1, \dots, v_\ell \leftarrow \mathcal{V}$ . Upon a corruption request for a party  $P_i$ , the simulator corrupts the dummy party  $\tilde{P}_i$ . For every input command of the form  $(\text{prove}, \text{sid}, x, w)$  that  $\tilde{P}_i$  received, set the random coins corresponding to the proof as  $(u_1, \dots, u_\ell)$ , where  $u_i = \text{HTDF.Inv}_{\text{sk}, w_i}(v_i)$ .

**Claim 3.4.**  $\text{IDEAL}_{\mathcal{S}_{\text{EXT}}, \mathcal{F}_{\text{nizk}}^R, \mathcal{Z}} \stackrel{s}{=} \text{IDEAL}_{\mathcal{S}_{\text{SIM}}, \mathcal{F}_{\text{nizk}}^R, \mathcal{Z}}$ .

*Proof.* This follows from the security of the HTDF scheme by using a standard hybrid argument. First, we show that the simulated proofs of  $\mathcal{S}_{\text{EXT}}$  are statistically close to the simulated proofs of  $\mathcal{S}_{\text{SIM}}$ . We define  $\ell + 1$  experiments, where in the  $i$ 'th experiment the first  $i - 1$  values  $v_j$  are computed as  $v_j \leftarrow \mathcal{V}$  and are explained as  $u_j \leftarrow \text{HTDF.Inv}_{\text{sk}, w_j}(v_j)$ . The remaining  $\ell - i + 1$  values are computed as  $v_j = f_{\text{pk}, w_j}(u_j)$  for  $u_j \leftarrow \mathcal{U}$ . The first experiment is exactly the way  $\mathcal{S}_{\text{EXT}}$  simulates the proof and the last experiment is exactly the way  $\mathcal{S}_{\text{SIM}}$  simulates the proof. By the properties of the HTDF scheme neighboring experiments are statistically close, and therefore also the proof simulated by  $\mathcal{S}_{\text{EXT}}$  and  $\mathcal{S}_{\text{SIM}}$ .

Second, since the number of **prove** requests is bounded by a polynomial number, we can use a second hybrid argument to replace the way the proofs are simulated and explained one by one, and conclude that the simulation of  $\mathcal{S}_{\text{EXT}}$  is statistically close to the simulation of  $\mathcal{S}_{\text{SIM}}$ .  $\square$

Note that  $\mathcal{S}_{\text{SIM}}$  does not require any access to the internals of  $\mathcal{F}_{\text{nizk}}^R$ . Therefore, we can now define the simulator  $\mathcal{S}$  that does not have any access to the internals of  $\mathcal{F}_{\text{nizk}}^R$ , and operates exactly like  $\mathcal{S}_{\text{SIM}}$ . As we have shown,  $\mathcal{S}$  running with  $\mathcal{F}_{\text{nizk}}^R$  is computationally indistinguishable from the protocol  $\pi_{\text{nizk}}$  running with  $\mathcal{A}$  in the  $(\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}, \mathcal{F}_{\text{crs}})$ -hybrid model. The protocol  $\pi_{\text{nizk}}$  therefore securely realizes  $\mathcal{F}_{\text{nizk}}^R$  in the  $(\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}, \mathcal{F}_{\text{crs}})$ -hybrid model.  $\square$

By instantiating the  $\mathcal{F}_{\text{nizk}}^{R_{\text{htdf}}}$  with the UC-NIZK protocol from [60], we obtain the following corollary.

**Corollary 3.5.** *Under the decision linear assumption (or equivalently, the subgroup decision problem) in bilinear group, and the existence of HTDF schemes, the following holds. For every NP-relation  $R$ , the functionality  $\mathcal{F}_{\text{nizk}}^R$  can be UC-realized in the  $\mathcal{F}_{\text{crs}}$ -hybrid model tolerating an adaptive, malicious adversary such that the size of the CRS is  $\text{poly}(\kappa)$  and the size of the proofs is  $|w| \cdot \text{poly}(\kappa)$ .*

## 4 Sublinear Communication in the Fully Adaptive Setting

In this section, we consider the fully adaptive setting (where the adversary can corrupt all parties) and construct two-round secure protocols with sublinear communication and online-computation complexity (in the circuit size). Our starting point is the protocol of Quach et al. [84] that is based on *laconic function evaluation (LFE)* (see Appendix A.2.1).

### 4.1 Cryptographic Primitives used in the Protocol

We informally describe the cryptographic primitives used in the construction. Formal definitions appear in the appendix.

**Laconic function evaluation.** We formally define laconic function evaluation (LFE) in Appendix A.2.1. Informally, an LFE scheme consists of 4 algorithms. The CRS generation algorithm generates a common random string given the security parameter and function parameters (e.g., function depth and input length)  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ . The compression algorithm produces a small digest of a circuit  $\text{digest}_C = \text{LFE.Compress}(\text{crs}, C; r)$ . The encryption algorithm encrypts the input based on the digest  $\text{ct} \leftarrow \text{LFE.Enc}(\text{crs}, \text{digest}_C, x)$ . The decryption algorithm decrypts the ciphertext using the random coins used in the compression  $y = \text{LFE.Dec}(\text{crs}, C, r, \text{ct})$ .

We require the LFE to be correct, i.e., using the notation above it holds that  $y = C(x)$ , and secure, meaning that the ciphertext can be simulated given the output value  $y$  without knowing the input  $x$ . LFE can be constructed with the function-hiding property, which ensures that the digest can be simulated based on the function parameters without knowing the function itself. If function hiding is not required (as is the case in this section) the compression algorithm can be made deterministic. We consider the “adaptive” version of LFE, where the inputs to the computation can be chosen *after* the CRS has been sampled. Quach et al. [84] constructed LFE schemes satisfying this property assuming *adaptive LWE* (see Appendix A.1.2).

**Explainability compilers.** We formally define explainability compilers in Appendix A.2.2. Informally, an explainability compiler takes as input a description of a randomized algorithm  $\text{Alg}$ , and outputs two algorithms:  $\widetilde{\text{Alg}}$  and  $\text{Explain}$ . The first algorithm  $\widetilde{\text{Alg}}$  computes the same functionality as  $\text{Alg}$ . The second algorithm  $\text{Explain}$  takes an input/output pair  $(x, y)$  and produces random coins  $r$  such that  $y = \widetilde{\text{Alg}}(x; r)$ .

### 4.2 Adaptive Security with Sublinear Communication: Secure-Erasures Setting

We will show that assuming LFE every function can be securely realized in two rounds facing an arbitrary number of adaptive corruptions with sublinear communication and online-computation complexity, in the common random string model with secure erasures. In Section 4.3, we will show how to replace the secure-erasures assumption by assuming explainability compilers, in which case

the protocol requires a reusable, input-independent common reference string, which depends on the depth of the function.

The basis of our protocol is the 2-round protocol of Quach et al. [84, Thm. 6.2] in the common random string model, that is secure against  $n - 1$  static corruptions and achieves sublinear communication and online-computation assuming the existence of LFE. The protocol from [84] is specified in a hybrid model with an ideally secure computation of the function  $\text{LFE.Enc}$  with abort (i.e., the  $\mathcal{F}_{\text{sfe-abort}}^{\text{LFE.Enc}}$ -hybrid model). That is, the ideal functionality receives  $(\text{crs}, \text{digest}_f, x_i, r_i)$  from each party  $P_i$  and computes

$$\text{ct} = \text{LFE.Enc}(\text{crs}, \text{digest}_f, x_1, \dots, x_n; \bigoplus_{i \in [n]} r_i).$$

In case of inconsistent inputs, or if the adversary sends `abort`, the functionality outputs  $\perp$ .

Given as a circuit  $C_f$  computing the function  $f$ , the protocol of [84] is defined as follows:

- The common random string is computed as  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, f, \text{params})$ .
- Upon receiving  $(\text{input}, \text{sid}, x_i)$ , every party  $P_i$  computes  $\text{digest}_f = \text{LFE.Compress}(\text{crs}, C_f)$ , samples a uniformly random  $r_i \leftarrow \{0, 1\}^*$ , and invokes the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^{\text{LFE.Enc}}$  with  $(\text{input}, \text{sid}, (\text{crs}, \text{digest}_f, x_i, r_i))$ .
- Upon receiving  $(\text{output}, \text{sid}, \text{ct})$  from the ideal functionality, party  $P_i$  check that  $\text{ct} \neq \perp$  (otherwise,  $P_i$  outputs  $(\text{output}, \text{sid}, \perp)$ ), computes  $y = \text{LFE.Dec}(\text{crs}, C_f, \text{ct})$ , and outputs  $(\text{output}, \text{sid}, y)$ .

Proving security of the protocol against a static adversary corrupting all-but-one of the parties is straightforward. Namely, by definition of LFE schemes, the simulator can simulate the ciphertext  $\text{ct}$  without knowing all of the input values as  $\text{ct} \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C_f, \text{digest}_f, y)$ . Furthermore, by the properties of LFE, the size of the circuit computing  $\text{LFE.Enc}$  is  $\text{poly}(\kappa, \ell_{\text{in}}, \ell_{\text{out}}, d, n)$ . By instantiating the ideal functionality using a statically secure 2-round protocol (e.g., the one from [78]), Quach et al. [84] achieved a statically secure protocol with sublinear communication and online-computational complexity.

A closer look at the protocol of [84] shows that it remains secure even facing adaptive corruptions of all-but-one of the parties, since a single honest party suffices to keep the randomness used for  $\text{LFE.Enc}$  hidden from the adversary. Furthermore, under the additional assumption of secure erasures, each party can erase its random coins  $r_i$  immediately after invoking  $\mathcal{F}_{\text{sfe-abort}}^{\text{LFE.Enc}}$ , and the protocol can satisfy adaptive corruptions of all the parties. By instantiating the functionality  $\mathcal{F}_{\text{sfe-abort}}^{\text{LFE.Enc}}$  with the 2-round adaptively secure MPC from [13], we obtain the following theorem.

**Theorem 4.1.** *Assume the existence of LFE schemes for  $\text{P/poly}$  and of 2-round adaptively and maliciously secure OT, and let  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  be an  $n$ -party function of depth  $d$ . Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized tolerating a malicious, adaptive PPT adversary by a 2-round protocol in the common random string model. The size of the common random string is  $\text{poly}(\kappa)$ , whereas the communication and online-computational complexity of the protocol are  $\text{poly}(\kappa, \ell_{\text{in}}, \ell_{\text{out}}, d, n)$ .*

Note that following [84, 13], the assumptions in Theorem 4.1 hold under the adaptive LWE assumption.

### 4.3 Adaptive Security with Sublinear Communication: Erasures-Free Setting

In the erasures-free setting, it is unclear how to simulate the output ciphertext, and later upon learning all of the inputs values of the parties, explain the random coins that are used to generate it. We get around this barrier by using explainability compilers.

To define the common reference string for the protocol, we define the distribution  $D_{\text{lfe}}(\text{params})$  that is parametrized by an LFE scheme and by the parameters of the function to be computed  $\text{params}$ . The distribution  $D_{\text{lfe}}$  computes  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$  and  $(\widetilde{\text{LFE.Enc}}, \text{Explain}) \leftarrow \text{Comp}(1^\kappa, \text{LFE.Enc})$ , and outputs the reference string  $(\text{crs}, \widetilde{\text{LFE.Enc}})$ .

We want to define the protocol in the  $\widetilde{\text{LFE.Enc}}$ -hybrid model, however, the function  $\widetilde{\text{LFE.Enc}}$  is only given in the CRS and is not known before the protocol begins. To get around this technicality, we define the function  $f_C((C_1, x_1, r_1), \dots, (C_n, x_n, r_n))$  that receives a circuit  $C_i$ , a value  $x_i$ , and random coins  $r_i$  from each party, and outputs  $C_1(x_1, \dots, x_n; \oplus r_i)$  in case  $C_1 = \dots = C_n$ , or  $\perp$  otherwise.

**Lemma 4.2.** *Assume the existence of LFE schemes for P/poly, and of explainability compilers for P/poly, and let  $f$  be a deterministic  $n$ -party function. Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized tolerating a malicious, adaptive PPT adversary in the  $(\mathcal{F}_{\text{crs}}^{D_{\text{lfe}}(f, \text{params})}, \mathcal{F}_{\text{sfe-abort}}^{f_C})$ -hybrid model.*

*Proof.* We adjust the protocol from [84] as follows:

- Upon receiving  $(\text{input}, \text{sid}, x_i)$ , every party  $P_i$  invokes  $\mathcal{F}_{\text{crs}}^{D_{\text{lfe}}(f, \text{params})}$  to get  $(\text{crs}, \widetilde{\text{LFE.Enc}})$ , computes  $\text{digest}_f = \text{LFE.Compress}(\text{crs}, C_f)$ , samples a uniformly random  $r_i \leftarrow \{0, 1\}^*$ , and invokes  $\mathcal{F}_{\text{sfe-abort}}^{f_C}$  with  $(\text{input}, \text{sid}, (\widetilde{\text{LFE.Enc}}, (\text{crs}, \text{digest}_f, x_i), r_i))$ .
- Upon receiving  $\text{ct}$  from the ideal functionality, party  $P_i$  checks that  $\text{ct} \neq \perp$  (if so  $P_i$  outputs  $(\text{output}, \text{sid}, \perp)$ ), computes  $y = \text{LFE.Dec}(\text{crs}, C_f, \text{ct})$ , and outputs  $(\text{output}, \text{sid}, y)$ .

Let  $\mathcal{A}$  be an adaptive, malicious adversary attacking in the  $(\mathcal{F}_{\text{crs}}^{D_{\text{lfe}}(f, \text{params})}, \mathcal{F}_{\text{sfe-abort}}^{f_C})$ -hybrid model. We will construct an ideal-process adversary  $\mathcal{S}$ , interacting with the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^f$  and with ideal (dummy) parties  $\tilde{P}_1, \dots, \tilde{P}_n$ , such that no environment can distinguish between  $\mathcal{S}$  and  $\mathcal{A}$ . Let  $\mathcal{Z}$  be an environment. As there is no interaction between the parties, the simulation should only explain corruption requests. The simulator initializes a value  $\text{ct}^* = \perp$ .

- To simulate the CRS,  $\mathcal{S}$  computes  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$  and  $(\widetilde{\text{LFE.Enc}}, \text{Explain}) \leftarrow \text{Comp}(1^\kappa, \text{LFE.Enc})$ . Next,  $\mathcal{S}$  gives  $\mathcal{A}$  the string  $(\text{crs}, \widetilde{\text{LFE.Enc}})$  and stores the function  $\text{Explain}$ .
- Upon a corruption request of a party  $P_i$  proceeds as follows;
  - If not all parties are corrupted, sample uniformly distributed random coins  $r_i \leftarrow \{0, 1\}^*$  and set the message to  $\mathcal{F}_{\text{sfe-abort}}^{f_C}$  to be  $(\text{input}, \text{sid}, (\widetilde{\text{LFE.Enc}}, (\text{crs}, \text{digest}_f, x_i), r_i))$ . If the party  $P_i$  received output  $y$  from  $\mathcal{F}_{\text{sfe-abort}}^f$ , set the response from  $\widetilde{\text{LFE.Enc}}$  to be  $(\text{output}, \text{sid}, \text{ct}^*)$ , where if  $\text{ct}^*$  has not been initialized yet (i.e.,  $\text{ct}^* = \perp$ ), compute  $\text{ct}^* \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C_f, \text{digest}_f, y)$ .
  - To explain the last corruption request (say of party  $P_n$ ), the simulator computes  $r \leftarrow \text{Explain}((\text{crs}, \text{digest}_f, x_1, \dots, x_n), \text{ct}^*)$ , and sets  $r_n = r \oplus r_1 \oplus \dots \oplus r_{n-1}$ .



We prove computational indistinguishability of the execution of the protocol with adversary  $\mathcal{A}$  from the ideal process with  $\mathcal{F}_{\text{sfe-abort}}^f$  and  $\mathcal{S}$  by defining two hybrid experiments:

1. The first hybrid is defined as the execution of the protocol, with the following modification. Upon a corruption of the first  $n - 1$  parties, set its random coins to be a uniformly random string  $r_i \leftarrow \{0, 1\}^*$ . The coins of the last corrupted party are computed by first setting  $r \leftarrow \text{Explain}((\text{crs}, \text{digest}_f, x_1, \dots, x_n), \text{ct})$ , and next computing  $r_n = r \oplus r_1 \oplus \dots \oplus r_{n-1}$ . This experiment is computationally indistinguishable from the real execution of the protocol by the security of the explainability compiler.
2. The second hybrid is defined as the first hybrid, but the functionality  $\widetilde{\mathcal{F}_{\text{sfe-abort}}^{fC}}$  operates differently. Upon receiving inputs  $(\text{input}, \text{sid}, (\widetilde{\text{LFE.Enc}}, (\text{crs}, \text{digest}_f, x_i), r_i))$  from all the parties, the functionality computes  $y = f(x_1, \dots, x_n)$ , creates a simulated ciphertext  $\text{ct} \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C_f, \text{digest}_f, y)$ , and returns  $\text{ct}$  to the parties. This experiment is computationally indistinguishable from the first hybrid by the security of the LFE scheme.

The second hybrid experiment operates identically as the simulator in the ideal model computing  $\mathcal{F}_{\text{sfe-abort}}^f$ , and the lemma follows.  $\square$

Using the protocol from [13] to instantiate the  $\mathcal{F}_{\text{sfe-abort}}^{fC}$  functionality that is used to compute  $\widetilde{\text{LFE.Enc}}$ , we obtain the following theorem.

**Theorem 4.3.** *Assume the existence of LFE schemes for  $\text{P/poly}$ , of explainability compilers for  $\text{P/poly}$ , and of 2-round adaptively and maliciously secure OT, and let  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  be a deterministic  $n$ -party function of depth  $d$ . Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized tolerating a malicious, adaptive PPT adversary by a 2-round protocol in the  $\mathcal{F}_{\text{crs}}^{D_{\text{lfe}}}$ -hybrid model. The size of the common reference string, the communication complexity, and online-computational complexity of the protocol are  $\text{poly}(\kappa, \ell_{\text{in}}, \ell_{\text{out}}, d, n)$ .*

## 5 Adaptively Secure Alice/Bob-Optimized Protocols

In this section, we consider 2-message protocols between Alice and Bob, with respective inputs  $x_A \in \{0, 1\}^{\ell_A}$  and  $x_B \in \{0, 1\}^{\ell_B}$ , where only Alice learns the output  $y = f(x_A, x_B)$ . We say that a protocol is “Alice-optimized” if Alice’s computation and the total communication of the protocol are proportional to  $|x_A| + |y|$ , while the computation complexity of Bob is proportional to  $f$ . We say that a protocol is “Bob-optimized” if Bob’s computation and the total communication of the protocol are proportional to  $|x_B| + |y|$ , while the computation complexity of Alice is proportional to  $f$ .

There exist insecure protocols which are Alice-optimized, where Alice sends her input to Bob who computes to function and returns the output to Alice. Similarly, there exist insecure protocols which are Bob-optimized, where Bob sends his input to Alice when she asks for it, and Alice computes the function.

Assuming FHE [51], there exist statically secure Alice-optimized protocols, where Alice sends her encrypted input to Bob who homomorphically evaluates the function and returns the encrypted output to Alice. Alice’s computation and the total communication of the protocol are  $(|x_A| + |y|) \cdot \text{poly}(\kappa)$ . Assuming function-hiding LFE [84] there exist statically secure Bob-optimized protocols,

where Alice sends  $\text{digest} \leftarrow \text{LFE.Compress}(\text{crs}, f_{x_A}(\cdot))$  to Bob, who replies with his encrypted input  $\text{ct} \leftarrow \text{LFE.Enc}(\text{digest}, x_B)$ , and finally Alice recovers the output. Bob’s computation and the total communication of the protocol are  $(|x_B| + |y|) \cdot \text{poly}(\kappa, d)$ , where  $d$  is the depth of the function  $f$ .

The question we consider is whether there exist adaptively secure protocols which are Alice-optimized or Bob-optimized.

## 5.1 Fully Adaptive Bob-Optimized Protocol

The elegant protocol from [84] is secure in the common random string model tolerating a static corruption of one of the parties by a semi-malicious adversary (that can choose arbitrary random coins for the corrupted party, but acts honestly otherwise).

Adjusting this protocol to the adaptive setting requires overcoming a few obstacles. Namely, the simulator should be able to generate an equivocal first message, i.e., to simulate the digest without knowing the input value of Alice, and upon a later corruption of Alice generate appropriate random coins explaining the message. Similarly, the simulator should be able to generate an equivocal second message, i.e., generate the ciphertext without knowing the input of Bob, and upon a later corruption of Bob provide appropriate random coins.

To support an adaptive corruption of Alice, we enhance the LFE scheme to support an equivocal mode (see Section 5.1.1). In this mode, the CRS is generated along with a trapdoor information. The trapdoor can be used to explain a simulated digest as a compression of any circuit with the appropriate parameters. Similarly to Section 4, to support an adaptive corruption of Bob, we can use either secure erasures or explainability compilers (see Appendix A.2.2).

**Theorem 5.1.** *Assume the existence of NCE, of equivocal, function-hiding LFE schemes for  $\text{P/poly}$ , and of explainability compilers for  $\text{P/poly}$ , and let  $f : \{0, 1\}^{\ell_A} \times \{0, 1\}^{\ell_B} \rightarrow \{0, 1\}^{\ell_{out}}$  be a deterministic two-party function computable by a depth- $d$  circuit. Then,  $\mathcal{F}_{\text{sfe}}^f$  can be UC-realized tolerating a semi-malicious, adaptive PPT adversary by a 2-message protocol in the common reference string model. The size of the common reference string, the communication complexity, and the computational complexity of Bob are  $(\ell_B + \ell_{out}) \cdot \text{poly}(\kappa, d)$ .*

*Proof.* The proof is given in Lemma 5.5 using secure channels, that can be instantiated over authenticated channels assuming NCE [21, 40, 31, 35].  $\square$

In the secure-erasures setting, we can remove the explainability compilers assumption, and get the following corollary.

**Corollary 5.2.** *Assume the existence of NCE and of equivocal, function-hiding LFE schemes for  $\text{P/poly}$ , and let  $f$  be a two-party function as above. Then,  $\mathcal{F}_{\text{sfe}}^f$  can be UC-realized in the secure-erasures model tolerating a semi-malicious, adaptive PPT adversary by a 2-message protocol in the common random string model. The size of the common random string is  $\text{poly}(\kappa)$ , and the communication complexity and computational complexity of Bob are  $(\ell_B + \ell_{out}) \cdot \text{poly}(\kappa, d)$ .*

### 5.1.1 Equivocal LFE

We start by extending the notion of LFE to support an equivocal mode.

