Continuous NMC Secure Against Permutations and Overwrites, with Applications to CCA Secure Commitments

Ivan Damgård[1], Tomasz Kazana[2], Maciej Obremski[1], Varun Raj[3], Luisa Siniscalchi[4]

[1] Aarhus University
[2] University of Warsaw
[3] National University Singapore
[4] University of Salerno

Abstract. Non-Malleable Codes (NMC) were introduced by Dziembowski, Pietrzak and Wichs in ICS 2010 as a relaxation of error correcting codes and error detecting codes. Faust, Mukherjee, Nielsen, and Venturi in TCC 2014 introduced an even stronger notion of non-malleable codes called continuous non-malleable codes where security is achieved against continuous tampering of a single codeword *without* re-encoding.
We construct information theoretically secure CNMC resilient to bit permutations and overwrites, this is the first Continuous NMC constructed outside of the split-state model.
In this work we also study relations between the CNMC and parallel CCA commitments. We show that the CNMC can be used to bootstrap a self-destruct parallel CCA bit commitment to a self-destruct parallel CCA string commitment, where self-destruct parallel CCA is a weak form of parallel CCA security. Then we can get rid of the self-destruct limitation obtaining a parallel CCA commitment, requiring only one-way functions.

1 Introduction

In this paper, we study the interesting relationship between the notions of non-malleable codes and non-malleable commitments, and advance state of art for both of them. Before giving our results, we introduce the notions.

1.1 Introduction to Non-Malleable Codes

Non-Malleable Codes (NMC) were introduced by Dziembowski, Pietrzak and Wichs [DPW10] as a relaxation of error correcting codes and error detecting codes. An NMC takes a message m and encodes it as a possibly longer and randomized codeword $c \leftarrow \mathsf{Enc}(m)$. The adversary chooses and submits a tampering function Tamper, that is applied to the code word to yield $c' = \mathsf{Tamper}(c)$. Applying the decoding algorithm yields a message $m' = \mathsf{Dec}(c')$. The security

guarantee for an NMC now is that the decoded message m' is either identical to the original message m or, in case of a decoding error, a message *unrelated* to m. Correspondingly, the adversary is given either m' or a symbol "same" indicating that decoding was successful. Technically, we require that if $m' \neq m$, then m' can be simulated using just the tampering function Tamper, but without knowing anything about the tampered codeword c'.

It is generally impossible to give any meaningful guarantees if the tampering function is unrestricted (the tamper function could decode, and then encode a modified message). Therefore, the tampering function Tamper is always assumed to come from some class $\mathbb{T}$ of functions. An immediate example application of NM codes is for tamper resilient cryptography: if a secret key is stored in a hardware device, the adversary could try to tamper with the device and observe its behavior after the modification. But if the key is encoded with an NM code, the security guarantees immediately imply that either the tampering had no effect or the effect can be simulated without the device.

Continuous Non-Malleable Codes (CNMC) As mentioned in [JW15], non-malleable codes can provide protection against these kind of attacks if the device is allowed to freshly re-encode its state after each invocation to make sure that the tampering is applied to a fresh codeword at each step. After each execution the entire content of the memory is erased. While such perfect erasures may be feasible in some settings, they are rather problematic in the presence of tampering. Due to this reason, Faust et al. [FMNV14] introduced an even stronger notion of non-malleable codes called continuous non-malleable codes where security is achieved against continuous tampering of a single codeword *without* re-encoding. In this model the adversary can iteratively submit tampering functions Tamper_i and learn $m_i = \mathsf{Dec}(\mathsf{Tamper}_i(c))$. We call this the *continuous tampering model*. This stronger security notion is needed in many setting, for instance when using NMCs to make tamper resilient computations on von Neumann architectures [FMNV15].

Some additional restrictions are, however, necessary in the continuous tampering model. If the adversary was given an unlimited budget of tampering queries, then, given that the class of tampering functions is sufficiently expressive (e.g. it allows to overwrite single bits of the codeword), the adversary can efficiently learn the entire message just by observing whether tampering queries leave the codeword unmodified or lead to decoding errors, see e.g. [GLM+04].

To overcome this general issue, [FMNV14] assume a *self-destruct* mechanism which is triggered by decoding errors. In particular, once the decoder outputs a special symbol $\perp$ the device *self-destructs* and the adversary loses access to his tampering oracle. This model still allows an adversary many tamper attempts, as long as his attack remains covert. Jafargholi and Wichs [JW15] considered four variants of continuous non-malleable codes depending on

- Whether tampering is *persistent* in the sense that the tampering is always applied to the current version of the tampered codeword, and all previous versions of the codeword are lost. The alternative definition considers non-

persistent tampering where the device resets after each tampering, and the tampering always occurs on the original codeword.
– Whether tampering to an invalid codeword (i.e., when the decoder outputs $\perp$) causes a *"self-destruct"* and the experiment stops and the attacker cannot gain any additional information, or alternatively whether the attacker can always continue to tamper and gain information.

A long line of research has tried to optimize the performance of NM codes with respect to the number of allowed tampering queries and the class of allowed tampering functions (see the related work section for details). In this paper we will be concerned with the case of CNMCs where there is no a priori bound on the number of queries. This model must include a *self-destruct* mechanism.Further we will be concerned with information theoretic NM codes where security holds for an unbounded adversary, and we will look at the single state model, where the tampering function is allowed to access the entire codeword. This is in contrast to the split-state model where the tamper function must consider disjoint parts of the codeword separately.

1.2 NMC- Our result

We give a construction of a *self-destruct