**Definition 5.3** (equivocal LFE). *A function-hiding LFE scheme  $\Pi$  is equivocal if there exists a PPT simulator  $(\text{Sim}_{\text{EQUIV-FH}}^1, \text{Sim}_{\text{EQUIV-FH}}^2)$  for the scheme  $\Pi$  such that for all stateful PPT adversary  $\mathcal{A}$ , it holds that*

$$\left| \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{EquivFH-real}}(\kappa) = 1 \right] - \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{EquivFH-ideal}}(\kappa) = 1 \right] \right| \leq \text{negl}(\kappa),$$

for the experiments  $\text{Expt}^{\text{EquivFH-real}}$  and  $\text{Expt}^{\text{EquivFH-ideal}}$  defined below:

$\text{Expt}_{\Pi, \mathcal{A}}^{\text{EquivFH-real}}(\kappa)$	$\text{Expt}_{\Pi, \mathcal{A}}^{\text{EquivFH-ideal}}(\kappa)$
$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ $C \leftarrow \mathcal{A}(\text{crs})$ <i>s.t.</i> $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $r \leftarrow \{0, 1\}^*$ $\text{digest} = \text{LFE.Compress}(\text{crs}, C; r)$ <i>Output</i> $\mathcal{A}(\text{crs}, \text{digest}, r)$	$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $(\text{crs}, \text{digest}, \text{state}) \leftarrow \text{Sim}_{\text{EQUIV-FH}}^1(1^\kappa, \text{params})$ $C \leftarrow \mathcal{A}(\text{crs})$ <i>s.t.</i> $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $r \leftarrow \text{Sim}_{\text{EQUIV-FH}}^2(C, \text{state})$ <i>Output</i> $\mathcal{A}(\text{crs}, \text{digest}, r)$

We next show that the generic construction of function-hiding LFE from standard LFE presented in [84] can be adjusted to provide equivocality.

**Lemma 5.4.** *Assuming the existence of standard LFE schemes there exists a function-hiding, equivocal LFE scheme.*

*Proof sketch.* Quach et al. [84, Sec. 5] showed how to construct a function-hiding LFE scheme given any standard LFE scheme. Recall that in an LFE scheme Alice sends  $\text{digest}$  to Bob, who replies with  $\text{LFE.Enc}(\text{crs}, \text{digest}, x)$ . Instead, Alice and Bob run a semi-malicious, 2-message 2PC protocol (e.g., a protocol based on semi-malicious OT and garbled circuits), where Alice has input  $\text{digest}$ , Bob has input  $x$ , and Alice learns  $\text{LFE.Enc}(\text{crs}, \text{digest}, x)$ . The message of Alice is the new digest and the message of Bob is the new ciphertext. By using a 2-round 2PC based on 2-round OT and garbled circuits, the transformation preserves the asymptotic efficiency of the construction, i.e., the digest size and the ciphertext size only grow by some fixed  $\text{poly}(\kappa)$  multiplicative factor.

In the FH-LFE scheme described above, the digest consists of the receiver messages in the 2-round OT protocol. It follows that if the OT protocol satisfies *equivocal receiver's security* [49], the FH-LFE scheme will be equivocal. Loosely speaking, equivocal receiver's security means that there exists a simulator that can simulate the first message of the receiver, and later, given the input bit, can provide random coins that explain this message according to the bit. Garg and Srinivasan [49, Lem. B.1] showed that assuming two-round maliciously secure OT protocol (in the static setting), there exists a two-round maliciously secure OT protocol with equivocal receiver's security. The construction from [49] immediately translates to the semi-malicious, and the proof follows as LFE implies two-round semi-malicious OT [84].  $\square$

### 5.1.2 Semi-Malicious Bob-optimized Protocol

We proceed to our Bob-optimized protocol. Recall that the distribution  $D_{\text{lfe}}(\text{params})$  samples a  $\text{crs}$  for the LFE scheme.

**Lemma 5.5.** *Consider the notations and assumptions in Theorem 5.1. Then, Protocol  $\pi_{\text{bob}}$  securely realizes the functionality  $\mathcal{F}_{\text{sfe}}^f$  tolerating a semi-malicious, adaptive PPT adversary in the  $(\mathcal{F}_{\text{smt}}, \mathcal{F}_{\text{crs}}^{D_{\text{lfe}}(f, \text{params})})$ -hybrid model.*

**Protocol  $\pi_{\text{bob}}$**

- **Common Input:** An LFE scheme and a circuit  $C_f$  computing the function  $f$ .
- **Notation:** Define the algorithm  $\text{LFE.Compress}_{\text{crs}, C_f}(x) := \text{LFE.Compress}(\text{crs}, C_f(x, \cdot))$  by hard-wiring  $\text{crs}$  and the circuit  $C_f$  to the compression algorithm, and given input  $x$  compresses the circuit  $C_f(x, \cdot)$  with the input  $x$  hard-wired.
- **The Protocol:**
  1. Upon receiving (input, sid,  $x_A$ ), Alice proceeds as follows:
    - (a) Compute  $\text{digest} = \text{LFE.Compress}_{\text{crs}, C_f}(x_A; r)$  for a uniformly distributed  $r \leftarrow \{0, 1\}^*$ .
    - (b) Send (sid, digest) to Bob.
  2. Upon receiving a message (sid, digest) from Alice, and having received (input, sid,  $x_B$ ), Bob proceeds as follows:
    - (a) Compute  $\text{ct} \leftarrow \widetilde{\text{LFE.Enc}}(\text{crs}, \text{digest}, x_B)$ .
    - (b) Send (sid, ct) to Alice.
  3. Upon receiving a message (sid, ct) from Bob, Alice computes  $y = \text{LFE.Dec}(\text{crs}, C, r, \text{ct})$  and outputs (output, sid,  $y$ ).

Figure 3: Two-round, Bob-optimized protocol with adaptive, semi-malicious security

*Proof.* Let  $\mathcal{A}$  be an adaptive, semi-malicious adversary attacking  $\pi_{\text{bob}}$  in the  $(\mathcal{F}_{\text{smt}}, \mathcal{F}_{\text{crs}}^{D_{\text{lfe}}(f.\text{params})})$ -hybrid model. We will construct an ideal-process adversary  $\mathcal{S}$ , interacting with the ideal functionality  $\mathcal{F}_{\text{sfe}}^f$  and with ideal (dummy) parties  $\tilde{A}$  and  $\tilde{B}$ , such that no environment can distinguish between  $\mathcal{S}$  and  $\mathcal{A}$ . Let  $\mathcal{Z}$  be an environment. Let  $\text{Sim}_{\text{LFE}}$  and  $\text{Sim}_{\text{EQUIV-FH}}$  be the simulators guaranteed to exist by the security of the equivocal FH-LFE scheme. The simulator  $\mathcal{S}$  constructs virtual parties Alice and Bob, and runs the adversary  $\mathcal{A}$ .

- To simulate the communication with  $\mathcal{Z}$ , every input value that  $\mathcal{S}$  receives from  $\mathcal{Z}$  is written on  $\mathcal{A}$ 's input tape. Likewise, every output value written by  $\mathcal{A}$  on its output tape is copied to  $\mathcal{S}$ 's own output tape.
- To simulate the common reference string, the simulator  $\mathcal{S}$  first computes  $(\text{crs}, \text{digest}, \text{state}) \leftarrow \text{Sim}_{\text{EQUIV-FH}}^1(1^\kappa, C_f.\text{params})$  and  $(\widetilde{\text{LFE.Enc}}, \text{Explain}) \leftarrow \text{Comp}(\text{LFE.Enc})$ . Next,  $\mathcal{S}$  sets the reference string to be  $(\text{crs}, \widetilde{\text{LFE.Enc}})$  and stores  $(\text{digest}, \text{state}, \text{Explain})$ .

Since the protocol is defined over secure channels, there is no need to simulate anything as long as both parties are honest. Upon the first corruption of one of the parties, the simulator learns its input and can simulate the internal state as if the party was corrupted from the beginning. We can therefore simplify the proof and consider two cases: the case where Alice is statically corrupted at the beginning and Bob is dynamically corrupted, and the case where Bob is statically corrupted at the beginning and Alice is dynamically corrupted.<sup>8</sup>

<sup>8</sup>This is effectively the compiler from semi-adaptive security to adaptive security of Garay et al. [45].

**Case 1: Alice is corrupted first.** The simulator corrupts  $\tilde{A}$  and proceeds as follows:

- Upon receiving  $(\text{input}, \text{sid}, x_A)$  from the environment,  $\mathcal{S}$  invokes  $\mathcal{A}$  with  $(\text{input}, \text{sid}, x_A)$  and receives back the message  $(\text{sid}, \text{digest})$  from  $\mathcal{A}$ . The simulator reads the random coins  $r_A$  that were used by  $\mathcal{A}$  from the witness tape of the semi-malicious  $\mathcal{A}$ .
- Next,  $\mathcal{S}$  sends  $(\text{input}, \text{sid}, x_A)$  to  $\mathcal{F}_{\text{sfe}}^f$  and receives back the output  $(\text{output}, \text{sid}, y)$ .
- To simulate the second message,  $\mathcal{S}$  computes  $\text{ct} \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C_f(x_A, \cdot), r_A, \text{digest}, y)$  and sends  $(\text{sid}, \text{ct})$  to  $\mathcal{A}$ .
- To explain a corruption request of Bob, the simulator corrupts  $\tilde{B}$ , learn its input  $x_B$ , and computes the random coins as  $r_B \leftarrow \text{Explain}((\text{crs}, \text{digest}, x_B), \text{ct})$ . (In case Bob was not activated with an input yet, set  $r_B$  to be a uniformly random string.)

**Proposition 5.6.** *Assuming the first corruption request is of Alice, it holds that*

$$\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{IDEAL}_{\mathcal{F}_{\text{sfe}}^f, \mathcal{S}, \mathcal{Z}}.$$

*Proof.* We prove the indistinguishability of the real and ideal worlds by defining a series of hybrid games. The output of each game is the output of the environment.

**The game  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^1$ .** In this game, we modify the real-model experiment  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$  as follows. Instead of computing the common reference string  $(\text{crs}, \widetilde{\text{LFE.Enc}})$  using  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, C_f.\text{params})$ , it is computed using  $(\text{crs}, \text{digest}, \text{state}) \leftarrow \text{Sim}_{\text{EQUIV-FH}}^1(1^\kappa, C_f.\text{params})$ .

**Claim 5.7.**  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^1$ .

*Proof.* This follows by the definition of equivocal LFE.  $\square$

**The game  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2$ .** In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^1$  as follows. Bob first computes  $\text{ct} = \widetilde{\text{LFE.Enc}}(\text{digest}, x_B; r)$  followed by  $r_B \leftarrow \text{Explain}((\text{crs}, \text{digest}, x_B), \text{ct})$ . Next, the original coins  $r$  are ignored, and the new coins  $r_B$  are set as the random tape of Bob.

**Claim 5.8.**  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^1 \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2$ .

*Proof.* This follows by the security of the explainability compiler  $\text{Comp}$ , since any distinguisher between  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^1$  and  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2$  immediately translates into an adversary for the experiment  $\text{Expt}^{\text{Explain}}$  (see Definition A.5) that succeeds with the same probability.  $\square$

**The game  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^3$ .** In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2$  as follows. Instead of computing  $\text{ct} \leftarrow \widetilde{\text{LFE.Enc}}(\text{crs}, \text{digest}, x_B)$  followed by  $r_B \leftarrow \text{Explain}((\text{crs}, \text{digest}, x_B), \text{ct})$ , Bob computes  $\text{ct}^* \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C_f(x_A, \cdot), r_A, \text{digest}, y)$  (where  $r_A$  are the coins used to compute  $\text{digest}$ ) followed by  $r_B \leftarrow \text{Explain}((\text{crs}, \text{digest}, x_B), \text{ct}^*)$ .

**Claim 5.9.**  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2 \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^3$ .

*Proof.* This follows by the security of the LFE scheme, since any distinguisher between  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^2$  and  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^3$  immediately translates into a distinguisher between the experiments  $\text{Expt}_{\text{LFE-real}}^{\text{LFE-ideal}}$  (see Appendix A.2.1) that succeeds with the same probability.  $\square$

**Claim 5.10.**  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}^3 \equiv \text{IDEAL}_{f, \mathcal{S}, \mathcal{Z}}$ .

*Proof.* This follows since in the experiment  $\text{HYB}^3$  behaviour of Bob is identical to the simulation done by  $\mathcal{S}$ . In particular, the second message  $\text{ct}$  is computed via the LFE simulator without knowing the input value  $x_B$ , and the random coins are set using  $\text{Explain}$ .  $\square$

Combining Claims 5.7 to 5.10, we get that  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{IDEAL}_{\mathcal{F}_{\text{sfe}}^f, \mathcal{S}, \mathcal{Z}}$  when the first corruption is of Alice. This concludes the proof of Proposition 5.6.  $\square$

**Case 2: Bob is corrupted first.** The simulator corrupts  $\tilde{B}$  and proceeds as follows:

- Upon receiving  $(\text{input}, \text{sid}, x_B)$  from the environment, and having received the notification  $(\text{input}, \text{sid}, \tilde{A})$  from  $\mathcal{F}_{\text{sfe}}^f$ , the simulator sends  $(\text{sid}, \text{digest})$  to  $\mathcal{A}$  (where  $\text{digest}$  was computed at the beginning of the simulation).
- Next,  $\mathcal{S}$  sends  $(\text{input}, \text{sid}, x_B)$  to  $\mathcal{F}_{\text{sfe}}^f$  and receives back the output  $(\text{output}, \text{sid}, y)$ .
- Finally,  $\mathcal{S}$  receives the second message  $(\text{sid}, \text{ct})$  from  $\mathcal{A}$ .
- To explain a corruption request of Alice, the simulator corrupts  $\tilde{A}$ , learn her input  $x_A$ , denotes by  $C_1 = C_f(x_A, \cdot)$  the circuit  $C_f$  with the value  $x_A$  hard-wired, and computes the random coins as  $r_A \leftarrow \text{Sim}_{\text{EQUIV-FH}}^2(C_1, \text{state})$ .

**Proposition 5.11.** *Assuming the first corruption request is of Bob, it holds that*

$$\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{IDEAL}_{\mathcal{F}_{\text{sfe}}^f, \mathcal{S}, \mathcal{Z}}$$

*Proof.* We prove the indistinguishability of the real and ideal worlds by defining a hybrid game. The output of the game is the output of the environment.

**The game**  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$ . In this game, we modify the real-model experiment  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$  as follows. Instead of computing the common reference string  $(\text{crs}, \widetilde{\text{LFE.Enc}})$  using  $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, C_f.\text{params})$ , it is computed using  $(\text{crs}, \text{digest}, \text{state}) \leftarrow \text{Sim}_{\text{EQUIV-FH}}^1(1^\kappa, C_f.\text{params})$ . Instead of computing  $\text{digest} \leftarrow \text{LFE.Compress}(\text{crs}, C_1; r_A)$  for  $C_1 = C_f(x_A, \cdot)$ , Alice uses  $\text{digest}$  as computed by  $\text{Sim}_{\text{EQUIV-FH}}^1$  followed by  $r_A \leftarrow \text{Sim}_{\text{EQUIV-FH}}^2(C_1, \text{state})$ . Next, the original coins  $r$  are ignored, and the new coins  $r_A$  are set as the random tape of Alice.

**Claim 5.12.**  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$ .

*Proof.* This follows by the security of the equivocal LFE scheme, since any distinguisher between  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$  and  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}}$  immediately translates into an adversary violating Definition 5.3.  $\square$

**Claim 5.13.**  $\text{HYB}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \equiv \text{IDEAL}_{\mathcal{F}_{\text{sfe}}^f, \mathcal{S}, \mathcal{Z}}$ .

*Proof.* This follows since in the experiment HYB the behaviour of Alice is identical to the simulation done by  $\mathcal{S}$ . In particular, the first message digest is computed using  $\text{Sim}_{\text{EQUIV-FH}}^1$  without knowing the input value  $x_A$ , and the random coins are set using  $\text{Sim}_{\text{EQUIV-FH}}^2$ .  $\square$

It follows that  $\text{REAL}_{\pi_{\text{bob}}, \mathcal{A}, \mathcal{Z}} \stackrel{c}{\equiv} \text{IDEAL}_{\mathcal{F}_{\text{sf}}, \mathcal{S}, \mathcal{Z}}$  when Bob is corrupted first. This concludes the proof of Proposition 5.11.  $\square$

This concludes the proof of Lemma 5.5.  $\square$

## 5.2 Impossibility of Adaptively Secure Alice-Optimized Protocol

We now turn to show that the impossibility of adaptively secure FHE from [74] can be extended to rule out adaptively secure Alice-optimized protocols. In fact, we prove a stronger impossibility showing that for some functions the size of Bob's message cannot be smaller than his input, even if Alice's message and the CRS are long. Intuitively, if the output of the function is simply Bob's input, then clearly Bob's message cannot be compressing. We show that this is the case even if the output is short.

For  $n \in \mathbb{N}$ , we define the two-party functionality  $f_n(x_A, g_B) = (g_B(x_A), \lambda)$ , where Alice has input  $x_A \in \{0, 1\}^{\log n}$ , Bob has input a function  $g_B : \{0, 1\}^{\log n} \rightarrow \{0, 1\}$ , represented by its truth table as an  $n$ -bit string, and Alice learns the output  $g_B(x_A)$ .

**Theorem 5.14.** *Let  $\pi_n$  be a 2-message protocol in the common reference string model for computing  $f_n$ , where Alice sends first the message  $m_1$  and Bob replies with the message  $m_2$ . If the protocol tolerates a semi-honest, adaptive adversary in the secure-erasures model, then  $|m_2| \geq n$ .*

*Proof.* Assume towards contradiction that  $|m_2| < n$ . The protocol  $\pi_n$  can be described as follows:

- The common reference string is set as  $\text{crs} \leftarrow \text{crsGen}(1^\kappa, 1^n)$  according to some distribution.
- Alice generates the first message as  $m_1 = \text{Alice}_1(\text{crs}, x_A; r_A)$ , and possibly erases some of the coins  $r_A$ , leaving  $r'_A$ .
- Bob generates the second message as  $m_2 = \text{Bob}(\text{crs}, g_B, m_1; r_B)$ , and possibly erases  $r_B$ .
- Alice computes the output as  $y = \text{Alice}_2(\text{crs}, x_A, r'_A, m_2)$ .

Consider the adversary  $\mathcal{A}$  and the environment  $\mathcal{Z}$  that operate as follows:

1. The adversary  $\mathcal{A}$  corrupts Bob.
2. The environment activates Alice with input  $x_A$ .
3. The adversary waits to receive the message  $m_1$ , corrupts Alice to learn her (partial) random coins  $r'_A$ , and sends  $(\text{crs}, m_1, r'_A)$  to the environment.
4. The environment, chooses a random function  $g_B \leftarrow \{0, 1\}^n$  and random coins  $r_B \leftarrow \{0, 1\}^*$ , runs the code for Bob to generate the message  $m_2 = \text{Bob}(\text{crs}, g_B, m_1; r_B)$ , and checks whether  $g_B(x_A) = \text{Alice}_2(\text{crs}, x_A, r'_A, m_2)$ . if so, the environment outputs 1 (real) and otherwise outputs 0 (ideal).



By the security of the protocol  $\pi_n$ , there exists a simulator  $\mathcal{S}$  that can simulate this attack. Initially, the simulator generates the simulated  $\text{crs}$  by running an algorithm  $(\text{crs}, \text{state}_1) \leftarrow \mathcal{S}_1(1^\kappa, 1^n)$ . On a corruption request of Bob, the simulator corrupts the ideal (dummy) Bob, and sends  $\text{crs}$  to  $\mathcal{A}$ . Next, to simulate the first message  $\mathcal{S}$  computes  $(m_1, \text{state}_2) \leftarrow \mathcal{S}_2(\text{state}_1)$  and sends  $m_1$  to  $\mathcal{A}$ . Finally, upon a corruption request of Alice, the simulator corrupts the ideal (dummy) Alice, learns her input  $x_A$ , and computes  $r'_A \leftarrow \mathcal{S}_3(x_A, \text{state}_2)$ .

Given a random function  $g : \{0, 1\}^{\log n} \rightarrow \{0, 1\}$ , we will use the simulator  $\mathcal{S}$  to construct a circuit  $C^*(x)$  for computing the function  $g$ . We start by computing  $(\text{crs}, \text{state}_1) \leftarrow \mathcal{S}_1(1^\kappa, 1^n)$  and  $(m_1, \text{state}_2) \leftarrow \mathcal{S}_2(\text{state}_1)$  just like  $\mathcal{S}$ . Note that this part of the simulation is independent of the parties' inputs. Next, run the code of Bob on input  $g$  to compute  $m_2 = \text{Bob}(\text{crs}, g, m_1; r_B)$  for a uniformly distributed  $r_B \leftarrow \{0, 1\}^*$ . Finally, sample random coins  $r \leftarrow \{0, 1\}^*$ , and hardwire the value  $(\text{crs}, \text{state}_2, m_2, r)$  to  $C^*$ . On input  $x$ , the resulting circuit  $C_{\text{crs}, \text{state}_2, m_2, r}^*(x)$  first computes  $r'_A = \mathcal{S}_3(x, \text{state}_2; r)$  and later  $y = \text{Alice}_2(\text{crs}, x, r'_A, m_2)$ .

By the security of the protocol  $\pi_n$  it holds that with overwhelming probability, it is possible to recover the string  $g$  by running  $C_{\text{crs}, \text{state}_2, m_2, r}^*(x)$  on every  $x \in \{0, 1\}^{\log n}$ . However, since the only component in  $C_{\text{crs}, \text{state}_2, m_2, r}^*$  that depends on  $g$  is  $m_2$ , it holds that

$$H(G \mid C_{\text{crs}, \text{state}_2, m_2, r}^*) = n - |m_2| > 0,$$

where  $G$  is a random variable that takes a value uniformly at random in  $\{0, 1\}^n$ , and  $C_{\text{crs}, \text{state}_2, m_2, r}^*$  is the random variable representing the output of the algorithm described above. Combined together, we derive a contradiction.  $\square$

## 6 Adaptive Corruptions of All-But-One of the Parties

In this section, we prove an analogue result in the adaptive setting to the result of Asharov et al. [5], who showed how to compute any function tolerating all-but-one corruptions using a two-round protocol in the threshold-PKI model assuming threshold FHE, which in turn can be instantiated using LWE. Our construction relies on *threshold equivocal FHE* (to be defined in Section 6.1.1) that allows simulating ciphertexts for honest parties and explaining them properly upon later corruptions.

We note that the simulation technique used in [5] (and similarly in [78]) does not translate to the adaptive setting. As observed in [5, 78], the threshold decryption protocol may leak some information about the shares of the secret key, and the simulator for the decryption protocol can be used to protect *exactly* one party. Since [5, 78] considered static corruptions, the set of corrupted parties was known ahead of time, and the simulator could choose one of the honest parties  $P_h$  as a special party for the simulation. The decryption protocol was simulated with respect to  $P_h$ , as if he is the only honest party. For this reason, proving security of *exactly*  $n - 1$  corruptions in [78] was considerably simpler than proving security of *up to*  $n - 1$  corruptions.<sup>9</sup>

The simulation strategy that was used in [5, 78] does not translate to the adaptive setting, since the party  $P_h$  that is chosen by the simulator may get corrupted after simulating the decryption protocol. The simulator cannot know in advance which party will be the last to remain honest. For this reason, we use a different simulation strategy, which allows the simulator to “correct” its

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<sup>9</sup>We note that the same problem arises also in the threshold FHE scheme for more general access structures [15, Def. 5.5], where the simulation is defined only for *maximal* invalid party sets.

choice of the party that is simulated as honest for the decryption protocol. Technically, this is done by having each party send shares of zero to each other party over a secure channel (that can be instantiated via NCE). These shares are used to hide the partial decryptions without changing their value. Since shares exchanged between pairs of honest parties remain hidden from the eyes of the adversary, the simulator has more freedom to replace the special party  $P_h$  upon corruption, by another honest party, even after simulating the decryption protocol.

## 6.1 Cryptographic Primitives used in the Protocol

Initially, we define equivocal FHE and show how to instantiate it assuming LWE. Next, we define threshold equivocal FHE.

### 6.1.1 Equivocal FHE

An equivocal FHE is an FHE scheme that is augmented with the capability to generate a public key in an “equivocal mode,” allowing to explain any ciphertext as an encryption of any value.

**Definition 6.1.** *An equivocal fully homomorphic encryption (EFHE) scheme is a six-tuple of algorithms (EFHE.Gen, EFHE.Enc, EFHE.Eval, EFHE.Dec, EFHE.GenEquiv, EFHE.Equiv) satisfying the following properties:*

- (EFHE.Gen, EFHE.Enc, EFHE.Eval, EFHE.Dec) is an FHE scheme.
- $\text{EFHE.GenEquiv}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{td})$ : on input the security parameter  $\kappa$  and a depth bound  $d$ , the equivocal key-generation algorithm outputs a public-key  $\text{pk}$  and a trapdoor  $\text{td}$ .
- $\text{EFHE.Equiv}(\text{td}, \text{ct}, m) \rightarrow r$ : on input a trapdoor  $\text{td}$ , a ciphertext  $\text{ct}$ , and a plaintext  $m$ , the equivocation algorithm outputs a value  $r$  in the randomness space.

We require the following properties:

1. (EFHE.Gen, EFHE.Enc, EFHE.Eval, EFHE.Dec) is a correct, compact, and semantically secure FHE scheme (see Appendix A.2.6).
2. For every depth bound  $d$  and for every message  $\mu \in \{0, 1\}$ , the following distributions are computationally indistinguishable:
  - The public key, a ciphertext encrypting  $\mu$ , the plaintext  $\mu$ , and the random coins used in an honest encryption

$$\{(\text{pk}, \text{ct}, \mu, r) \mid (\text{sk}, \text{pk}) \leftarrow \text{EFHE.Gen}(1^\kappa, 1^d), \text{ct} = \text{EFHE.Enc}(\text{pk}, \mu; r)\}_{\kappa}.$$

- The equivocal public key, an encryption of zero, the plaintext  $\mu$ , and the computed random coins generated by the equivocation algorithm

$$\{(\text{pk}, \text{ct}, \mu, \text{EFHE.Equiv}(\text{td}, \text{ct}, \mu)) \mid (\text{pk}, \text{td}) \leftarrow \text{EFHE.GenEquiv}(1^\kappa, 1^d), \text{ct} \leftarrow \text{EFHE.Enc}_{\text{pk}}(0)\}_{\kappa}.$$

**Constructing Equivocal FHE.** We next show that equivocal FHE schemes exist assuming LWE. We will prove this statement using an extended, “dual-mode” version of homomorphic trapdoor functions from [57]. Recall that HTDF, were defined in [57] in a “meaningless” (or “equivocal”) mode (see also Appendix A.2.3). That is, by computing  $(\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d)$  it is possible to “commit” to  $x \in \{0, 1\}$  by computing  $v = f_{\mathbf{pk}, x}(u)$  for a random  $u \in \mathcal{U}$ . Equivalently, it is possible to sample a random element  $v \leftarrow \mathcal{V}$  and “explain” it to every  $x \in \{0, 1\}$  by computing  $u \leftarrow \text{HTDF.Inv}_{\mathbf{sk}, x}(v)$ . To avoid confusion, we will denote in the remaining of this section the key-generation algorithm in this mode by  $\text{HTDF.GenEquiv}$ . Note that using the equivocal mode, the commitment to  $x$  is *statistically hiding*.

Gorbunov et al. [57] also defined a “meaningful” (or “extractable”) mode of homomorphic trapdoor functions. In this mode, the keys are generated via a different algorithm  $(\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.GenExtract}(1^\kappa, 1^d)$ . The public key can again be used to “commit” to  $x \in \{0, 1\}$  by computing  $v = f_{\mathbf{pk}, x}(u)$  for a random  $u \in \mathcal{U}$ . However, the secret key is now used to extract the bit  $x$  from  $v$ , i.e.,  $x \leftarrow \text{HTDF.Extract}_{\mathbf{sk}}(v)$ . Note that using the extractable mode, the commitment to  $x$  is *statistically binding* and *computationally hiding*.

Gorbunov et al. [57] constructed a dual-mode HTDF scheme and proved that under the LWE assumption, the following properties are satisfied:

1.  $(\text{HTDF.GenEquiv}, f, \text{HTDF.Inv}, \text{HTDF.Eval}^{\text{in}}, \text{HTDF.Eval}^{\text{out}})$  is an HTDF scheme according to Definition A.7.
2. For any depth-bound  $d$  and every  $x \in \{0, 1\}$  it holds that

$$\Pr \left[ \text{HTDF.Extract}_{\mathbf{sk}}(v) \neq x \mid \begin{array}{l} (\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.GenExtract}(1^\kappa, 1^d) \\ v = f_{\mathbf{pk}, x}(u) \text{ for } u \leftarrow \mathcal{U} \end{array} \right] \leq \text{negl}(\kappa).$$

3. The public key in the equivocal mode is computationally indistinguishable from the public key in the extractable mode:

$$\left\{ \mathbf{pk} \mid (\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.GenEquiv}(1^\kappa, 1^d) \right\}_\kappa \stackrel{c}{\equiv} \left\{ \mathbf{pk} \mid (\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.GenExtract}(1^\kappa, 1^d) \right\}_\kappa.$$

We can now proceed to construct an equivocal FHE from dual-mode HTDF. The extractable mode of the HTDF scheme will be used to define the “standard FHE,” since extracting the bit  $x$  from  $v$  corresponds to decrypting a ciphertext (indeed, the extractable mode in [57] corresponds to the GSW FHE scheme [54]). The equivocal mode of the HTDF scheme will be used to define the equivocal mode of the FHE scheme, as it allows explaining a ciphertext to every bit. Formally:

- $\text{EFHE.Gen}(1^\kappa, 1^d)$ : run  $(\mathbf{pk}, \mathbf{sk}) \leftarrow \text{HTDF.GenExtract}(1^\kappa, 1^d)$  and return  $(\mathbf{pk}, \mathbf{sk})$ .
- $\text{EFHE.Enc}(\mathbf{pk}, x; r)$ : parse  $r$  as an element  $u \in \mathcal{U}$ , compute  $v = f_{\mathbf{pk}, x}(u)$ , and return  $v$  as the ciphertext.
- $\text{EFHE.Eval}(\mathbf{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell)$ : parse each ciphertext  $\text{ct}_i$  as an element  $v_i \in \mathcal{V}$ , compute  $v^* = \text{HTDF.Eval}^{\text{out}}(C, v_1, \dots, v_\ell)$ , and return  $v^*$  as the evaluated ciphertext.
- $\text{EFHE.Dec}(\mathbf{sk}, \text{ct})$ : parse  $\text{ct}$  as an element  $v \in \mathcal{V}$ , compute  $x \leftarrow \text{HTDF.Extract}_{\mathbf{sk}}(v)$ , and return  $x$  as the plaintext.

- $\text{EFHE.GenEquiv}(1^\kappa, 1^d)$ : run  $(\text{pk}, \text{sk}) \leftarrow \text{HTDF.GenEquiv}(1^\kappa, 1^d)$  and return  $(\text{pk}, \text{sk})$ , i.e.,  $\text{sk}$  acts as the trapdoor  $\text{td}$ .
- $\text{EFHE.Equiv}(\text{td}, \text{ct}, x)$ : parse  $\text{ct}$  as an element  $v \in \mathcal{V}$  and  $\text{td}$  as  $\text{sk}$ , compute  $u \leftarrow \text{HTDF.Inv}_{\text{sk}, x}(v)$ , and return  $u$  as the random coins.

The security of the equivocal FHE scheme follows immediately from the security of the dual-mode HTDF scheme. This proves the following lemma.

**Lemma 6.2.** *Assuming LWE there exist equivocal FHE schemes.*

*Proof.* As proven above, equivocal FHE schemes can be constructed from dual-mode HTDF schemes, and the security of the EFHE scheme follows immediately from the security of the dual-mode HTDF scheme. The lemma follows as dual-mode HTDF schemes exist under the LWE assumption [57]. For completeness, we describe the construction in Appendix C.  $\square$

### 6.1.2 Threshold Equivocal FHE

In a threshold FHE scheme, the key generation and the decryption algorithms are in fact  $n$ -party protocols. We consider the simplest case of  $n$ -out-of- $n$  threshold FHE and require a single round decryption protocol (following [5, 58, 78, 43]). We note that threshold FHE for more general access structures are also known assuming LWE [15]. A threshold equivocal FHE scheme (TEFHE) is a threshold FHE scheme that admits an equivocal mode as in Section 6.1.1.

**Definition 6.3** (TEFHE). *A threshold equivocal fully homomorphic encryption (TEFHE) is a seven-tuple of algorithms  $(\text{TEFHE.Gen}, \text{TEFHE.Enc}, \text{TEFHE.Eval}, \text{TEFHE.PartDec}, \text{TEFHE.FinDec}, \text{TEFHE.GenEquiv}, \text{TEFHE.Equiv})$  satisfying the following properties:*

- $\text{TEFHE.Gen}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{sk}_1, \dots, \text{sk}_n)$ : on input the security parameter  $\kappa$  and a depth bound  $d$ , the key-generation algorithm outputs a public key  $\text{pk}$  and  $n$  secret key shares  $\text{sk}_1, \dots, \text{sk}_n$ .
- $\text{TEFHE.Enc}(\text{pk}, \mu) \rightarrow \text{ct}$ : on input a public key  $\text{pk}$  and a plaintext  $\mu \in \{0, 1\}$ , the encryption algorithm outputs a ciphertext  $\text{ct}$ .
- $\text{TEFHE.Eval}(\text{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell) \rightarrow \text{ct}$ : on input a public key  $\text{pk}$ , a circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$ , and a tuple of ciphertexts  $(\text{ct}_1, \dots, \text{ct}_\ell)$ , the homomorphic evaluation algorithm outputs a ciphertext  $\text{ct}$ .
- $\text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}) \rightarrow \text{p}_i$ : on input a secret key share  $\text{sk}_i$  and a ciphertext  $\text{ct}$ , the partial decryption algorithm outputs a partial decryption  $\text{p}_i$ .
- $\text{TEFHE.FinDec}(\text{pk}, \text{p}_1, \dots, \text{p}_n) \rightarrow \tilde{\mu}$ : on input a public key  $\text{pk}$  and a set  $\{\text{p}_i\}_{i \in [n]}$ , the final decryption algorithm outputs  $\tilde{\mu} \in \{0, 1, \perp\}$ .
- $\text{TEFHE.GenEquiv}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{td})$ : on input the security parameter  $\kappa$  and a depth bound  $d$ , the equivocal key-generation algorithm outputs a public-key  $\text{pk}$  and a trapdoor  $\text{td}$ .
- $\text{TEFHE.Equiv}(\text{td}, \text{ct}, m) \rightarrow r$ : on input a trapdoor  $\text{td}$ , a ciphertext  $\text{ct}$ , and a plaintext  $m$ , the equivocation algorithm outputs a value  $r$  in the randomness space.

We require the following properties:

1. The FHE scheme that is defined by setting the decryption key  $\mathbf{sk} = (\mathbf{sk}_1, \dots, \mathbf{sk}_n)$  and the decryption algorithm is composed of executing  $\text{TEFHE.PartDec}(i, \mathbf{sk}_i, \text{ct})$  for every  $i \in [n]$  followed by  $\text{TEFHE.FinDec}(\mathbf{pk}, \mathbf{p}_1, \dots, \mathbf{p}_n)$  is a correct, compact, and secure equivocal FHE scheme for circuits of depth  $d$  (according to Definition 6.1, see also Appendix A.2.6).
2. Simulatability of partial decryption: there exists a PPT simulator  $\text{Sim}_{\text{TEFHE}}$  such that on input  $i \in [n]$ , and all decryption keys except of the  $i$ 'th one  $\{\mathbf{sk}_j\}_{j \neq i}$ . The following distributions are statistically close:

$$\{\mathbf{p}_i \mid \mathbf{p}_i \leftarrow \text{TEFHE.PartDec}(i, \mathbf{sk}_i, \text{ct})\} \stackrel{c}{\equiv} \{\mathbf{p}'_i \mid \mathbf{p}'_i \leftarrow \text{Sim}_{\text{TEFHE}}(i, \text{ct}, \mu, \{\mathbf{sk}_j\}_{j \neq i})\},$$

where the keys are set as  $(\mathbf{pk}, \mathbf{sk}_1, \dots, \mathbf{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$ , the ciphertext is set as  $\text{ct} \leftarrow \text{TEFHE.Eval}(\mathbf{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell)$  for a circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$  and for  $i \in [\ell]$  ciphertext  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\mathbf{pk}, \mu_i)$  with  $\mu_i \in \{0, 1\}$ , and  $\mu = C(\mu_1, \dots, \mu_\ell)$ .

In the protocol, we will require some additional properties regarding the key-generation and threshold-decryption protocols.

**Definition 6.4** (special TEFHE). *A special TEFHE is a TEFHE scheme satisfying the following properties:*

1. On input  $1^\kappa$  and  $1^d$ , the key-generation algorithm  $\text{TEFHE.Gen}$  outputs  $(\mathbf{pk}, \mathbf{sk}_1, \dots, \mathbf{sk}_n)$  where the public key  $\mathbf{pk}$  defines a prime number  $q$ , and each secret key  $\mathbf{sk}_i$  is uniformly distributed in  $\mathbb{Z}_q^{n'}$  for some  $n' = \text{poly}(\kappa, d)$ .
2. The partial decryption algorithm  $\mathbf{p}_i \leftarrow \text{TEFHE.PartDec}(i, \mathbf{sk}_i, \text{ct})$  operates by computing  $\mathbf{p}_i = \langle \text{ct}, \mathbf{sk}_i \rangle + e \pmod q$ .
3. For every  $v_1, \dots, v_n \in \mathbb{Z}_q$ , the final decryption algorithm  $\text{TEFHE.FinDec}(\mathbf{pk}, \mathbf{p}_1, \dots, \mathbf{p}_n)$  satisfies the following linearity property

$$\text{TEFHE.FinDec}(\mathbf{pk}, \mathbf{p}_1 + v_1, \dots, \mathbf{p}_n + v_n) = \text{TEFHE.FinDec}(\mathbf{pk}, \mathbf{p}_1, \dots, \mathbf{p}_n) + \sum_{i \in [n]} v_i.$$

**Lemma 6.5.** *Assuming LWE there exist special TEFHE schemes.*

*Proof.* The lemma follows from Lemma 6.2 using standard techniques for constructing threshold FHE [5, 78]. For completeness, we describe the construction in Appendix C.  $\square$

## 6.2 The Threshold-PKI Functionality

We define the protocol assuming a trusted pre-process phase where the keys of the TEFHE scheme are generated and distributed to the parties. We capture this assumption with the threshold-PKI ideal functionality.

**Functionality**  $\mathcal{F}_{\text{thresh-pki}}(\Pi, d)$

$\mathcal{F}_{\text{thresh-pki}}$  proceeds as follows, interacting with parties  $P_1, \dots, P_n$  and an adversary  $\mathcal{S}$ , and parameterized by a TEFHE scheme  $\Pi$  and a depth bound  $d$ .

- Upon receiving a message  $(\text{init}, \text{sid})$  from party  $P_i$ , do:
  1. If there is no value  $(\text{sid}, \text{pk}, \text{sk}_1, \dots, \text{sk}_n)$  recorded, then compute  $(\text{pk}, \text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$ , record it, and send  $(\text{sid}, P_i)$  to  $\mathcal{S}$ .
  2. Send  $(\text{sid}, P_i, \text{pk})$  to  $\mathcal{S}$  and a delayed output  $(\text{sid}, \text{pk}, \text{sk}_i)$  to  $P_i$ .

Figure 4: The threshold-PKI functionality

### 6.3 Semi-Malicious Security

**Theorem 6.6.** *Assume that special TEFHE exists, let  $t < n$ , and let  $f : (\{0, 1\}^{\ell_{in}})^n \rightarrow \{0, 1\}^{\ell_{out}}$  be an efficiently computable function. Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model, tolerating an adaptive, semi-malicious, PPT  $t$ -adversary, by a two-round protocol with communication complexity  $\text{poly}(\ell_{in}, \ell_{out}, \kappa, n)$ .*

**Protocol**  $\pi_{\text{allbutone}}$

- **Private Input:** Every party  $P_i$ , for  $i \in [n]$ , has private input  $x_i \in \{0, 1\}^{\ell_{in}}$ .
- **Common Input:** A special TEFHE scheme  $\Pi$  and a circuit  $C_f$  of depth  $d$ .
- **The Protocol:**
  1. Upon receiving  $(\text{input}, \text{sid}, x_i)$ , party  $P_i$  proceeds as follows:
    - (a) Invoke  $\mathcal{F}_{\text{thresh-pki}}(\Pi, d)$  with  $(\text{init}, \text{sid})$  to receive  $(\text{sid}, \text{pk}, \text{sk}_i)$ . Let  $q$  be the prime associated with the public key  $\text{pk}$  (as per Definition 6.4).
    - (b) Encrypt the input as  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, x_i)$ .
    - (c) Sample uniformly distributed  $\text{s}_i^1, \dots, \text{s}_i^n \leftarrow \mathbb{Z}_q$ , conditioned on  $\sum_{j=1}^n \text{s}_i^j = 0 \pmod q$ .
    - (d) Send  $(\text{sid}, \text{ct}_i, \text{s}_i^j)$  to  $P_j$  over a secure channel (via  $\mathcal{F}_{\text{smt}}$ ).
  2. In case some party aborts, output  $(\text{output}, \text{sid}, \perp)$  and halt. Otherwise, upon receiving  $(\text{sid}, \cdot)$  messages from all the parties, party  $P_i$  proceeds as follows:
    - (a) Compute  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_f, \text{ct}_1, \dots, \text{ct}_n)$ .
    - (b) Partially decrypt the result as  $\text{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$ .
    - (c) Set  $\text{m}_i = \text{p}_i + \sum_{j=1}^n \text{s}_j^i \pmod q$  and send  $(\text{sid}, \text{m}_i)$  to every party.
  3. In case some party aborts, output  $(\text{output}, \text{sid}, \perp)$  and halt. Otherwise, upon receiving  $(\text{sid}, \cdot)$  from all the parties, party  $P_i$  runs the final decrypt as  $y = \text{TEFHE.FinDec}(\text{pk}, \{\text{m}_1, \dots, \text{m}_n\})$  and outputs  $(\text{output}, \text{sid}, y)$ .

Figure 5: Two-round MPC with semi-malicious security in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model

*Proof.* We start by showing correctness of a non-aborting execution. In this case, for every  $i \in [n]$  it holds that  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, x_i)$  and  $\sum_{j=1}^n s_j^i = 0 \pmod q$ ; in addition,  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_f, \text{ct}_1, \dots, \text{ct}_n)$ . Next, each party computes  $\mathbf{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$  and sends  $\mathbf{m}_i = \mathbf{p}_i + \sum_{j=1}^n s_j^i \pmod q$ . By the properties of special TEFHE, it holds that

$$\begin{aligned}
\text{TEFHE.FinDec}(\text{pk}, \{\mathbf{m}_1, \dots, \mathbf{m}_n\}) &= \text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_1 + \sum_{j=1}^n s_j^1, \dots, \mathbf{p}_n + \sum_{j=1}^n s_j^n\}) \\
&= \text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_1, \dots, \mathbf{p}_n\}) + \sum_{j=1}^n s_j^1 + \dots + \sum_{j=1}^n s_j^n \\
&= \text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_1, \dots, \mathbf{p}_n\}) + \sum_{j=1}^n s_j^1 + \dots + \sum_{j=1}^n s_j^n \\
&= \text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_1, \dots, \mathbf{p}_n\}) + 0 + \dots + 0 \pmod q \\
&= C(x_1, \dots, x_n).
\end{aligned}$$

We proceed to prove security of the protocol. Let  $\mathcal{A}$  be an adaptive, semi-malicious adversary attacking  $\pi_{\text{allbutone}}$  in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model. We will construct an ideal-process adversary  $\mathcal{S}$ , interacting with the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^f$  and with ideal (dummy) parties  $\tilde{P}_1, \dots, \tilde{P}_n$ , such that no environment can distinguish between  $\mathcal{S}$  and  $\mathcal{A}$ . The simulator  $\mathcal{S}$  constructs virtual parties  $P_1, \dots, P_n$ , and runs the adversary  $\mathcal{A}$ . Let  $\text{Sim}_{\text{TEFHE}}$  be the simulator that is guaranteed to exist for the TEFHE scheme  $\Pi$ .

Note that the internal state of a party  $P_i$  in the protocol consists of its input  $x_i$ , the message  $(\text{pk}, \text{sk}_i)$  from  $\mathcal{F}_{\text{thresh-pki}}$ , the random coins that include coins  $r_i^E$  for encrypting the input in the first round, coins  $r_i^D$  for partial decryption of the evaluated ciphertext for the second round, and the secret shares of zero  $(s_i^1, \dots, s_i^n)$ .<sup>10</sup> In addition, the internal state contains the messages  $P_i$  received in the first round, including the values  $(\text{ct}_1, \dots, \text{ct}_n)$  and  $(s_1^i, \dots, s_n^i)$ , and in the second round  $(\mathbf{m}_1, \dots, \mathbf{m}_n)$ .

**Simulating the communication  $\mathcal{Z}$ :** Every input value that  $\mathcal{S}$  receives from  $\mathcal{Z}$  is written on  $\mathcal{A}$ 's input tape. Likewise, every output value written by  $\mathcal{A}$  on its output tape is copied to  $\mathcal{S}$ 's own output tape.

**Simulating the threshold-PKI functionality  $\mathcal{F}_{\text{thresh-pki}}$ :** The simulator  $\mathcal{S}$  computes  $(\text{pk}, \text{td}) \leftarrow \text{TEFHE.GenEquip}(1^\kappa, 1^d)$ , for every  $i \in [n]$  samples  $\text{sk}_i \leftarrow \mathbb{Z}_q^{n'}$  (as per Definition 6.4), and gives  $(\text{pk}, \text{sk}_i)$  to  $\mathcal{A}$  for every corrupted party  $P_i$ .

**Simulating the first round:** For every honest party  $P_i$ , compute  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, 0^{\ell_{\text{in}}})$ , sample a secret sharing of zero  $(s_i^1, \dots, s_i^n)$  (i.e.,  $\sum_j s_j^i = 0 \pmod q$ ), and send  $(\text{sid}, \text{ct}_i, s_i^j)$  to every corrupted  $P_j$ . Finally, the simulator receives from  $\mathcal{A}$  the message  $(\text{sid}, \text{ct}_j, s_j^i)$  on behalf of every corrupted party  $P_j$  and every honest party  $P_i$ , and reads from the special witness tape of the semi-malicious adversary  $\mathcal{A}$  the input and random coins  $(x_i', r_i)$ .

<sup>10</sup>The coin that determine  $(s_i^1, \dots, s_i^n)$  can be, for example, an encoding of the first  $n-1$  elements  $(s_i^1, \dots, s_i^{n-1})$  that fully determine  $s_i^n$ .



**Sending input to the functionality  $\mathcal{F}_{\text{sfe-abort}}^f$ :** In case the adversary aborted in the first round, send  $(\text{abort}, \text{sid})$  to the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^f$ . Otherwise, for every corrupted party  $P_i$ , send the input value  $x'_i$  to the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^f$  and receive back the output  $y$ . The simulator does not release the output to the ideal (dummy) honest parties yet.

**Simulating the second round:** The simulator chooses an arbitrary honest party  $P_{h_1}$ . Let  $\text{ct}_1, \dots, \text{ct}_n$  be the ciphertexts in the first round of the simulation,  $\mathcal{S}$  homomorphically evaluates the circuit as  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_f, \text{ct}_1, \dots, \text{ct}_n)$ , computes for every honest party  $P_i$  (except for  $P_{h_1}$ ) the partial decryption  $\mathbf{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}; r_i^D)$ , for uniformly random  $r_i^D$ , and for  $P_{h_1}$  the simulator computes the simulated partial decryption  $\mathbf{p}_{h_1} \leftarrow \text{Sim}_{\text{TEFHE}}(h_1, \text{ct}, y, \{\text{sk}_j\}_{j \neq h_1})$ .

Next, for each honest party  $P_i$  (including  $P_{h_1}$ ), let  $(\mathbf{s}_1^i, \dots, \mathbf{s}_n^i)$  be the shares sent to  $P_i$  from every corrupted party in the first round (by  $\mathcal{A}$ ), and were sampled to be sent to  $P_i$  from every other honest party during the simulation (by  $\mathcal{S}$ ). The simulator  $\mathcal{S}$  sets  $\mathbf{m}_i = \mathbf{p}_i + \sum_j \mathbf{s}_j^i \pmod q$ , gives  $(\text{sid}, \mathbf{m}_i)$  to  $\mathcal{A}$ , and receives messages from  $\mathcal{A}$  on behalf of corrupted parties.

In case the adversary aborts in the second round, send  $(\text{abort}, \text{sid})$  to the ideal functionality  $\mathcal{F}_{\text{sfe-abort}}^f$ . Otherwise, release the output values to the ideal honest parties.

**Simulating corruption requests of party  $P_i$ :** The simulator  $\mathcal{S}$  corrupts the dummy party  $\tilde{P}_i$  and continues as follows, depending on the timing of the corruption:

- **Corruptions before the first round:** In this case  $\tilde{P}_i$  has not been activated with input yet.  $\mathcal{S}$  sets the contents of  $P_i$ 's random tape with uniformly random coins.
- **Corruptions after the first round and before the second round:** In this case  $\tilde{P}_i$  has been activated with input  $(\text{input}, \text{sid}, x_i)$ , but did not receive output yet.  $\mathcal{S}$  sets the contents of  $P_i$ 's input tape to  $(\text{input}, \text{sid}, x_i)$  and the contents of the random tape to  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$  for the ciphertext  $\text{ct}_i$  that was used to simulate  $P_i$ 's message, and to a uniformly random  $r_i^D$  (representing the random coins for the partial-decryption algorithm). Finally, the secret shares of zero  $(\mathbf{s}_1^i, \dots, \mathbf{s}_n^i)$  have already been set for  $P_i$  during the simulation. Similarly, the messages received in the first round, i.e.,  $(\mathbf{s}_1^i, \dots, \mathbf{s}_n^i)$  are set according to the simulation both for corrupted parties (as received from  $\mathcal{A}$ ) and for honest parties (as simulated by  $\mathcal{S}$ ).
- **Corruptions after the second round:** In this case  $\tilde{P}_i$  has received the output.  $\mathcal{S}$  computes  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$  as in the previous case, but in order to set the random coins of the partial decryption,  $\mathcal{S}$  proceeds as follows. Let  $h_k$  be the index of the recent party for which the decryption share was simulated (i.e.,  $h_k$  with the largest  $k$ ).
  - If  $i \neq h_k$ , i.e., the corrupted party is *not* the party that was used to simulate the threshold decryption, then set the decryption coins  $r_i^D$  as the coins used in the simulation of the second round. Set the shares of zero  $(\mathbf{s}_1^i, \dots, \mathbf{s}_n^i)$  and the incoming messages  $(\mathbf{s}_1^i, \dots, \mathbf{s}_n^i)$  according to the values set in the simulation.
  - If  $i = h_k$ , then sample uniformly random  $r_{h_k}^D$  and compute the partial decryption  $\mathbf{p}_{h_k} = \text{TEFHE.PartDec}(h_k, \text{sk}_{h_k}, \text{ct}; r_{h_k}^D)$ . Next, choose an arbitrary honest party  $P_{h_{k+1}}$  and simulate the partial decryption as  $\mathbf{p}_{h_{k+1}} \leftarrow \text{Sim}_{\text{TEFHE}}(h_{k+1}, \text{ct}, y, \{\text{sk}_j\}_{j \neq h_{k+1}})$ . Finally, the simulator needs to “adjust” the values for  $P_{h_k}$  and  $P_{h_{k+1}}$ . In particular, it must hold that:

1.  $\mathbf{m}_{h_k} = \mathbf{p}_{h_k} + \sum_j \mathbf{s}_j^{h_k} \pmod q.$
2.  $\sum_j \mathbf{s}_{h_k}^j = 0 \pmod q.$
3.  $\mathbf{m}_{h_{k+1}} = \mathbf{p}_{h_{k+1}} + \sum_j \mathbf{s}_j^{h_{k+1}} \pmod q.$
4.  $\sum_j \mathbf{s}_{h_{k+1}}^j = 0 \pmod q.$

The free variables the simulator can use (since they have not been released to the adversary yet) are  $\mathbf{s}_{h_{k+1}}^{h_{k+1}}$ ,  $\mathbf{s}_{h_k}^{h_k}$ ,  $\mathbf{s}_{h_k}^{h_{k+1}}$ , and  $\mathbf{s}_{h_{k+1}}^{h_k}$ . The simulator proceeds to solve these 4 equations with respect to the 4 variables mentioned above.

We now turn to show that no environment can distinguish between the simulation of  $\mathcal{S}$  with  $\mathcal{F}_{\text{sfe-abort}}^f$  and the execution of  $\pi_{\text{allbutone}}$  with  $\mathcal{A}$  by defining a series of hybrid games. The output of each game is the output of the environment. Let  $\mathcal{Z}$  be a PPT environment.

**The game**  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^1$ . In this game, we modify the real-model experiment  $\text{REAL}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}$  as follows. In the second round, an arbitrary honest party  $P_{h_1}$  (e.g., the honest party  $P_i$  with smallest index  $i$ ) computes a simulated decryption share as  $\mathbf{p}_{h_1} \leftarrow \text{Sim}_{\text{TEFHE}}(h_1, \text{ct}, y, \{\mathbf{sk}_j\}_{j \neq h_1})$ .

Upon a corruption request of some party  $P_i$ , let  $h_k$  be the index of the recent party for which the decryption share was simulated (i.e.,  $h_k$  with the largest  $k$ ). If  $i = h_k$ , then proceed to explain the state of  $P_{h_k}$  in the same way as done in the simulation, by computing  $\mathbf{p}_{h_k} = \text{TEFHE.PartDec}(h_k, \mathbf{sk}_{h_k}, \text{ct}; r_{h_k}^D)$ , choosing another honest party  $P_{h_{k+1}}$  and simulating its partial decryption as  $\mathbf{p}_{h_{k+1}} \leftarrow \text{Sim}_{\text{TEFHE}}(h_{k+1}, \text{ct}, y, \{\mathbf{sk}_j\}_{j \neq h_{k+1}})$ . Finally, adjust  $\mathbf{s}_{h_{k+1}}^{h_{k+1}}$ ,  $\mathbf{s}_{h_k}^{h_k}$ ,  $\mathbf{s}_{h_k}^{h_{k+1}}$ , and  $\mathbf{s}_{h_{k+1}}^{h_k}$  by solving the four equations.

**Claim 6.7.**  $\text{REAL}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}} \stackrel{\text{s}}{\equiv} \text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^1$ .

*Proof.* This follows by the security of the threshold-decryption protocol of the TEFHE scheme. More precisely, since there are at most  $n - 1$  corruptions, we can list the set of parties that were used for simulating the decryption protocol as  $h_1, h_2, \dots, h_m$ , with  $m < n$ . Define the sequence of  $n + 1$  hybrids where in the  $i$ 'th hybrid, for  $1 \leq k \leq i - 1$  the decryption share of  $P_{h_k}$  is computed as  $\mathbf{p}_{h_k} \leftarrow \text{Sim}_{\text{TEFHE}}(h_k, \text{ct}, y, \{\mathbf{sk}_j\}_{j \neq h_k})$ , and for  $i \leq k \leq n$  as  $\mathbf{p}_{h_k} \leftarrow \text{TEFHE.PartDec}(h_k, \mathbf{sk}_{h_k}, \text{ct})$ . The first hybrid is exactly the execution of  $\pi_{\text{allbutone}}$  in the real world, whereas the  $(n + 1)$ 'th hybrid is the execution in  $\text{HYB}^1$ . By the security of the threshold decryption protocol it holds that every two neighboring hybrids are statistically close (in particular, in each hybrid running the final decryption on the decryption shares of all the parties gives the value  $y$ ), and the claim follows using a standard hybrid argument.  $\square$

**The game**  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^2$ . In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^1$  as follows. Instead of computing  $(\text{pk}, \mathbf{sk}_1, \dots, \mathbf{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$ , the threshold-PKI functionality sets the keys as  $(\text{pk}, \text{td}) \leftarrow \text{TEFHE.GenEquiv}(1^\kappa, 1^d)$ , for every  $i \in [n]$  samples  $\mathbf{s}_i \leftarrow \mathbb{Z}_q^{n'-1}$  (as per Definition 6.4), sets  $\mathbf{sk}_i = (\mathbf{s}_i, 1)$ , and gives  $(\text{pk}, \mathbf{sk}_i)$  to every party  $P_i$ .

**Claim 6.8.**  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^1 \stackrel{\text{c}}{\equiv} \text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^2$ .

*Proof.* In  $\text{HYB}^1$  the decryption shares are set such that the output is  $y$  for any set of secret keys. Therefore, the claim follows by the security of the threshold equivocal FHE scheme.  $\square$

**The game**  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^3$ . In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^2$  as follows. Instead of having each party encrypt its input as  $\text{ct}_i = \text{TEFHE.Enc}(\text{pk}, x_i; r_i^E)$ , each party  $P_i$  encrypts  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, 0^{\ell_{\text{in}}})$  and computes  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$ .

**Claim 6.9.**  $\text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^2 \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{allbutone}}, \mathcal{A}, \mathcal{Z}}^3$ .

*Proof.* In  $\text{HYB}^1$  and  $\text{HYB}^2$  the output value is forced to be  $y$  independently of the ciphertexts, so it is possible to replace the encryption of the parties one-by-one from the real inputs to encryptions of zero. The claim follows from the security of the TEFHE scheme.  $\square$

The execution in  $\text{HYB}^3$  is exactly like the simulation done by  $\mathcal{S}$  in the ideal model, therefore the theorem follows from Claims 6.7 to 6.9.  $\square$

## 6.4 Malicious Security

We now proceed to achieve security against malicious and adaptive adversaries. The underlying idea is to compile the semi-malicious protocol from Section 6.3 (Figure 5) to provide malicious security. Following the GMW paradigm, Asharov et al. [5, Thm. E.3] showed how to compile any semi-maliciously secure protocol into a maliciously secure protocol in the ZK-hybrid model that tolerates the same number of corruptions without increasing the round complexity. The compiler in [5] was analyzed in the static-corruption setting; however, it is not hard to see that if the semi-malicious protocol is adaptively secure, then so is the maliciously secure protocol.

It is important that the semi-malicious protocol will be defined purely over a broadcast channel, and that no point-to-point messages will be used. Therefore, the secret shares of zero that were sent over secure point-to-point channels will be encrypted and transmitted over the broadcast channel. As we consider adaptive corruptions, we need to use non-committing encryption (see Appendix A.2.5) instead of standard public-key encryption, and each non-committing public key should be used to encrypt  $n$  elements in  $\mathbb{Z}_q$ . We consider the distribution of the NCE public keys as part of the threshold-PKI functionality.<sup>11</sup>

Instantiating the ZK functionality using adaptively secure NIZK yields a maliciously secure protocol with the same number of rounds. The CRS for the NIZK can also be combined with the threshold-PKI functionality. Using the NIZK from Section 3, we achieve a two-round protocol in the threshold-PKI model with communication complexity that is independent of the function.

**Corollary 6.10.** *Assume the existence of special TEFHE schemes, NCE schemes, and the additional assumptions in Corollary 3.5, let  $t < n$ , and let  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  be an efficiently computable function. Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{bc}})$ -hybrid model, tolerating an adaptive, malicious, PPT  $t$ -adversary, by a two-round protocol with communication complexity  $\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$ .*

By instantiating the threshold-PKI functionality using an adaptively secure protocol, we get a protocol that is defined in the CRS-hybrid model. Using the recent protocol of Benhamouda et al. [13] we obtain the following corollary.

**Corollary 6.11.** *Assume the existence of two-round adaptively secure OT in the CRS model, and the additional assumptions in Corollary 6.10, let  $t < n$ , and let  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  be an*

<sup>11</sup>Note that unlike the TEFHE keys, the NCE keys can only be used an a priori bounded number of times.

efficiently computable function. Then,  $\mathcal{F}_{\text{sfe-abort}}^f$  can be UC-realized in the  $(\mathcal{F}_{\text{crs}}, \mathcal{F}_{\text{bc}})$ -hybrid model, tolerating an adaptive, malicious PPT  $t$ -adversary, by a four-round protocol with communication complexity  $\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$ .

## 7 The Honest-Majority Setting

In this section, we show how to adjust the protocol from Section 6 that provides security with abort, into guaranteeing output delivery in the honest-majority setting. Gordon et al. [58] constructed a two-round protocol in the threshold-PKI model that guarantees output delivery assuming an honest majority in the static-corruption setting and has communication complexity that is independent of the function. We apply some of the techniques from [58] on our adaptively secure protocol designed for the all-but-one setting, and achieve a matching result tolerating adaptive corruptions.

In the all-but-one case (Section 6) the secret decryption key was shared using additive secret sharing. As observed in [58], since the decryption of the GSW-based threshold FHE consists of linear operations, it is possible to use Shamir’s secret sharing [89] instead. The problem with a naïve use of this idea is that when the partial decryptions are reconstructed, each decryption share is multiplied by the Lagrange coefficient, and thus also the smudging noise. This will result in blowing up the noise and may end up with an incorrect decryption. Gordon et al. [58] overcame this problem by having each party secret share (using Shamir’s scheme) its smudging noise in the first round of the protocol, and parties added shares of the smudging noise of non-aborting parties in a way that is compatible with the decryption algorithm.<sup>12</sup>

### 7.1 Cryptographic Primitives used in the Protocol

#### 7.1.1 Threshold Secret Sharing Scheme

A key ingredient in the construction of the TEFHE scheme for the honest-majority setting is Shamir’s  $(t + 1)$ -out-of- $n$  secret sharing [89]. Given integers  $t < n$  and a prime  $q > n$ , Shamir’s scheme consists of two algorithms.

- $(\mathbf{s}_1, \dots, \mathbf{s}_n) \leftarrow \text{Share}(s)$  : to share a secret  $s \in \mathbb{Z}_q$ , choose a random polynomial  $p(x)$  over  $\mathbb{Z}_q$  of degree  $t$  conditioned on  $p(0) = s$ , and set the shares as  $\mathbf{s}_i = p(i)$  for every  $i \in [n]$ .
- $s' = \text{Recon}(\{\mathbf{s}_i\}_{i \in S})$  : given shares  $\{\mathbf{s}_i\}_{i \in S}$  for some subset  $S \subseteq [n]$  of size  $|S| \geq t + 1$ , compute  $s' = \sum_{j \in S} L_j(0) \cdot \mathbf{s}_j$ , where  $L_j(x)$  is the Lagrange polynomial with respect to the set  $S$

$$L_j(x) = \prod_{k \in S \setminus \{j\}} \frac{x - k}{j - k}.$$

Shamir’s secret sharing perfectly hides the secret given any subset of at most  $t$  shares, and, as shown above, any subset of  $t + 1$  shares enables an efficient reconstruction of the secret. For the remaining of the section we will consider  $t = \lceil n/2 \rceil - 1$ .

In the proof of protocol we will use an additional property of Shamir’s scheme, which allows to generate “new” shares for a “new” secret in a way that is compatible to less than  $t$  “old” shares.

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<sup>12</sup>Recently, Boneh et al. [15] showed that this problem can be overcome in a different way, by using a special secret sharing scheme that ensures the Lagrange coefficients are binary values.

- $(\tilde{s}_1, \dots, \tilde{s}_n) \leftarrow \text{Reshare}(s, \{s_i\}_{i \in S})$ : To re-share a secret  $s \in \mathbb{Z}_q$  given a set of shares  $\{s_i\}_{i \in S}$ , where  $|S| \leq t$ , choose a random polynomial  $p(x)$  over  $\mathbb{Z}_q$  of degree  $t$  conditioned on  $p(0) = s$  and  $p(i) = s_i$  for  $i \in S$ , and set the new shares as  $\tilde{s}_j = p(j)$  for every  $j \in [n]$ .

### 7.1.2 Honest-Majority Threshold Equivocal FHE

In Section 6, the threshold FHE was defined using an  $n$ -out-of- $n$  secret sharing of the decryption key. We now adjust the definition of TEFHE to the honest majority setting. The main adjustments are using an  $n/2$ -out-of- $n$  Shamir secret sharing of the decryption key, and separating the smudging noise from the encryption and decryption algorithms so it is external to the scheme. We still require though the ability to simulate the smudging noise in a similar way to Definition 6.3.

**Definition 7.1.** *A special  $n/2$ -out-of- $n$  TEFHE scheme is defined using the same syntax as in Definition 6.3, with the following adjustments:*

- *The key-generation algorithm  $(\text{pk}, \text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$  outputs a public key  $\text{pk}$  that defines a prime  $q$  and a smudging-noise bound  $B_{\text{smdg}}$ . In addition, the algorithm outputs Shamir shares  $(\text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{Share}(\mathbf{t})$  of a value  $\mathbf{t} = (\mathbf{s}, 1)$ , where  $\mathbf{s} \leftarrow \mathbb{Z}_q^{n'-1}$  for some  $n' = \text{poly}(\kappa, d)$ .*
- *The partial decryption algorithm  $\mathbf{p}_i \leftarrow \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$  operates by computing  $\mathbf{p}_i = \langle \text{ct}, \text{sk}_i \rangle \bmod q$ .*
- *The final decryption algorithm  $\text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_i\}_{i \in S})$ , for  $S \subseteq [n]$  of size  $n/2$ , executes Shamir's reconstruction, i.e., outputs  $\text{Recon}(\{\mathbf{p}_i\}_{i \in S})$ .*

We require the following properties:

1. **Correctness:** Consider the FHE scheme that is defined by setting the decryption key  $\text{sk} = \text{Recon}(\text{sk}_1, \dots, \text{sk}_n)$ , and where the decryption algorithm consists of choosing an arbitrary set  $S \subseteq [n]$ , sampling for every  $j \in S$  smudging noise  $e_j \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$ , and outputting  $\langle \text{sk}, \text{ct} \rangle + \sum_{j \in S} e_j$ . Then, this FHE scheme is a correct, compact, and secure equivocal FHE scheme for circuits of depth  $d$  (according to Definition 6.1).
2. **Simulatability of partial decryption:** There exists a PPT simulator  $\text{Sim}_{\text{TEFHE}}$  such that on input  $i \in [n]$ , and all decryption keys except of the  $i$ 'th one  $\{\text{sk}_j\}_{j \neq i}$ . The following distributions are statistically close:

$$\left\{ \begin{array}{l} \mathbf{p}_i + e_i \\ e_i \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}] \end{array} \middle| \begin{array}{l} \mathbf{p}_i \leftarrow \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}) \\ \mathbf{p}_i \leftarrow \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}) \end{array} \right\} \stackrel{s}{\equiv} \{ \mathbf{p}'_i \mid \mathbf{p}'_i \leftarrow \text{Sim}_{\text{TEFHE}}(i, \text{ct}, \mu, \{\text{sk}_j\}_{j \neq i}) \},$$

where the keys are set as  $(\text{pk}, \text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$ , the ciphertext is set as  $\text{ct} \leftarrow \text{TEFHE.Eval}(\text{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell)$  for a circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$  and for  $i \in [\ell]$  ciphertext  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, \mu_i)$  with  $\mu_i \in \{0, 1\}$ , and  $\mu = C(\mu_1, \dots, \mu_\ell)$ .

**Lemma 7.2.** *Assuming LWE there exist special  $n/2$ -out-of- $n$  TEFHE schemes.*

*Proof.* The lemma follows from Lemma 6.2 by sharing the secret key  $\text{sk}$  using Shamir's scheme, defining the partial decryption as the inner product of the ciphertext and the share of the key,

and the final decryption as the Shamir's reconstruction. Correctness follows by linearity of inner product. Simulatability of partial decryption follows identically as in Section 6.1.2. We note that the security of the scheme is not proven directly, but is combined with the security of the protocol below.  $\square$

## 7.2 Semi-Malicious Security

**Theorem 7.3.** *Assume the existence of special  $n/2$ -out-of- $n$  TEFHE schemes, let  $t < n/2$ , and let  $f : (\{0, 1\}^{\ell_{in}})^n \rightarrow \{0, 1\}^{\ell_{out}}$  be an efficiently computable function. Then,  $\mathcal{F}_{\text{sfe-god}}^f$  can be UC-realized in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model, tolerating an adaptive, semi-malicious, PPT  $t$ -adversary, by a two-round protocol with communication complexity  $\text{poly}(\ell_{in}, \ell_{out}, \kappa, n)$ .*

**Protocol  $\pi_{\text{hm}}$**

- **Private Input:** Every party  $P_i$ , for  $i \in [n]$ , has private input  $x_i \in \{0, 1\}^{\ell_{in}}$ .
- **Common Input:** A special  $n/2$ -out-of- $n$  TEFHE scheme  $\Pi$  and a circuit  $C_f$  of depth  $d$ .
- **The Protocol:**

1. Upon receiving  $(\text{input}, \text{sid}, x_i)$ , party  $P_i$  proceeds as follows:
  - (a) Invoke  $\mathcal{F}_{\text{thresh-pki}}(\Pi, d)$  with  $(\text{init}, \text{sid})$  to receive  $(\text{sid}, \text{pk}, \text{sk}_i)$ . Let  $q$  be the prime associated with  $\text{pk}$  and  $B_{\text{smdg}}$  the associated smudging-noise bound (as per Definition 7.1).
  - (b) Encrypt the input as  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, x_i)$ .
  - (c) Sample smudging noise  $e_i \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$ .
  - (d) Secret share  $e_i$  using Shamir's secret sharing as  $(s_{i,1}, \dots, s_{i,n}) \leftarrow \text{Share}(e_i)$ .
  - (e) Send  $(\text{sid}, \text{ct}_i, s_{i,j})$  to  $P_j$  over a secure channel (via  $\mathcal{F}_{\text{smt}}$ ).
2. Denote by  $S_1$  the set of parties who sent an  $(\text{sid}, \cdot)$  message, party  $P_i$  proceeds as follows:
  - (a) Let  $C_{f,S_1}$  be the circuit  $C_f$  where for every  $j \notin S_1$  a default input value is hard-wired.
  - (b) Compute  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_{f,S_1}, \{\text{ct}_j\}_{j \in S_1})$ .
  - (c) Partially decrypt the result as  $\text{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$ .
  - (d) Set  $\text{m}_i = \text{p}_i + \sum_{j \in S_1} s_{j,i}$  and send  $(\text{sid}, \text{m}_i)$  to every party.
3. Denote by  $S_2$  the set of parties who sent an  $(\text{sid}, \cdot)$  message. Party  $P_i$  computes the final decryption as  $y = \text{TEFHE.FinDec}(\text{pk}, \{\text{m}_j\}_{j \in S_2})$  and outputs  $(\text{output}, \text{sid}, y)$ .

Figure 6: MPC with g.o.d. and semi-malicious security in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model

*Proof.* We start by showing correctness. Denote by  $S_1$  the set of non-aborting parties in round 1, then for every  $i \in S_1$  it holds that  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, x_i)$  and  $(s_{i,1}, \dots, s_{i,n})$  are Shamir shares of  $e_i$ ; in addition,  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_{f,S_1}, \{\text{ct}_j\}_{j \in S_1})$ . Next, each party computes  $\text{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$  and sends  $\text{m}_i = \text{p}_i + \sum_{j \in S_1} s_{j,i} \pmod q$ . Denote by  $S_2$  the set of non-aborting parties in round 1. By linearity of inner product, and since  $|S_2| \geq n/2$  (due to the

honest-majority assumption), it holds that:

$$\begin{aligned}
\text{TEFHE.FinDec}(\text{pk}, \{\mathbf{m}_i\}_{i \in S_2}) &= \text{TEFHE.FinDec}(\text{pk}, \{\mathbf{p}_i + \sum_{j \in S_1} \mathbf{s}_{j,i}\}_{i \in S_2}) \\
&= \sum_{i \in S_2} L_i(0) \cdot (\mathbf{p}_i + \sum_{j \in S_1} \mathbf{s}_{j,i}) \\
&= \sum_{i \in S_2} L_i(0) \cdot (\langle \text{ct}, \mathbf{sk}_i \rangle + \sum_{j \in S_1} \mathbf{s}_{j,i}) \\
&= \langle \text{ct}, \sum_{i \in S_2} L_i(0) \cdot \mathbf{sk}_i \rangle + \sum_{j \in S_1} \sum_{i \in S_2} L_i(0) \cdot \mathbf{s}_{j,i} \\
&= \langle \text{ct}, \mathbf{sk} \rangle + \sum_{j \in S_1} e_j.
\end{aligned}$$

By *correctness* of the special  $n/2$ -out-of- $n$  TEFHE scheme, the decryption is correct and outputs  $C_{f, S_1}(\{x_i\}_{i \in S_1})$

We proceed to prove security of the protocol. Let  $\mathcal{A}$  be an adaptive, semi-malicious adversary attacking  $\pi_{\text{hm}}$  in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{smt}})$ -hybrid model. We will construct an ideal-process adversary  $\mathcal{S}$ , interacting with the ideal functionality  $\mathcal{F}_{\text{sfe-god}}^f$  and with ideal (dummy) parties  $\tilde{P}_1, \dots, \tilde{P}_n$ , such that no environment can distinguish between  $\mathcal{S}$  and  $\mathcal{A}$ . The simulator  $\mathcal{S}$  constructs virtual parties  $P_1, \dots, P_n$ , and runs the adversary  $\mathcal{A}$ . Let  $\text{Sim}_{\text{TEFHE}}$  be the simulator that is guaranteed to exist for the TEFHE scheme  $\Pi$ .

Note that the internal state of a party  $P_i$  in the protocol consists of its input  $x_i$ , the message  $(\text{pk}, \mathbf{sk}_i)$  from  $\mathcal{F}_{\text{thresh-pki}}$ , the random coins that include coins  $r_i^E$  for encrypting the input in the first round, coins  $r_i^{\text{sm}}$  for sampling the smudging noise  $e_i$ , and its secret sharing  $(\mathbf{s}_{i,1}, \dots, \mathbf{s}_{i,n})$  of  $e_i$ .<sup>13</sup> In addition, the internal state contains the messages  $P_i$  received in the first round, including the values  $(\text{ct}_1, \dots, \text{ct}_n)$  and  $(\mathbf{s}_{1,i}, \dots, \mathbf{s}_{n,i})$ , and in the second round  $(\mathbf{m}_1, \dots, \mathbf{m}_n)$ .

**Simulating the communication  $\mathcal{Z}$ :** Every input value that  $\mathcal{S}$  receives from  $\mathcal{Z}$  is written on  $\mathcal{A}$ 's input tape. Likewise, every output value written by  $\mathcal{A}$  on its output tape is copied to  $\mathcal{S}$ 's own output tape.

**Simulating the threshold-PKI functionality  $\mathcal{F}_{\text{thresh-pki}}$ :** The simulator  $\mathcal{S}$  computes  $(\text{pk}, \text{td}) \leftarrow \text{TEFHE.GenEquiv}(1^\kappa, 1^d)$ , samples  $\mathbf{s} \leftarrow \mathbb{Z}_q^{n'-1}$  (as per Definition 7.1), sets  $\mathbf{sk} = \mathbf{t} = (\mathbf{s}, 1)$ , computes Shamir shares  $(\text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{Share}(\mathbf{sk})$ , and gives  $(\text{pk}, \mathbf{sk}_i)$  to  $\mathcal{A}$  for every corrupted party  $P_i$ .

**Simulating the first round:** For every honest party  $P_i$ , compute  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, 0^{\ell_{\text{in}}})$ , sample  $e_i \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$ , share it as  $(\mathbf{s}_{i,1}^{(0)}, \dots, \mathbf{s}_{i,n}^{(0)}) \leftarrow \text{Share}(e_i)$ , and send  $(\text{sid}, \text{ct}_i, \mathbf{s}_{i,j}^{(0)})$  to every corrupted  $P_j$ . Finally, the simulator receives from  $\mathcal{A}$  the message  $(\text{sid}, \text{ct}_j, \mathbf{s}_{j,i}^{(0)})$  on behalf of every corrupted party  $P_j$  and every honest party  $P_i$ , and reads from the special witness tape of the semi-malicious adversary  $\mathcal{A}$  the input and random coins  $(x'_i, r_i)$ .

<sup>13</sup>Without loss of generality we can assume that  $(\mathbf{s}_i^1, \dots, \mathbf{s}_i^n)$  fully determine the coins used to sample  $e_i$  and to share it.



**Sending input to the functionality  $\mathcal{F}_{\text{sfe-god}}^f$ :** On behalf of every aborting corrupted party, set  $x'_i$  to be the default input value (e.g., zero), and for every non-aborting corrupted party  $P_i$  set the input value  $x'_i$  as read from the witness tape. Send the input values  $\{x'_i\}$  to the ideal functionality  $\mathcal{F}_{\text{sfe-god}}^f$  and receive back the output  $y$ .

**Simulating the second round:** Let  $\mathcal{I}$  be the set of currently corrupted parties, and let  $\{\text{ct}_j\}_{j \in S_1}$  be the ciphertexts of non-aborting parties in the first round of the simulation.  $\mathcal{S}$  homomorphically evaluates the circuit as  $\text{ct} = \text{TEFHE.Eval}(\text{pk}, C_{f, S_1}, \{\text{ct}_j\}_{j \in S_1})$  and computes for every honest party  $P_i$  the partial decryption  $\mathbf{p}_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct})$ .

Next, the simulator chooses an arbitrary honest party  $P_{h_1}$ , and computes the simulated partial decryption  $\tilde{\mathbf{p}}_{h_1} \leftarrow \text{Sim}_{\text{TEFHE}}(h_1, \text{ct}, y, \{\text{sk}_j\}_{j \neq h_1})$  (recall that the simulated  $\tilde{\mathbf{p}}_{h_1}$  includes the smudging noise).  $\mathcal{S}$  extracts the smudging noise for  $P_{h_1}$  as  $\tilde{e}_{h_1} = \mathbf{p}_{h_1} - \tilde{\mathbf{p}}_{h_1} \bmod q$  and re-shares it as  $(\mathbf{s}_{h_1,1}^{(1)}, \dots, \mathbf{s}_{h_1,n}^{(1)}) \leftarrow \text{Reshare}(\tilde{e}_{h_1}, \{\mathbf{s}_{h_1,j}^{(0)}\}_{j \in \mathcal{I}})$ . For every  $i \neq h_1$  and  $j$  set  $\mathbf{s}_{i,j}^{(1)} = \mathbf{s}_{i,j}^{(0)}$  (i.e., the shares sent to honest parties from corrupted parties in the first round (by  $\mathcal{A}$ ), and were sampled (by  $\mathcal{S}$ ) to be sent by honest parties, except for  $P_{h_1}$ ).

Finally, the simulator  $\mathcal{S}$  sets  $\mathbf{m}_i = \mathbf{p}_i + \sum_j \mathbf{s}_{j,i}^{(1)} \bmod q$ , gives  $(\text{sid}, \mathbf{m}_i)$  to  $\mathcal{A}$ , and receives messages from  $\mathcal{A}$  on behalf of corrupted parties.

**Simulating corruption requests of party  $P_i$ :** The simulator  $\mathcal{S}$  corrupts the dummy party  $\tilde{P}_i$  and continues as follows, depending on the timing of the corruption:

- **Corruptions before the first round:** In this case  $\tilde{P}_i$  has not been activated with input yet.  $\mathcal{S}$  sets the contents of  $P_i$ 's random tape with uniformly random coins.
- **Corruptions after the first round and before the second round:** In this case  $\tilde{P}_i$  has been activated with input  $(\text{input}, \text{sid}, x_i)$ , but did not receive output yet.  $\mathcal{S}$  sets the contents of  $P_i$ 's input tape to  $(\text{input}, \text{sid}, x_i)$  and the contents of the random tape to  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$  for the ciphertext  $\text{ct}_i$  that was used to simulate  $P_i$ 's message.

Finally, the secret shares of the smudging noise  $(\mathbf{s}_{i,1}^{(0)}, \dots, \mathbf{s}_{i,n}^{(0)})$  have already been set for  $P_i$  during the simulation. Similarly, the messages received in the first round, i.e.,  $(\mathbf{s}_{1,i}^{(0)}, \dots, \mathbf{s}_{n,i}^{(0)})$  are set according to the simulation both for corrupted parties (as received from  $\mathcal{A}$ ) and for honest parties (as simulated by  $\mathcal{S}$ ).

- **Corruptions after the second round:** In this case  $\tilde{P}_i$  has received the output.  $\mathcal{S}$  computes  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$  as in the previous case.

Let  $h_k$  be the index of the recent honest party that was ‘‘chosen’’ during the simulation; in particular, all of the secret shares are denoted as  $\mathbf{s}_{i,j}^{(k)}$ .

- If  $i \neq h_k$ , i.e., the corrupted party is *not* the ‘‘chosen’’ honest party, then set the shares of the smudging noise  $(\mathbf{s}_{i,1}^{(k)}, \dots, \mathbf{s}_{i,n}^{(k)})$ , and the incoming messages  $(\mathbf{s}_{1,i}^{(k)}, \dots, \mathbf{s}_{n,i}^{(k)})$  according to the values set in the simulation.
- If  $i = h_k$ , then the simulator needs to adjust the secret shares to be compatible with the view of the adversary. Let  $\mathcal{I}$  be the set of currently corrupted parties. To ‘‘fix’’ the shares sent by  $P_{h_k}$ , the simulator samples a new smudging noise  $e'_{h_k} \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  and re-shares it as  $(\mathbf{s}_{h_k,1}^{(k+1)}, \dots, \mathbf{s}_{h_k,n}^{(k+1)}) \leftarrow \text{Reshare}(e'_{h_k}, \{\mathbf{s}_{h_k,j}^{(k)}\}_{j \in \mathcal{I}})$ .

Next, to “fix” the incoming shares, the simulator chooses another honest party  $P_{h_{k+1}}$ , and sets the value

$$\mathbf{s}_{h_{k+1},h_k}^{(k+1)} = \sum_{j \in [n]} \mathbf{s}_{j,h_k}^{(1)} - \sum_{j \in [n] \setminus h_{k+1}} \mathbf{s}_{j,h_k}^{(k)} \pmod{q}.$$

For all other values  $i, j$  set  $\mathbf{s}_{i,j}^{(k+1)} = \mathbf{s}_{i,j}^{(k)}$ .

We now turn to show that no environment can distinguish between the simulation of  $\mathcal{S}$  with  $\mathcal{F}_{\text{sfe-god}}^f$  and the execution of  $\pi_{\text{allbutone}}$  with  $\mathcal{A}$  by defining a series of hybrid games. The output of each game is the output of the environment. Let  $\mathcal{Z}$  be a PPT environment.

**The game**  $\text{HYB}_{\pi_{\text{hm}},\mathcal{A},\mathcal{Z}}^1$ . In this game, we modify the real-model experiment  $\text{REAL}_{\pi_{\text{hm}},\mathcal{A},\mathcal{Z}}$  as follows.

Denote the shares sent from party  $P_i$  to party  $P_j$  by  $\mathbf{s}_{i,j}^{(0)}$ . After the first round, once the output value  $y$  is already determined by the inputs, an arbitrary honest party  $P_{h_1}$  (e.g., the honest party  $P_i$  with smallest index  $i$ ) replaces the smudging noise  $e_{h_1}$  used for the first round (and the shares  $(\mathbf{s}_{h_1,1}^{(0)}, \dots, \mathbf{s}_{h_1,n}^{(0)}) \leftarrow \text{Share}(e_{h_1})$ ) as follows:

Let  $\mathcal{I}$  be the set of currently corrupted parties. First, the simulated decryption share is computed as  $\tilde{\mathbf{p}}_{h_1} \leftarrow \text{Sim}_{\text{TEFHE}}(h_1, \text{ct}, y, \{\text{sk}_j\}_{j \neq h_1})$  (recall that the simulated  $\tilde{\mathbf{p}}_{h_1}$  includes the simulated smudging noise). Next, new smudging noise for  $P_{h_1}$  is extracted as  $\tilde{e}_{h_1} = \mathbf{p}_{h_1} - \tilde{\mathbf{p}}_{h_1} \pmod{q}$ , and re-shared as  $(\mathbf{s}_{h_1,1}^{(1)}, \dots, \mathbf{s}_{h_1,n}^{(1)}) \leftarrow \text{Reshare}(\tilde{e}_{h_1}, \{\mathbf{s}_{h_1,j}^{(0)}\}_{j \in \mathcal{I}})$ . For all other  $i, j$ , denote the shares  $\mathbf{s}_{i,j}^{(1)} = \mathbf{s}_{i,j}^{(0)}$ . The parties compute the second round messages using the “new” shares  $\mathbf{s}_{i,j}^{(1)}$ .

Upon a corruption request of party  $P_i$ , let  $h_k$  be the index of the recent honest party that was “chosen.” In particular, all of the secret shares are denoted as  $\mathbf{s}_{i,j}^{(k)}$ . If  $i \neq h_k$ , then the state of  $P_{h_k}$  is revealed to the adversary without any change. If  $i = h_k$ , then proceed to explain the state of  $P_{h_k}$  in the same way as done in the simulation, by adjusting the secret shares to be compatible with the view of the adversary. Let  $\mathcal{I}$  be the set of currently corrupted parties.

- To “fix” the shares sent by  $P_{h_k}$ , sample a new smudging noise  $e'_{h_k} \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  and re-share it as  $(\mathbf{s}_{h_k,1}^{(k+1)}, \dots, \mathbf{s}_{h_k,n}^{(k+1)}) \leftarrow \text{Reshare}(e'_{h_k}, \{\mathbf{s}_{h_k,j}^{(k)}\}_{j \in \mathcal{I}})$ .
- To “fix” the incoming shares, choose another honest party  $P_{h_{k+1}}$ , and sets the value

$$\mathbf{s}_{h_{k+1},h_k}^{(k+1)} = \sum_{j \in [n]} \mathbf{s}_{j,h_k}^{(1)} - \sum_{j \in [n] \setminus h_{k+1}} \mathbf{s}_{j,h_k}^{(k)} \pmod{q}.$$

- For all other values  $i, j$  set  $\mathbf{s}_{i,j}^{(k+1)} = \mathbf{s}_{i,j}^{(k)}$ .

**Claim 7.4.**  $\text{REAL}_{\pi_{\text{hm}},\mathcal{A},\mathcal{Z}} \stackrel{\text{S}}{\equiv} \text{HYB}_{\pi_{\text{hm}},\mathcal{A},\mathcal{Z}}^1$ .

*Proof.* This follows by the security of the threshold-decryption protocol of the TEFHE scheme and by the honest-majority assumption. Note that the view of the adversary at the second round contains the messages  $m_i = \text{TEFHE.PartDec}(\text{sk}_i, \text{ct}) + \sum_{j \in \mathcal{S}_1} \mathbf{s}_{j,i}^{(1)}$  sent by each honest party  $P_i$ , as well as all share sent to a corrupted  $P_j$  i.e.,  $\mathbf{s}_{i,j}^{(1)}$  and from every corrupted party  $\mathbf{s}_{j,i}^{(1)}$ . By the *simulatability of partial decryption* property, it holds that the view of the adversary is statistically close in the real execution and the hybrid experiment. The rest of the simulation ensures the following invariants are retained under a corruption of a new party  $P_i$ :

- $(s_{i,1}^{(k)}, \dots, s_{i,n}^{(k)})$  form a Shamir sharing of a uniformly distributed element  $e_i \in [-B_{\text{smdg}}, B_{\text{smdg}}]$ .
- $\sum_{j \in S_1} s_{i,n}^{(k)} = \sum_{j \in S_1} s_{i,n}^{(1)}$ , therefore  $m_i = \text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}) + \sum_{j \in S_1} s_{i,n}^{(1)}$ .
- For every honest party  $P_h$ , the adversary learns at most  $t$  shares  $s_{h,j}^{(k)}$ .

□

**The game**  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^2$ . In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^1$  as follows. Instead of computing  $(\text{pk}, \text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{TEFHE.Gen}(1^\kappa, 1^d)$ , the threshold-PKI functionality sets the keys as  $(\text{pk}, \text{td}) \leftarrow \text{TEFHE.GenEquiv}(1^\kappa, 1^d)$ , for every  $i \in [n]$  samples  $\mathbf{s} \leftarrow \mathbb{Z}_q^{n'-1}$  (as per Definition 7.1), sets  $\text{sk} = \mathbf{t} = (\mathbf{s}, 1)$ , computes  $(\text{sk}_1, \dots, \text{sk}_n) \leftarrow \text{Share}(\text{sk})$ , and gives  $(\text{pk}, \text{sk}_i)$  to every party  $P_i$ .

**Claim 7.5.**  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^1 \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^2$ .

*Proof.* In  $\text{HYB}^1$  the decryption shares are set such that the output is  $y$  for any set of secret keys. Therefore, the claim follows by the security of the threshold equivocal FHE scheme. □

**The game**  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^3$ . In this game, we modify the experiment  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^2$  as follows. Instead of having each party encrypt its input as  $\text{ct}_i = \text{TEFHE.Enc}(\text{pk}, x_i; r_i^E)$ , each party  $P_i$  encrypts  $\text{ct}_i \leftarrow \text{TEFHE.Enc}(\text{pk}, 0^{\ell_{\text{in}}})$  and computes  $r_i^E \leftarrow \text{TEFHE.Equiv}(\text{td}, \text{ct}_i, x_i)$ .

**Claim 7.6.**  $\text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^2 \stackrel{c}{\equiv} \text{HYB}_{\pi_{\text{hm}}, \mathcal{A}, \mathcal{Z}}^3$ .

*Proof.* In  $\text{HYB}^1$  and  $\text{HYB}^2$  the output value is forced to be  $y$  independently of the ciphertexts, so it is possible to replace the encryption of the parties one-by-one from the real inputs to encryptions of zero. The claim follows from the security of the TEFHE scheme. □

The execution in  $\text{HYB}^3$  is exactly like the simulation done by  $\mathcal{S}$  in the ideal model, therefore the theorem follows from Claims 7.4 to 7.6. □

### 7.3 Malicious Security

We proceed to prove security against malicious and adaptive adversaries in an analogue way to Section 6.4, using the compiler from semi-malicious security to malicious security [5]. As in the all-but-one case, we require all communication to be sent over a broadcast channel, hence the secret shares of the smudging noise should be encrypted using non-committing encryption (see Appendix A.2.5). We consider the distribution of the NCE public keys as part of the threshold-PKI functionality. Similarly, the CRS for the adaptively secure NIZK from Section 3 is also considered as part of public key delivered by the threshold-PKI functionality.

**Theorem 7.7.** *Assume the existence of special  $n/2$ -out-of- $n$  TEFHE schemes and of NCE schemes, let  $t < n/2$ , and let  $f : (\{0, 1\}^{\ell_{\text{in}}})^n \rightarrow \{0, 1\}^{\ell_{\text{out}}}$  be an efficiently computable function. Then,  $\mathcal{F}_{\text{sfe-god}}^f$  can be UC-realized in the  $(\mathcal{F}_{\text{thresh-pki}}, \mathcal{F}_{\text{bc}})$ -hybrid model, tolerating an adaptive, malicious PPT  $t$ -adversary, by a two-round protocol with communication complexity  $\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$ .*

By instantiating the threshold-PKI functionality using the adaptively secure protocol of Damgård and Ishai [39], we get a protocol that is defined in the plain model.

**Corollary 7.8.** *Consider the same assumptions as in Theorem 7.7. Then,  $\mathcal{F}_{\text{sfe-god}}^f$  can be UC-realized in the  $\mathcal{F}_{\text{bc}}$ -hybrid model, tolerating an adaptive, malicious PPT  $t$ -adversary, by a constant-round protocol with communication complexity  $\text{poly}(\ell_{\text{in}}, \ell_{\text{out}}, \kappa, n)$ .*

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## A Preliminaries (Cont’d)

In Appendix [A.1](#), we define the LWE and adaptive LWE assumptions. In Appendix [A.2](#), we formally define the cryptographic primitives used in the paper.

### A.1 Cryptographic Assumptions

#### A.1.1 Learning With Errors

The decisional learning with errors (LWE) problem, introduced by Regev [\[86\]](#), is defined as follows.

**Definition A.1** (decision LWE). Let  $n = n(\kappa)$  and  $q = q(\kappa)$  be integer parameters and  $\chi = \chi(\kappa)$  be a distribution over  $\mathbb{Z}$ . The learning with errors (LWE) assumption  $\text{LWE}_{n,q,\chi}$  states that for all polynomials  $m = \text{poly}(\kappa)$  the following distributions are computationally indistinguishable:

$$(\mathbf{A}, \mathbf{s}^T \mathbf{A} + \mathbf{e}) \stackrel{c}{\equiv} (\mathbf{A}, \mathbf{u}),$$

where  $\mathbf{A} \leftarrow \mathbb{Z}_q^{n \times m}$ ,  $\mathbf{s} \leftarrow \mathbb{Z}_q^n$ ,  $\mathbf{e} \leftarrow \chi^m$ , and  $\mathbf{u} \leftarrow \mathbb{Z}_q^m$ .

We rely on LWE security with the following range of parameters. We assume that for any polynomial  $p = p(\kappa) = \text{poly}(\kappa)$  there exists some polynomial  $n = n(\kappa) = \text{poly}(\kappa)$ , some  $q = q(\kappa) = 2^{\text{poly}(\kappa)}$ , and some  $B = B(\kappa)$ -bounded distribution  $\chi = \chi(\kappa)$  such that  $q/B \geq 2^p$  and the  $\text{LWE}_{n,q,\chi}$  assumption holds. Throughout the paper, the LWE assumption without further specification refers to the above parameters. The sub-exponentially secure LWE assumption further assumes that  $\text{LWE}_{n,q,\chi}$  with the above parameters is sub-exponentially secure, meaning that there exists some  $\epsilon > 0$  such that the distinguishing advantage of any polynomial-time distinguisher is  $2^{-\kappa^\epsilon}$ .

### A.1.2 Adaptive Learning With Errors

Quach et al. [84] used the following natural variant of the LWE problem, denoted *adaptive LWE*.

**Definition A.2** (decision ALWE). We define the decision adaptive LWE assumption  $\text{ALWE}_{n,k,q,\chi}$  with parameter  $n, k, q \in \mathbb{Z}$  and a distribution  $\chi$  over  $\mathbb{Z}$  which are all parametrized by the security parameter  $\kappa$ . Let  $m = n \cdot \lceil \log q \rceil$ . We let  $\mathbf{G} \in \mathbb{Z}^{n \times m}$  be the gadget matrix (as defined in [77], see also Appendix C.1). For any polynomial  $m' = m'(\kappa)$ , we consider the following two games  $\text{GAME}^\beta$  for  $\beta \in \{0, 1\}$  between a challenger and an adversary  $\mathcal{A}$ .

- The Challenger picks  $k$  random matrices  $\mathbf{A}_i \leftarrow \mathbb{Z}_q^{n \times m}$  for  $i \in [k]$ , and sends them to  $\mathcal{A}$ .
- $\mathcal{A}$  adaptively picks  $x_1, \dots, x_k \in \{0, 1\}$ , and sends it to the Challenger.
- The Challenger samples  $\mathbf{s} \leftarrow \mathbb{Z}_q^n$  and computes for all  $i \in [k]$

$$\begin{cases} \mathbf{b}_i = \mathbf{s}^T (\mathbf{A}_i - x_i \cdot \mathbf{G}) + \mathbf{e}_i \text{ where } \mathbf{e}_i \leftarrow \chi^m, & \text{if } \beta = 0. \\ \mathbf{b}_i \leftarrow \mathbb{Z}_q^m, & \text{if } \beta = 1. \end{cases}$$

The Challenger also picks  $\mathbf{A}_{k+1} \leftarrow \mathbb{Z}_q^{n \times m'}$  and computes

$$\begin{cases} \mathbf{b}_{k+1} = \mathbf{s}^T \mathbf{A}_{k+1} + \mathbf{e}_{k+1} \text{ where } \mathbf{e}_{k+1} \leftarrow \chi^{m'}, & \text{if } \beta = 0. \\ \mathbf{b}_{k+1} \leftarrow \mathbb{Z}_q^{m'}, & \text{if } \beta = 1. \end{cases}$$

The challenger sends  $\mathbf{A}_{k+1}$  and  $\{\mathbf{b}_i\}_{i \in [k+1]}$  to the adversary.

The  $\text{ALWE}_{n,k,q,\chi}$  assumption states that for all polynomial  $m = m(\kappa)$ , the games  $\text{GAME}^0$  and  $\text{GAME}^1$  are computationally indistinguishable.

## A.2 Cryptographic Primitives

### A.2.1 Laconic Function Evaluation

Quach et al. [84] introduced the notion of laconic function evaluation (LFE), and constructed an LFE scheme for all circuits under the adaptive LWE assumption. At a high level, LFE allows to “compress” a function into a short digest that enables encrypting the input to the function as a short ciphertext. The size of the digest and the ciphertext (and therefore the computational cost of the encryption algorithm) only depend on the depth of the circuit representing the function, and are independent of the circuit size.

Quach et al. [84] considered another variant of LFE that is *function-hiding*, meaning that the digest hides all partial information about the underlying function. Following [84], we assume that the circuit class  $\mathcal{C}$  associates every circuit  $C \in \mathcal{C}$  with some circuit parameters  $C.\text{params}$  that remain unhidden. Unless specified otherwise, we will consider  $\mathcal{C}$  to be the class of all circuits with  $C.\text{params} = (1^{\ell_{\text{in}}}, 1^d)$  consisting of the input size  $\ell_{\text{in}}$  and the depth  $d$  of the circuit.

**Definition A.3** (LFE). *A laconic function evaluation (LFE) scheme for a class of circuits  $\mathcal{C}$  consists of four algorithm  $\Pi = (\text{LFE.crsGen}, \text{LFE.Compress}, \text{LFE.Enc}, \text{LFE.Dec})$ .*

- $\text{LFE.crsGen}(1^\kappa, \text{params}) \rightarrow \text{crs}$ : *given an input the security parameter and circuit parameters  $\text{params}$  the CRS generation algorithm outputs a uniformly random common random string  $\text{crs}$  of appropriate length.*
- $\text{LFE.Compress}(\text{crs}, C) \rightarrow \text{digest}_C$ : *given as input the common random string  $\text{crs}$  and a circuit  $C \in \mathcal{C}$ , the compression algorithm outputs a digest  $\text{digest}_C$ .*
- $\text{LFE.Enc}(\text{crs}, \text{digest}_C, x) \rightarrow \text{ct}$ : *given as input the common random string  $\text{crs}$ , a digest  $\text{digest}_C$ , and a message  $x$ , the encryption algorithm outputs a ciphertext  $\text{ct}$ .*
- $\text{LFE.Dec}(\text{crs}, C, r, \text{ct}) \rightarrow y$ : *given as input the common random string  $\text{crs}$ , a circuit  $C \in \mathcal{C}$ , the compression random coins  $r$ , and a ciphertext  $\text{ct}$ , the deterministic decryption algorithm outputs a message  $y$ .*

We require the following properties from an LFE scheme  $\Pi$ :

**Correctness.** For every security parameter  $\kappa$ , parameters  $\text{params}$ , and circuit  $C \in \mathcal{C}$  with  $C.\text{params} = \text{params}$  it holds that:

$$\Pr \left[ y = C(x) \mid \begin{array}{l} \text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params}) \\ \text{digest}_C = \text{LFE.Compress}(\text{crs}, C; r) \\ \text{ct} \leftarrow \text{LFE.Enc}(\text{crs}, \text{digest}_C, x) \\ y = \text{LFE.Dec}(\text{crs}, C, r, \text{ct}) \end{array} \right] = 1.$$

**Security.** There exists a PPT simulator  $\text{Sim}_{\text{LFE}}$  for the scheme  $\Pi$  such that for every stateful PPT adversary  $\mathcal{A}$ , it holds that

$$\left| \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{LFE-real}}(\kappa) = 1 \right] - \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{LFE-ideal}}(\kappa) = 1 \right] \right| \leq \text{negl}(\kappa),$$

for the experiments  $\text{Expt}^{\text{LFE-real}}$  and  $\text{Expt}^{\text{LFE-ideal}}$  defined below:

$\text{Expt}_{\Pi, \mathcal{A}}^{\text{LFE-real}}(\kappa)$	$\text{Expt}_{\Pi, \mathcal{A}}^{\text{LFE-ideal}}(\kappa)$
$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ $x^*, C, r \leftarrow \mathcal{A}(\text{crs})$ s.t. $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $\text{digest}_C = \text{LFE.Compress}(\text{crs}, C; r)$ $\text{ct} \leftarrow \text{LFE.Enc}(\text{crs}, \text{digest}_C, x^*)$ Output $\mathcal{A}(\text{ct})$	$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ $x^*, C, r \leftarrow \mathcal{A}(\text{crs})$ s.t. $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $\text{digest}_C = \text{LFE.Compress}(\text{crs}, C; r)$ $\text{ct} \leftarrow \text{Sim}_{\text{LFE}}(\text{crs}, C, r, \text{digest}_C, C(x^*))$ Output $\mathcal{A}(\text{ct})$

**(Statistical) Function-Hiding.** An LFE scheme  $\Pi$  is function hiding if there exists a PPT simulator  $\text{Sim}_{\text{FH}}$  for the scheme  $\Pi$  such that for all stateful PPT adversary  $\mathcal{A}$ , it holds that

$$\left| \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{FH-real}}(\kappa) = 1 \right] - \Pr \left[ \text{Expt}_{\Pi, \mathcal{A}}^{\text{FH-ideal}}(\kappa) = 1 \right] \right| \leq \text{negl}(\kappa),$$

for the experiments  $\text{Expt}^{\text{FH-real}}$  and  $\text{Expt}^{\text{FH-ideal}}$  defined below:

$\text{Expt}_{\Pi, \mathcal{A}}^{\text{FH-real}}(\kappa)$	$\text{Expt}_{\Pi, \mathcal{A}}^{\text{FH-ideal}}(\kappa)$
$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ $C \leftarrow \mathcal{A}(\text{crs})$ s.t. $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $\text{digest}_C \leftarrow \text{LFE.Compress}(\text{crs}, C)$ Output $\mathcal{A}(\text{digest}_C)$	$\text{params} \leftarrow \mathcal{A}(1^\kappa)$ $\text{crs} \leftarrow \text{LFE.crsGen}(1^\kappa, \text{params})$ $C \leftarrow \mathcal{A}(\text{crs})$ s.t. $C \in \mathcal{C}$ and $C.\text{params} = \text{params}$ $\text{digest}_C \leftarrow \text{Sim}_{\text{FH}}(\text{crs}, C.\text{params})$ Output $\mathcal{A}(\text{digest}_C)$

**Theorem A.4** ([84]). *Assuming sub-exponential hardness of LWE there exists a function-hiding LFE scheme for circuits  $C$  of depth  $d$ , input size  $\ell_{in}$ , and output size  $\ell_{out}$  where:*

- *The CRS is of size  $\text{poly}(\kappa, \ell_{in}, d)$ . The digest is of size  $\text{poly}(\kappa)$ .*
- *The running time of the encryption algorithm and the size of its output (the ciphertext) are  $\tilde{O}(\ell_{out}) \cdot \text{poly}(\kappa, \ell_{in}, d)$ .*
- *The running time of the compression and the decryption algorithms is  $\tilde{O}(|C|) \cdot \text{poly}(\kappa, \ell_{in}, d)$ .*
- *If the function-hiding property is not required, the compression algorithm  $\text{LFE.Compress}$  can be made deterministic.*

## A.2.2 Explainability Compiler

The notion of an explainability compiler was introduced by Dachman-Soled et al. [38] in the context of adaptively secure MPC as an extension of the technique of Sahai and Waters [88] for constructing sender-deniable encryption. At a high-level, the compiler can take any randomized algorithm  $\text{Alg}$  and produce an algorithm  $\widetilde{\text{Alg}}$  with the same functionality and has roughly the same size, along with an Explain algorithm. For any input/output pair  $(x, y)$ , the Explain algorithm can produce coins  $r$  such that  $y = \widetilde{\text{Alg}}(x; r)$ . A similar notion was also used by Canetti et al. [25].

**Definition A.5** (explainability compiler). *A PPT algorithm  $\text{Comp}$  is an explainability compiler for a circuit class  $\mathcal{C}$  if for every efficient, randomized circuit  $\text{Alg} \in \mathcal{C}$ , the following hold:*

- **Polynomial slowdown.** *There is a polynomial  $p(\cdot)$  such that, for any  $(\widetilde{\text{Alg}}, \text{Explain})$  output by  $\text{Comp}(1^\kappa, \text{Alg})$  it holds that  $|\widetilde{\text{Alg}}| \leq p(\kappa) \cdot |\text{Alg}|$ .*

- **Statistical functional equivalence.** With overwhelming probability over the choice of  $(\widetilde{\text{Alg}}, \cdot)$  as output by  $\text{Comp}(1^\kappa, \text{Alg})$ , the distribution of  $\widetilde{\text{Alg}}(x)$  is statistically close to the distribution of  $\text{Alg}(x)$  for every input  $x$ .
- **Explainability.** For every stateful PPT adversary  $\mathcal{A}$  it holds that

$$\left| \Pr \left[ \text{Expt}_{\text{Comp}, \text{Alg}, \mathcal{A}}^{\text{Explain}}(\kappa) = 1 \right] \right| \leq 1/2 + \text{negl}(\kappa),$$

for the experiment  $\text{Expt}^{\text{Explain}}$  defined below:

$\text{Expt}_{\text{Comp}, \text{Alg}, \mathcal{A}}^{\text{Explain}}(\kappa)$
$x^* \leftarrow \mathcal{A}(1^\kappa)$ $(\text{Alg}, \text{Explain}) \leftarrow \text{Comp}(1^\kappa, \text{Alg})$ Sample $r_0 \leftarrow \{0, 1\}^*$ Compute $y^* = \text{Alg}(x^*; r_0)$ Compute $r_1 \leftarrow \text{Explain}(x^*, y^*)$ Sample $b \leftarrow \{0, 1\}$ Compute $b' \leftarrow \mathcal{A}(\text{Alg}, y^*, r_b)$ Output 1 if and only if $b' = b$

**Theorem A.6** ([38]). Assuming the existence of an indistinguishable obfuscator for P/poly and of one-way functions, there exists an explainability compiler for P/poly.

### A.2.3 Homomorphic Trapdoor Functions

Homomorphic trapdoor functions (HTDF) were introduced by [57] as a unification of homomorphic encryption and homomorphic signatures.

**Definition A.7** (HTDF [57]). A homomorphic trapdoor function (HTDF) consists of the following five polynomial-time algorithms ( $\text{HTDF.Gen}$ ,  $f$ ,  $\text{HTDF.Inv}$ ,  $\text{HTDF.Eval}^{\text{in}}$ ,  $\text{HTDF.Eval}^{\text{out}}$ ) with the following syntax:

- $\text{HTDF.Gen}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{sk})$ : given an input the security parameter and the depth-bound, the key-generation procedure outputs a public key  $\text{pk}$  and a secret key  $\text{sk}$ . The security parameter defines the index space  $\mathcal{X}$ , the input space  $\mathcal{U}$ , the output space  $\mathcal{V}$  and some efficiently samplable input distribution  $D_{\mathcal{U}}$  over  $\mathcal{U}$ . We require that membership in the sets  $\mathcal{U}, \mathcal{V}, \mathcal{X}$  can be efficiently tested and that one can efficiently sample uniformly at random from  $\mathcal{V}$ .
- $f_{\text{pk}, x} : \mathcal{U} \rightarrow \mathcal{V}$ : the algorithm  $f$  is parametrized by a public key  $\text{pk}$  and an index  $x \in \mathcal{X}$ .
- $\text{HTDF.Inv}_{\text{sk}, x} : \mathcal{V} \rightarrow \mathcal{U}$ : the algorithm  $\text{HTDF.Inv}$  is parametrized by a secret key  $\text{pk}$  and an index  $x \in \mathcal{X}$ .
- $u^* = \text{HTDF.Eval}^{\text{in}}(g, (x_1, u_1), \dots, (x_\ell, u_\ell))$  and  $v^* = \text{HTDF.Eval}^{\text{out}}(g, v_1, \dots, v_\ell)$  are deterministic input/output homomorphic evaluation algorithms. The algorithms take as input some function  $g : \mathcal{X}^\ell \rightarrow \mathcal{X}$  and values  $x_i \in \mathcal{X}, u_i \in \mathcal{U}, v_i \in \mathcal{V}$ . The outputs are  $u^* \in \mathcal{U}$  and  $v^* \in \mathcal{V}$ .

We require the following properties from an HTDF scheme:

- **Correctness.** Let  $(\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d)$ , let  $x_1, \dots, x_\ell \in \mathcal{X}$ , let  $g : \mathcal{X}^\ell \rightarrow \mathcal{X}$  of depth at most  $d$ , and let  $y := g(x_1, \dots, x_\ell)$ . Let  $u_1, \dots, u_\ell \in \mathcal{U}$  and set  $v_i := f_{\text{pk}, x_i}(u_i)$  for  $i \in [\ell]$ . Let  $u^* = \text{HTDF.Eval}^{\text{in}}(g, (x_1, u_1), \dots, (x_\ell, u_\ell))$  and let  $v^* = \text{HTDF.Eval}^{\text{out}}(g, v_1, \dots, v_\ell)$ . Then, we require that  $u^* \in \mathcal{U}$  and  $f_{\text{pk}, y}(u^*) = v^*$ .
- **Distributional equivalence of inversion.** The following distributions are statistically close:

$$\left\{ (\text{pk}, \text{sk}, x, u, v) \mid (\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d), x \in \mathcal{X}, u \leftarrow \mathcal{U}, v = f_{\text{pk}, x}(u) \right\}_\kappa$$

$$\stackrel{s}{\equiv}$$

$$\left\{ (\text{pk}, \text{sk}, x, u', v') \mid (\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d), x \in \mathcal{X}, v' \leftarrow \mathcal{V}, u' = \text{HTDF.Inv}_{\text{sk}, x}(v') \right\}_\kappa,$$

where  $x \in \mathcal{X}$  is an arbitrary random variable that depends on  $(\text{pk}, \text{sk})$ .

- **Claw-free security.** For every PPT adversary  $\mathcal{A}$  it holds that

$$\Pr \left[ \begin{array}{c} f_{\text{pk}, x}(u) = f_{\text{pk}, x'}(u') \\ u, u' \in \mathcal{U}, \quad x, x' \in \mathcal{X}, \quad x \neq x' \end{array} \middle| \begin{array}{c} (\text{pk}, \text{sk}) \leftarrow \text{HTDF.Gen}(1^\kappa, 1^d) \\ (x, x', u, u') \leftarrow \mathcal{A}(1^\kappa, \text{pk}) \end{array} \right] \leq \text{negl}(\kappa).$$

**Theorem A.8** ([57]). Under the LWE assumption, there exists an HTDF scheme.

#### A.2.4 Strong One-Time Signatures

**Definition A.9** (strong one-time signatures). A *strong one-time signatures scheme* consists of three polynomial-time algorithms  $(\text{Sig.Gen}, \text{Sign}, \text{Vrfy})$  with the following syntax:

- $\text{Sig.Gen}(1^\kappa) \rightarrow (\text{vk}, \text{sigk})$ : given an input the security parameter, the key-generation procedure outputs a public verification key  $\text{vk}$  and a secret signing key  $\text{sigk}$ .
- $\text{Sign}_{\text{sigk}}(m) \rightarrow \sigma$ : given an input a signing key  $\text{sigk}$  and a message  $m$ , the signing algorithm outputs a signature  $\sigma$ .
- $b = \text{Vrfy}_{\text{vk}}(\sigma, m)$ : given an input a verification key  $\text{vk}$ , a signature  $\sigma$ , and a message  $m$ , the verification algorithm outputs a bit  $b \in \{0, 1\}$ .

We require the following properties from the signature scheme:

- **(Perfect) correctness.** For every message  $m \in \{0, 1\}^*$  it holds that

$$\Pr \left[ \text{Vrfy}_{\text{vk}}(\sigma, m) = 1 \mid (\text{vk}, \text{sigk}) \leftarrow \text{Sig.Gen}(1^\kappa), \sigma \leftarrow \text{Sign}_{\text{sigk}}(m) \right] = 1.$$

- **Strong existential unforgeability under one-time chosen message attack.** For every PPT adversary  $\mathcal{A}$  it holds that

$$\Pr \left[ \begin{array}{c} (m', \sigma') \neq (m, \sigma) \\ \text{Vrfy}_{\text{vk}}(\sigma', m') = 1 \end{array} \middle| \begin{array}{c} (\text{vk}, \text{sigk}) \leftarrow \text{Sig.Gen}(1^\kappa) \\ m \leftarrow \mathcal{A}(\text{vk}) \\ \sigma \leftarrow \text{Sign}_{\text{sigk}}(m) \\ (m', \sigma') \leftarrow \mathcal{A}(\sigma) \end{array} \right] \leq \text{negl}(\kappa).$$



We will use a strong one-time signature scheme that has fixed-length signatures, i.e., where there is a polynomial upper bound  $\ell_{SIG}(\kappa)$  on the length of the signatures. Fixed-length strong one-time signatures can be constructed from one-way functions (from universal one-way hash-functions [79] and Lamport signatures used in combination with Merkle trees [87]). The existence of fully homomorphic trapdoor functions therefore implies the existence of fixed-length strong one-time signatures.

### A.2.5 Non-Committing Encryption

A non-committing encryption scheme [21, 9, 40] is a public-key encryption scheme with the capability to efficiently simulate a public key and a ciphertext that can be explained as an encryption of any message.

**Definition A.10** (NCE). *A non-non-committing (bit) encryption scheme consists of four algorithms (NC.Gen, NC.Enc, NC.Dec, NC.Sim) such that the following properties hold:*

- *The triplet (NC.Gen, NC.Enc, NC.Dec) forms a public-key encryption scheme.*
- *NC.Sim is a simulation algorithm that on input  $1^\kappa$ , outputs  $(\text{pk}, \text{ct}, \rho_G^0, \rho_E^0, \rho_G^1, \rho_E^1)$ , such that for any  $\mu \in \{0, 1\}$  the following distributions are computationally indistinguishable:*

– *the joint view of an honest sender and an honest receiver in a normal encryption of  $\mu$*

$$\{(\text{pk}, \text{ct}, r_G, r_E) \mid (\text{sk}, \text{pk}) = \text{NC.Gen}(1^\kappa; r_G), \text{ct} = \text{NC.Enc}(\text{pk}, \mu; r_E)\},$$

– *the simulated view of an encryption of  $\mu$*

$$\{(\text{pk}, \text{ct}, \rho_G^\mu, \rho_E^\mu) \mid (\text{pk}, \text{ct}, \rho_G^0, \rho_E^0, \rho_G^1, \rho_E^1) \leftarrow \text{NC.Sim}(1^\kappa)\}.$$

NCE schemes exist under wide range of assumptions, including LWE [31, 82, 66, 67].

### A.2.6 Fully Homomorphic Encryption

**Definition A.11.** *A (leveled) fully homomorphic encryption scheme (FHE) consists of 4 PPT algorithms:*

- *FHE.Gen( $1^\kappa, 1^d$ )  $\rightarrow$  (pk, sk): on input the security parameter  $\kappa$  and a depth bound  $d$ , the key generation algorithm outputs a public key  $\text{pk}$  and a secret key  $\text{sk}$ .*
- *FHE.Enc(pk,  $\mu$ )  $\rightarrow$  ct: on input a public key  $\text{pk}$  and a plaintext  $\mu \in \{0, 1\}$ , the encryption algorithm outputs a ciphertext  $\text{ct}$ .*
- *FHE.Eval(pk,  $C, \text{ct}_1, \dots, \text{ct}_\ell$ )  $\rightarrow$  ct: on input a public key  $\text{pk}$ , a circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$ , and a tuple of ciphertexts  $(\text{ct}_1, \dots, \text{ct}_\ell)$ , the homomorphic evaluation algorithm outputs a ciphertext  $\text{ct}$ .*
- *FHE.Dec(sk, ct)  $\rightarrow$   $\tilde{\mu}$ : on input a secret key  $\text{sk}$  and a ciphertext  $\text{ct}$ , the decryption algorithm outputs  $\tilde{\mu} \in \{0, 1\}$ .*

We require the FHE scheme to be *correct*, meaning that when initialized with depth bound  $d$  the scheme correctly evaluates all circuits of depth at most  $d$ , and *compact*, meaning that the size of the decryption circuit (and of the evaluated ciphertext) is independent of  $d$ .

**Definition A.12.** Let  $\Pi = (\text{FHE.Gen}, \text{FHE.Enc}, \text{FHE.Dec}, \text{FHE.Eval})$  be an FHE scheme.

- $\Pi$  is *correct* if for every depth bound  $d \in \mathbb{N}^+$ , every circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$  of depth at most  $d$  and every series of inputs  $\mu_1, \dots, \mu_\ell \in \{0, 1\}$  it holds that

$$\Pr[\text{FHE.Dec}(\text{sk}, \text{FHE.Eval}(\text{pk}, C, \text{FHE.Enc}(\text{pk}, \mu_1), \dots, \text{FHE.Enc}(\text{pk}, \mu_\ell))) \neq C(\mu_1, \dots, \mu_\ell)] \leq \text{negl}(\kappa).$$

- $\Pi$  is *compact* if there exists a polynomial  $s(\cdot)$  such that for every  $\kappa$ , every depth bound  $d$ , every circuit  $C : \{0, 1\}^\ell \rightarrow \{0, 1\}$  of depth at most  $d$ , and every  $\mu_1, \dots, \mu_\ell \in \{0, 1\}$ , the following holds. For  $(\text{pk}, \text{sk}) \leftarrow \text{FHE.Gen}(1^\kappa, 1^d)$ , ciphertexts  $\text{ct}_j \leftarrow \text{FHE.Enc}(\text{pk}, \mu_j)$  for  $j \in [\ell]$ , and  $\text{ct} \leftarrow \text{FHE.Eval}(\text{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell)$ , we have that  $|\text{ct}| \leq s(\kappa, d)$ .
- $\Pi$  is *semantically secure* if for every  $\kappa$  and every depth bound  $d$ , it holds that for every PPT adversary  $\mathcal{A}$  the following experiment  $\text{Expt}_{\mathcal{A}, \Pi}^{\text{FHE}}(1^\kappa, 1^d)$  outputs 1 with negligible probability.

$\text{Expt}_{\mathcal{A}, \Pi}^{\text{FHE}}(1^\kappa, 1^d)$

1. On input the security parameter  $1^\kappa$  and depth bound  $1^d$ , the challenger generates  $(\text{pk}, \text{sk}) \leftarrow \text{FHE.Gen}(1^\kappa, 1^d)$ , chooses a random  $b \xleftarrow{R} \{0, 1\}$ , and computes the ciphertext  $\text{ct} \leftarrow \text{FHE.Enc}(\text{pk}, b)$ . Next, the challenger hands  $(\text{pk}, \text{ct})$  to  $\mathcal{A}$ .
2.  $\mathcal{A}$  outputs  $b'$ . The experiment outputs 1 if  $b = b'$ .

By abuse of notation we consider in the paper encryption strings rather than bits. This should be interpreted as encrypting the string bit by bit.

## B The UC Framework

In this section, we describe the UC framework, for more details see [20].

### B.1 The Real Model

An execution of a protocol  $\pi$  in the real model consists of  $n$  PPT *interactive Turing machines* (ITMs)  $P_1, \dots, P_n$  representing the parties, along with two additional ITMs, an *adversary*  $\mathcal{A}$ , describing the behavior of the corrupted parties and an *environment*  $\mathcal{Z}$ , representing the external network environment in which the protocol operates. The environment gives inputs to the honest parties, receives their outputs, and can communicate with the adversary at any point during the execution. The adversary controls the operations of the corrupted parties and the delivery of messages between the parties.

In more details, each ITM is initialized with the security parameter  $\kappa$  and random coins, where the environment receives an additional auxiliary input. The protocol proceeds by a sequence of *activations*, where the environment is activated first and at each point a single ITM is active. When the environment is activated it can read the output tapes of all honest parties and of the adversary,

and it can activate one of the parties or the adversary by writing on its input tape. Once a party is activated it can perform a local computation, write on its output tape or send messages to other parties by writing on its outgoing communication tapes. After the party completes its operations the control is returned to the environment. Once the adversary is activated it can send messages on behalf of the corrupted parties or send a message to the environment by writing on its output tape. In addition,  $\mathcal{A}$  controls the communication between the parties, and so it can read the contents of the messages on outgoing tapes of honest parties and write messages on their incoming tapes. We assume that only messages that were sent in the past by some party can be delivered, and each message can be delivered at most once.<sup>14</sup>  $\mathcal{A}$  can also corrupt an honest party, gain access to all its tapes and control all its actions. Whenever a party is corrupted the environment is notified. If  $\mathcal{A}$  wrote on the incoming tape of an honest party, this party is activated next, otherwise the environment is activated. The protocol completes once  $\mathcal{Z}$  stops activating other parties and outputs a single bit.

A *semi-honest* adversary always instructs the corrupted parties to follow the protocol. A *malicious* adversary may instruct the corrupted parties to deviate from the protocol arbitrarily. In this work we also consider *semi-malicious* adversaries [5], that instruct the corrupted parties to follow the protocol but can choose arbitrary random coins for them. Formally, the adversary has a special *witness tape*. In each round of the protocol, whenever the adversary produces a new protocol message  $m$  on behalf of some party  $P_k$ , it must also write to its special witness tape some pair  $(x, r)$  of input  $x$  and randomness  $r$  that explains its behavior. More specifically, all of the protocol messages sent by the adversary on behalf of  $P_k$  up to that point, including the new message  $m$ , must exactly match the honest protocol specification for  $P_k$  when executed with input  $x$  and randomness  $r$ . Note that the witnesses given in different rounds need not be consistent. Also, we assume that the attacker is rushing and hence may choose the message  $m$  and the witness  $(x, r)$  in each round adaptively, after seeing the protocol messages of the honest parties in that round (and all prior rounds). Lastly, the adversary may also choose to abort the execution on behalf of  $P_k$  in any step of the interaction.

Let  $\text{REAL}_{\pi, \mathcal{A}, \mathcal{Z}}(\kappa, z, \mathbf{r})$  denote  $\mathcal{Z}$ 's output on input  $z$  and security parameter  $\kappa$ , after interacting with adversary  $\mathcal{A}$  and parties  $P_1, \dots, P_n$  running protocol  $\pi$  with random tapes  $\mathbf{r} = (r_1, \dots, r_n, r_{\mathcal{A}}, r_{\mathcal{Z}})$  as described above. Let  $\text{REAL}_{\pi, \mathcal{A}, \mathcal{Z}}(\kappa, z)$  denote the random variable  $\text{REAL}_{\pi, \mathcal{A}, \mathcal{Z}}(\kappa, z, \mathbf{r})$ , when the vector  $\mathbf{r}$  is uniformly chosen.

## B.2 The Ideal Model

A computation in the ideal model consists of  $n$  *dummy* parties  $P_1, \dots, P_n$ , an *ideal-process adversary* (simulator)  $\mathcal{S}$ , an *environment*  $\mathcal{Z}$ , and an *ideal functionality*  $\mathcal{F}$ . As in the real model, the environment gives inputs to the honest (dummy) parties, receives their outputs, and can communicate with the ideal-process adversary at any point during the execution. The dummy parties act as channels between the environment and the ideal functionality, meaning that they send the inputs received from  $\mathcal{Z}$  to  $\mathcal{F}$  and vice-versa. The ideal functionality  $\mathcal{F}$  defines the desired behaviour of the computation.  $\mathcal{F}$  receives the inputs from the dummy parties, executes the desired computation and sends the output to the parties. The ideal-process adversary does not see the communication between the parties and the ideal functionality, however,  $\mathcal{S}$  can communicate with  $\mathcal{F}$ .

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<sup>14</sup>We assume that all the communication is authenticated yet visible to the adversary; formally, we work in the  $\mathcal{F}_{\text{auth}}$ -hybrid model.

Hiding the communication between the ideal functionality and the parties from the adversary may be too restrictive; it is often desired to provide the adversary the power to determine *when* a party will receive the message. We say that the ideal functionality  $\mathcal{F}$  sends a *delayed output*  $v$  to a party  $P$  if  $\mathcal{F}$  first sends to the adversary a message that it is ready to generate an output to  $P$ . In case the output is public  $\mathcal{F}$  sends  $v$  to the adversary. When the adversary replies to the message,  $\mathcal{F}$  outputs the value  $v$  to  $P$ .<sup>15</sup>

Let  $\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}(\kappa, z, \mathbf{r})$  denote  $\mathcal{Z}$ 's output on input  $z$  and security parameter  $\kappa$ , after interacting with ideal-process adversary  $\mathcal{S}$  and dummy parties  $P_1, \dots, P_n$  that interact with ideal functionality  $\mathcal{F}$  with random tapes  $\mathbf{r} = (r_{\mathcal{S}}, r_{\mathcal{Z}})$  as described above. Let  $\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}(\kappa, z)$  denote the random variable  $\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}(\kappa, z, \mathbf{r})$ , when the vector  $\mathbf{r}$  is uniformly chosen.

**Definition B.1.** *We say that a protocol  $\pi$  UC-realizes an ideal functionality  $\mathcal{F}$  in the presence of adaptive malicious (resp., semi-malicious) adversaries, if for any PPT adaptive malicious (resp., semi-malicious) adversary  $\mathcal{A}$  and any PPT environment  $\mathcal{Z}$ , there exists a PPT ideal-process adversary  $\mathcal{S}$  such that the following two distribution ensembles are computationally indistinguishable*

$$\{\text{REAL}_{\pi,\mathcal{A},\mathcal{Z}}(\kappa, z)\}_{\kappa \in \mathbb{N}, z \in \{0,1\}^*} \stackrel{c}{\equiv} \{\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}(\kappa, z)\}_{\kappa \in \mathbb{N}, z \in \{0,1\}^*}.$$

### B.3 The Hybrid Model

The  $\mathcal{F}$ -hybrid model is a combination of the real and ideal models, it extends the real model with an ideal functionality  $\mathcal{F}$ . The parties communicate with each other in exactly the same way as in the real model described above, however, they can interact with  $\mathcal{F}$  as in the ideal model. An important property of the UC framework is that the ideal functionality  $\mathcal{F}$  in a  $\mathcal{F}$ -hybrid model can be replaced with a protocol that UC-realizes  $\mathcal{F}$ .

Let the global output  $\text{HYBRID}_{\pi,\mathcal{A},\mathcal{Z}}^{\mathcal{F}}(\kappa, z)$  denote  $\mathcal{Z}$ 's output on input  $z$  and security parameter  $\kappa$ , after interacting in a  $\mathcal{F}$ -hybrid model with adversary  $\mathcal{A}$  and parties  $P_1, \dots, P_n$  with uniformly distributed random tapes  $\mathbf{r} = (r_1, \dots, r_n, r_{\mathcal{A}}, r_{\mathcal{Z}})$  running protocol  $\pi$ .

**Theorem B.2** (Canetti [20]). *Let  $\mathcal{F}$  be an ideal functionality and let  $\rho$  be a protocol that UC-realizes  $\mathcal{F}$  in the presence of adaptive malicious (resp., semi-malicious) adversaries, and let  $\pi$  be a protocol that UC-realizes  $\mathcal{G}$  in the  $\mathcal{F}$ -hybrid model in the presence of adaptive malicious (resp., semi-malicious) adversaries. Then for any PPT adaptive malicious (resp., semi-malicious) real-model adversary  $\mathcal{A}$  and any PPT environment  $\mathcal{Z}$ , there exists a PPT adaptive malicious (resp., semi-malicious) adversary  $\mathcal{S}$  in the  $\mathcal{F}$ -hybrid model such that*

$$\{\text{REAL}_{\pi\rho,\mathcal{A},\mathcal{Z}}(\kappa, z)\}_{\kappa \in \mathbb{N}, z \in \{0,1\}^*} \stackrel{c}{\equiv} \{\text{HYBRID}_{\pi,\mathcal{S},\mathcal{Z}}^{\mathcal{F}}(\kappa, z)\}_{\kappa \in \mathbb{N}, z \in \{0,1\}^*}.$$

### B.4 Some Ideal Functionalities

We next describe several ideal functionalities that are used throughout the paper.

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<sup>15</sup>The ideal-process adversary may never release messages from the ideal functionality to the dummy parties and so termination of the computation is not guaranteed. In order to rule out trivial protocols that never produce output, we follow [22] and consider *non-trivial protocols* that have the following property: if the real-model adversary delivers all messages and does not corrupt any parties, then the ideal-process adversary also delivers all messages and does not corrupt any parties. We note that using techniques from [73] guaranteed termination can be enforced.

### B.4.1 Common Reference String

The *common reference string* functionality samples a string from some pre-determined distribution and provides the string to all the parties. The CRS functionality is described in Figure 7.

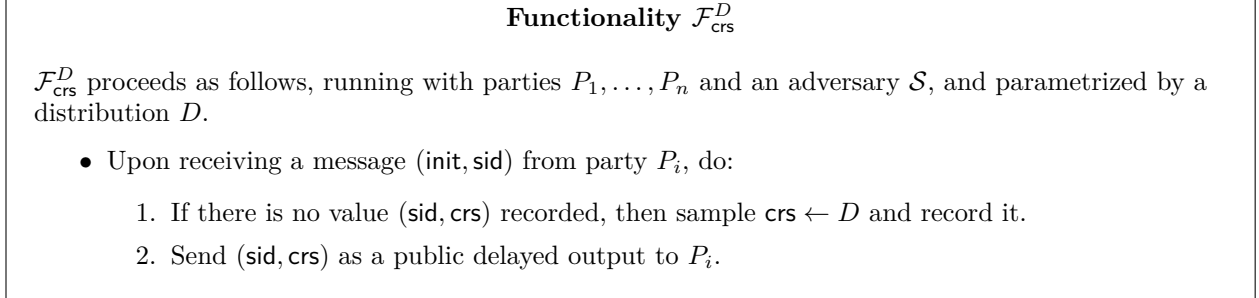


Figure 7: The common reference string functionality

### B.4.2 Secure Message Transmission

The *secure message transmission (SMT)* functionality models a secure and private channel between two parties. The sender can send a message to the receiver such that the adversary learns only a specified leakage of the message, e.g., its length. If the sender is corrupted before the message was delivered to the receiver, the adversary is allowed to change the message. The secure message transmission functionality is described in Figure 8.

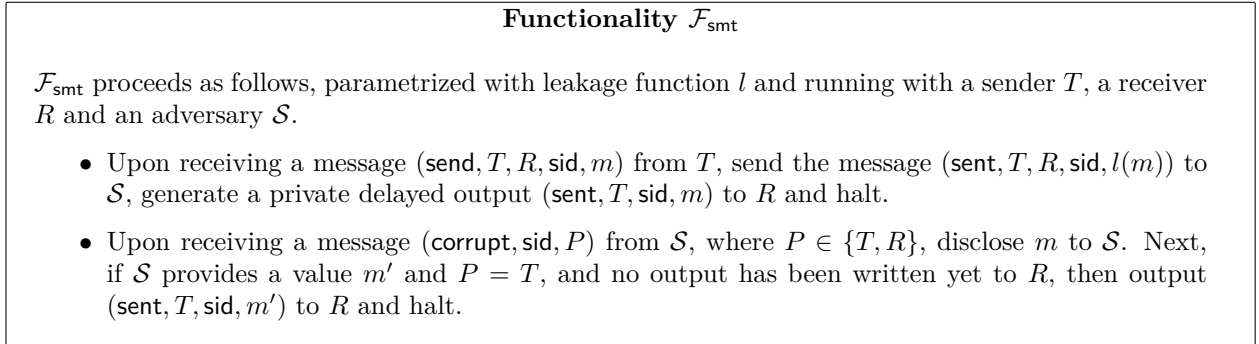


Figure 8: The secure message transmission functionality

### B.4.3 Secure Function Evaluation

*Secure function evaluation (SFE)* is a multiparty primitive where a set of  $n$  parties wish to compute a (possibly randomized) function  $f: (\{0, 1\}^*)^n \times \{0, 1\}^* \rightarrow (\{0, 1\}^*)^n$ , where  $f = (f_1, \dots, f_n)$ . That is, for a vector of inputs  $\mathbf{x} = (x_1, \dots, x_n) \in (\{0, 1\}^*)^n$  and random coins  $r \in_R \{0, 1\}^*$ , the output-vector is  $(f_1(\mathbf{x}; r), \dots, f_n(\mathbf{x}; r))$ . The output for the  $i$ 'th party (with input  $x_i$ ) is defined to be  $f_i(\mathbf{x}; r)$ . The secure function evaluation functionality,  $\mathcal{F}_{\text{sfe}}^f$ , is presented in Figure 9.

**Functionality  $\mathcal{F}_{\text{sfe}}^f$**

$\mathcal{F}_{\text{sfe}}^f$  proceeds as follows, running with parties  $P_1, \dots, P_n$  and an adversary  $\mathcal{S}$ , and parametrized by an  $n$ -party function  $f: (\{0, 1\}^*)^n \times \{0, 1\}^* \rightarrow (\{0, 1\}^*)^n$ . For every  $P_i$  initialize an input value  $x_i = \perp$  and an output value  $y_i = \perp$ .

- Upon receiving a message (input, sid,  $v$ ) from some party  $P_i$ , set  $x_i = v$  and send a message (input, sid,  $P_i, |v|$ ) to  $\mathcal{S}$ .
- Upon receiving a message (output, sid) from some party  $P_i$ , do:
  1. If  $x_j = \perp$  for some honest  $P_j$ , ignore the message.
  2. Otherwise, if  $y_1, \dots, y_n$  have not been set yet, then choose  $r \in_R \{0, 1\}^*$  and compute  $(y_1, \dots, y_n) = f(x_1, \dots, x_n; r)$ .
  3. Generate a delayed output (output, sid,  $y_i$ ) to  $P_i$  and send (output, sid,  $P_i$ ) to  $\mathcal{S}$ .

Figure 9: The secure function evaluation functionality

Note that UC protocols do not provide *guaranteed termination*, since the adversary has full control over the communication. Therefore, it is standard to claim about security in situations where the environment provides sufficiently many activations to the parties, and the adversary delivers all the messages (see [22] for further discussion). We would like to define UC analogues to *security with abort* and *guaranteed output delivery* that are normally defined in the stand-alone model (see [34]). One way is to use the synchronous model of UC [73], where guaranteed termination can be achieved independently of the adversary. Another way is to slightly adjust the  $\mathcal{F}_{\text{sfe}}$  functionality, as discussed below.

Note that the  $\mathcal{F}_{\text{sfe}}$  functionality in some sense guarantee the output delivery, since although the adversary has the power to “hang” the computation, he cannot force an honest party to output an incorrect result. Stated differently, if the protocol terminates, i.e., the adversary provides inputs and deliver the output then the protocol satisfies guaranteed output delivery. For clarity, we denote this functionality by  $\mathcal{F}_{\text{sfe-god}}$ . In the no-honest-majority setting, the adversary has an extra capability, as he can force all parties to output  $\perp$  even when the protocol terminates. To capture this capability, we denote by  $\mathcal{F}_{\text{sfe-abort}}$  the  $\mathcal{F}_{\text{sfe}}$  functionality that allows the adversary to send a special (abort, sid) message at any time. In case some honest party has already received the output value, the functionality ignores this message. Otherwise, the functionality sets the output of all honest parties to  $\perp$ .

#### B.4.4 Broadcast

The *broadcast* functionality enables a sender to reliably deliver a message to all other parties. If the sender is corrupted before the message was delivered to the receivers, the adversary is allowed to change the message. We model this functionality as a special case of the SFE functionality for the function  $f(x, \lambda, \dots, \lambda) = (x, \dots, x)$ .

## C Constructing TEFHE From LWE

In this section, we present an explicit construction of threshold equivocal FHE scheme. we start by describing in Appendix C.1 the GSW scheme, followed by its equivocal variant in Appendix C.2. In Appendix C.3, we present the threshold equivocal FHE. Unlike the rest of the paper, in this section we use  $n$  to denote the dimension of the lattice, and  $N$  to denote the number of parties.

### C.1 GSW Fully Homomorphic Encryption

We now describe the GSW [54] fully homomorphic encryption scheme. We use in the construction the public gadget matrix as defined by Micciancio and Peikert [77], which is a matrix  $\mathbf{G} \in \mathbb{Z}_q^{n \times m}$  with some special structure, such that given a matrix  $\mathbf{V} \in \mathbb{Z}_q^{n \times m}$  everyone can compute a “short” matrix  $\mathbf{G}^{-1}(\mathbf{V}) \in \mathbb{Z}_q^{m \times m}$  satisfying  $\mathbf{G} \cdot \mathbf{G}^{-1}(\mathbf{V}) = \mathbf{V}$ . The GSW scheme is defined as follows:

- $\text{params} \leftarrow \text{GSW.Setup}(1^\kappa, 1^d)$ : Choose a lattice dimension parameter  $n = n(\kappa, d)$ , a  $B_\chi$ -bounded error distribution  $\chi = \chi(\kappa, d)$ , and a modulus  $q$  of size  $q = B_\chi 2^{\omega(d\kappa \log \kappa)}$  such that  $\text{LWE}_{n-1, q, \chi, B_\chi}$  holds. Choose  $m = n \log(q) + \omega(\log \kappa)$ . Finally, choose a random matrix  $\mathbf{B} \in \mathbb{Z}_q^{n-1 \times m}$ . Output  $\text{params} := (q, n, m, \chi, B_\chi, \mathbf{B})$ . We stress that all the other algorithms implicitly get  $\text{params}$  as input even if we usually do not write this explicitly.
- $(\text{pk}, \text{sk}) \leftarrow \text{GSW.Keygen}(\text{params})$ : We separately describe two sub-algorithms to generate secret-key and public-key respectively:
  - $\text{sk} \leftarrow \text{GSW.SKGen}(\text{params})$ : Sample uniformly at random  $\mathbf{s} \leftarrow \mathbb{Z}_q^{n-1}$  and output  $\text{sk} = \mathbf{t} = (-\mathbf{s}, 1) \in \mathbb{Z}_q^n$ .
  - $\text{pk} \leftarrow \text{GSW.PKGen}(\text{params}, \text{sk})$ : Let  $\text{sk} = \mathbf{t}$ , sample  $\mathbf{e} \leftarrow \chi^m$ , set  $\mathbf{b} := \mathbf{t} \cdot \mathbf{B} + \mathbf{e} \in \mathbb{Z}_q^m$  and output  $\text{pk} = \mathbf{A}$  where  $\mathbf{A} \in \mathbb{Z}_q^{n \times m}$  is defined as  $\mathbf{A} = \begin{bmatrix} \mathbf{B} \\ \mathbf{b} \end{bmatrix}$ .
- $\text{ct} \leftarrow \text{GSW.Enc}(\text{pk}, \mu)$ : Choose a short random matrix as the randomness  $\mathbf{R} \leftarrow \{0, 1\}^{m \times m}$ . Then, output the encryption of the message  $\mu \in \{0, 1\}$  as  $\mathbf{C} \in \mathbb{Z}_q^{n \times m}$ , defined as

$$\mathbf{C} = \mathbf{A}\mathbf{R} + \mu\mathbf{G}.$$

- $\mu' = \text{GSW.Dec}(\text{sk}, \text{ct})$ : We decompose the decryption algorithm into two parts:
  - $v = \text{GSW.Dec}^1(\text{sk}, \text{ct})$ : Let  $\text{sk} = \mathbf{t}$  and  $\text{ct} = \mathbf{C}$ . Consider the public vector  $\mathbf{w} = (0, \dots, 0, \lceil q/2 \rceil) \in \mathbb{Z}_q^n$  and output  $v = \mathbf{t} \cdot \mathbf{C} \cdot \mathbf{G}^{-1}(\mathbf{w}^T) \in \mathbb{Z}_q$ .
  - $\mu' = \text{GSW.Dec}^2(v)$ : Output  $\mu' = \left\lfloor \left\lceil \frac{v}{q/2} \right\rceil \right\rfloor$ .
- $\text{GSW.Eval}$ : We define the homomorphic evaluation by defining addition and multiplication. Given ciphertexts  $\mathbf{C}_1, \mathbf{C}_2 \in \mathbb{Z}_q^{n \times m}$ , define:
  - $\text{GSW.Add}(\mathbf{C}_1, \mathbf{C}_2)$ : Output  $\mathbf{C}^{(+)} = \mathbf{C}_1 + \mathbf{C}_2 \in \mathbb{Z}_q^{n \times m}$ .
  - $\text{GSW.Mult}(\mathbf{C}_1, \mathbf{C}_2)$ : Output  $\mathbf{C}^{(\times)} = \mathbf{C}_1 \cdot \mathbf{G}^{-1}(\mathbf{C}_2) \in \mathbb{Z}_q^{n \times m}$ .

**Theorem C.1** ([54]). *The scheme as defined above is a secure FHE scheme under the  $\text{LWE}_{n-1, q, \chi, B_\chi}$  assumption.*



*Proof sketch.* Semantic security of the scheme is proved in two steps. First, the public key  $\mathbf{A}$  is replaced with a uniformly random matrix  $\mathbf{A} \leftarrow \mathbb{Z}_q^{n \times m}$ ; this step is secure under the LWE assumption. Second, the ciphertext  $\mathbf{C} = \mathbf{A}\mathbf{R} + \mu\mathbf{G}$  is replaced with a uniformly random matrix  $\mathbf{C} \leftarrow \mathbb{Z}_q^{n \times m}$ ; this step is secure due to the leftover hash lemma. We refer the reader to [54] for further details.

To prove correctness we start by defining noisy ciphertexts.

**Definition C.2.** A  $\beta$ -noisy ciphertext of some message  $\mu$  under secret key  $\mathbf{sk} = \mathbf{t} \in \mathbb{Z}_q^n$  is a matrix  $\mathbf{C} \in \mathbb{Z}_q^{n \times m}$  such that  $\mathbf{t}\mathbf{C} = \mu\mathbf{t}\mathbf{G} + \mathbf{e}$  for some  $\mathbf{e}$  satisfying  $\|\mathbf{e}\|_\infty \leq \beta$ .

We proceed to analyze the noise behaviour of ciphertexts under encryption, evaluation, and decryption operation.

- **Encryption.** Consider a public key  $\mathbf{A}$  and a secret key  $\mathbf{t}$  generated by  $\text{GSW.Keygen}(1^\kappa, 1^d)$ ; it holds that  $\mathbf{t}\mathbf{A} = \mathbf{e}$  with  $\|\mathbf{e}\|_\infty \leq B_\chi$ . Therefore, a ciphertext  $\mathbf{C} = \mathbf{A}\mathbf{R} + \mu\mathbf{G} \leftarrow \text{GSW.Enc}(\text{pk}, \mu)$  satisfies  $\mathbf{t}\mathbf{C} = \mathbf{e}\mathbf{R} + \mu\mathbf{t}\mathbf{G}$  with  $\|\mathbf{e}\mathbf{R}\|_\infty \leq mB_\chi$ . That is,  $\mathbf{C}$  is a  $mB_\chi$ -noisy ciphertext of  $\mu$  under secret key  $\mathbf{t}$ . We denote  $\beta_{\text{init}} = mB_\chi$ .
- **Addition.** Let  $\mathbf{C}_1$  (resp.  $\mathbf{C}_2$ ) be a  $\beta_1$ -noisy (resp.  $\beta_2$ -noisy) ciphertext of  $\mu_1$  (resp.  $\mu_2$ ) under secret key  $\mathbf{t}$ , i.e.,  $\mathbf{t}\mathbf{C}_i = \mu_i\mathbf{t}\mathbf{G} + \mathbf{e}_i$  with  $\|\mathbf{e}_i\|_\infty \leq \beta_i$  for  $i \in \{1, 2\}$ . Then, for  $\mathbf{C}^{(+)} = \mathbf{C}_1 + \mathbf{C}_2$  it holds that  $\mathbf{t}\mathbf{C}^{(+)} = (\mu_1 + \mu_2)\mathbf{t}\mathbf{G} + \mathbf{e}_1 + \mathbf{e}_2$ , i.e.,  $\mathbf{C}^{(+)}$  is a  $(\beta_1 + \beta_2)$ -noisy ciphertext of  $\mu_1 + \mu_2$  under  $\mathbf{t}$ .
- **Multiplication.** Let  $\mathbf{C}_1$  and  $\mathbf{C}_2$  as above. Then, for  $\mathbf{C}^{(\times)} = \mathbf{C}_1\mathbf{G}^{-1}(\mathbf{C}_2)$  it holds that  $\mathbf{t}\mathbf{C}^{(\times)} = \mu_1\mu_2\mathbf{t}\mathbf{G} + \mathbf{e}$  where  $\mathbf{e} = \mathbf{e}_1\mathbf{G}^{-1}(\mathbf{C}_2) + \mu_1\mathbf{e}_2$ . Therefore,  $\|\mathbf{e}\|_\infty \leq (m\beta_1 + \beta_2)$ , i.e.,  $\mathbf{C}^{(\times)}$  is a  $(m\beta_1 + \beta_2)$ -noisy ciphertext of  $\mu_1\mu_2$  under  $\mathbf{t}$ . The same calculation holds for NAND gates.
- **Decryption.** Let  $\mathbf{C}$  be a  $\beta$ -noisy ciphertext of  $\mu$  under  $\mathbf{t}$ , i.e.,  $\mathbf{t}\mathbf{C} = \mu\mathbf{t}\mathbf{G} + \mathbf{e}$  with  $\|\mathbf{e}\|_\infty \leq \beta$ . Then,  $\mathbf{t}\mathbf{C}\mathbf{G}^{-1}(\mathbf{w}^T) = \mu \lceil q/2 \rceil + e'$  with  $e' = \langle \mathbf{e}, \mathbf{G}^{-1}(\mathbf{w}^T) \rangle$ . It holds that  $|e'| \leq m\beta$ . The decryption will be correct as long as  $|e'| \leq q/4$ , i.e., as long as  $\beta \leq q/(4m)$ . We denote  $\beta_{\text{max}} = q/(4m)$ .

Consider a homomorphic evaluation of Boolean circuit of depth  $d$  consisting of NAND gates. The inputs are encrypted as  $\beta_{\text{init}}$ -noisy ciphertexts, and each level multiplies the noise by a factor of at most  $(m + 1)$ . Therefore, the final output is  $\beta_{\text{final}}$ -noisy ciphertexts, where  $\beta_{\text{final}} = (m + 1)^d \beta_{\text{init}}$ . To ensure correctness of decryption we require that  $\beta_{\text{final}} \leq \beta_{\text{max}}$  meaning  $B_\chi 4m^2 (m + 1)^d < q$  which is satisfied by the choice of parameters for the scheme.  $\square$

## C.2 GSW/GVW Equivocal Fully Homomorphic Encryption

We now describe the construction of the additional algorithms for the equivocal FHE scheme based on the GSW FHE scheme, following [57].

**Lemma C.3** ([1, 53, 2, 77]). *There exist efficient algorithms  $\text{TrapGen}$ ,  $\text{SamPre}$ , and  $\text{Sam}$  such that the following holds. Given integers  $n \geq 1$  and  $q \geq 2$ , there exists some  $m^* = m^*(n, q) = O(n \log q)$  and  $\beta_{\text{sam}} = \beta_{\text{sam}}(n, q) = O(n\sqrt{\log q})$  such that for all  $m \geq m^*$  and all  $k$  (polynomial in  $n$ ) it holds that:*

1.  $\mathbf{U} \leftarrow \text{Sam}(1^m, 1^k, q)$  samples a matrix  $\mathbf{U} \in \mathbb{Z}_q^{m \times k}$  which satisfies  $\|\mathbf{U}\|_\infty \leq \beta_{\text{sam}}$  (with probability 1).

2. We have the statistical indistinguishability requirements:

$$\mathbf{A} \stackrel{s}{\equiv} \mathbf{A}' \quad \text{and} \quad (\mathbf{A}, \text{td}, \mathbf{U}, \mathbf{V}) \stackrel{s}{\equiv} (\mathbf{A}, \text{td}, \mathbf{U}', \mathbf{V}'),$$

where  $\mathbf{A}$  is sampled as  $(\mathbf{A}, \text{td}) \leftarrow \text{TrapGen}(1^n, 1^m, q)$  and  $\mathbf{A}' \leftarrow \mathbb{Z}_q^{n \times m}$  is uniformly random. Likewise,  $\mathbf{U} \leftarrow \text{Sam}(1^m, 1^k, q)$ ,  $\mathbf{V} = \mathbf{A}\mathbf{U}$ ,  $\mathbf{V}' \leftarrow \mathbb{Z}_q^{n \times k}$  is uniformly random, and  $\mathbf{U}' \leftarrow \text{SamPre}(\mathbf{A}, \mathbf{V}', \text{td})$ . The statistical distance is negligible in  $n$ . Moreover, we guarantee that any  $\mathbf{U}' \in \text{SamPre}(\mathbf{A}, \mathbf{V}', \text{td})$  always satisfies  $\mathbf{A}\mathbf{U}' = \mathbf{V}'$  and  $\|\mathbf{U}'\|_\infty \leq \beta_{\text{sam}}$ .

We proceed to define the additional algorithms for the equivocal FHE:

- $(\text{pk}, \text{td}) \leftarrow \text{GSW.GenEquiv}(\text{params})$ : Select  $(\mathbf{A}, \text{td}) \leftarrow \text{TrapGen}(1^n, 1^m, q)$  and set  $\text{pk} = \mathbf{A}$ .
- $\text{GSW.Equiv}(\text{td}, \text{ct}, \mu)$ : Let  $\text{ct} = \mathbf{C}$  and output  $\text{SamPre}(\mathbf{A}, \mathbf{C} - \mu\mathbf{G}, \text{td})$ .

### C.3 Threshold Equivocal Fully Homomorphic Encryption

We proceed by adjusting the EFHE scheme to support threshold key-generation and decryption.

- $\text{TEFHE.Gen}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{sk}_1, \dots, \text{sk}_N)$ : Compute  $\text{params} \leftarrow \text{GSW.Setup}(1^\kappa, 1^d)$ , set  $(\text{pk}, \text{sk}) \leftarrow \text{GSW.Keygen}(\text{params})$ , and let  $(\text{sk}_1, \dots, \text{sk}_N)$  be a linear secret sharing of  $\text{sk}$ , i.e.,  $\sum_{i \in [N]} \text{sk}_i = \text{sk} \pmod q$ .
- $\text{TEFHE.Enc}(\text{pk}, \mu) \rightarrow \text{ct}$ : Output  $\text{GSW.Enc}(\text{pk}, \mu)$ .
- $\text{TEFHE.Eval}(\text{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell) \rightarrow \text{ct}$ : Output  $\text{GSW.Eval}(\text{pk}, C, \text{ct}_1, \dots, \text{ct}_\ell)$ .
- $\text{TEFHE.PartDec}(i, \text{sk}_i, \text{ct}) \rightarrow \mathbf{p}_i$ : Output partial decryption  $\mathbf{p}_i = \text{GSW.Dec}^1(\text{sk}_i, \text{ct}) + e_i$ , where  $e_i \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  is some random ‘‘smudging noise,’’ where  $B_{\text{smdg}} = 2^{d\kappa} \log^\kappa B_\chi$ .
- $\text{TEFHE.FinDec}(\text{pk}, \mathbf{p}_1, \dots, \mathbf{p}_N) \rightarrow \tilde{\mu}$ : Output  $\text{GSW.Dec}^2(\sum_{i=1}^N \mathbf{p}_i)$ .
- $\text{TEFHE.GenEquiv}(1^\kappa, 1^d) \rightarrow (\text{pk}, \text{td})$ : Output  $\text{GSW.GenEquiv}(\text{params})$ .
- $\text{TEFHE.Equiv}(\text{td}, \text{ct}, \mu) \rightarrow r$ : Output  $\text{GSW.Equiv}(\text{td}, \text{ct}, \mu)$ .

We will use the following ‘‘smudging lemma’’ [5] to prove correctness of the threshold scheme.

**Lemma C.4** ([5]). *Let  $B_1 = B_1(\kappa)$  and  $B_2 = B_2(\kappa)$  be positive integers, and let  $e_1 \in [-B_1, B_1]$  be a fixed integer. Let  $e_2 \leftarrow [-B_2, B_2]$  be chosen uniformly at random. Then, the distribution of  $e_2$  is statistically indistinguishable from that of  $e_2 + e_1$  as long as  $B_1/B_2 = \text{negl}(\kappa)$ .*

To prove correctness, we need to show that given a ciphertext matrix  $\mathbf{C}$ , it holds that by computing  $\mathbf{p}_i = \text{GSW.Dec}^1(\text{sk}_i, \text{ct}) + e_i$  for  $e_i \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  followed by  $\text{GSW.Dec}^2(\sum_{i=1}^N \mathbf{p}_i)$ , we get the same result as computing  $v = \text{GSW.Dec}^1(\text{sk}, \mathbf{C})$  followed by  $\text{GSW.Dec}^2(v)$ . Note that by defining  $\mathbf{C}_1 = \mathbf{C}\mathbf{G}^{-1}(\mathbf{w}^T) \in \mathbb{Z}_q^n$ , it holds by linearity of inner product that:

$$\begin{aligned} \sum_{i=1}^N \mathbf{p}_i &= \sum_{i=1}^N (\text{GSW.Dec}^1(\mathbf{t}_i, \mathbf{C}) + e_i) = \sum_{i=1}^N (\mathbf{t}_i \mathbf{C} \mathbf{G}^{-1}(\mathbf{w}^T) + e_i) = \sum_{i=1}^N (\mathbf{t}_i \mathbf{C}_1 + e_i) \\ &= \sum_{i=1}^N (\langle \mathbf{t}_i, \mathbf{C}_1 \rangle + e_i) = \langle \sum_{i=1}^N \mathbf{t}_i, \mathbf{C}_1 \rangle + \sum_{i=1}^N e_i = \langle \mathbf{t}, \mathbf{C}_1 \rangle + \sum_{i=1}^N e_i \\ &= \text{GSW.Dec}^1(\mathbf{t}, \mathbf{C}) + \sum_{i=1}^N e_i. \end{aligned}$$

To prove simulatability, we construct a simulator  $\text{Sim}_{\text{TEFHE}}$  that receives as input the ciphertext  $\text{ct}$  describing a matrix  $\mathbf{C} \in \mathbb{Z}_q^{n \times m}$ , a plaintext  $\mu \in \{0, 1\}$ , and the secret key of all parties but the  $i$ 'th  $\{\text{sk}_j\}_{j \neq i}$ , where each  $\text{sk}_j$  is of the form  $\mathbf{t}_j = (\mathbf{s}_j, 1) \in \mathbb{Z}_q^n$ . The simulator starts by setting  $\mathbf{C}_1 = \mathbf{C}\mathbf{G}^{-1}(\mathbf{w}^T)$  for  $\mathbf{w} = (0, \dots, 0, \lceil q/2 \rceil) \in \mathbb{Z}_q^n$ , and computing  $\gamma_j = \langle \mathbf{t}_j, \mathbf{C}_1 \rangle$  for every  $j \neq i$ . Next, sample smudging noise  $e_i^{\text{sm}} \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  and set

$$\mathbf{p}'_i = \mu \lceil q/2 \rceil + e_i^{\text{sm}} - \sum_{j \neq i} \gamma_j.$$

To prove the statistical indistinguishability, note that by the same calculation as used to argue correctness, we know that  $\sum_{j \in [N]} \gamma_j = \mu \lceil q/2 \rceil + e'$  with  $|e'| \leq \beta_{\text{final}} m N = 2^{O(d \log \kappa)} B_\chi$ . Denote that real partial decryption is  $\mathbf{p}_i = \gamma_i + e_i^{\text{sm}}$ , then it holds that

$$\mathbf{p}_i = \mu \lceil q/2 \rceil + e_i^{\text{sm}} + e' - \sum_{j \neq i} \gamma_j.$$

The difference between the real value  $\mathbf{p}_i$  and the simulated value  $\mathbf{p}'_i$  is the noise  $e'$  of norm  $|e'| = 2^{O(d \log \kappa)} B_\chi$ . By Lemma C.4, the distributions  $e_i^{\text{sm}}$  and  $e_i^{\text{sm}} + e'$  are statistically close since  $e_i^{\text{sm}} \leftarrow [-B_{\text{smdg}}, B_{\text{smdg}}]$  for  $B_{\text{smdg}} = 2^{O(d \log \kappa)} B_\chi$ , hence  $B_{\text{smdg}}/|e'| \geq 2^\kappa$ .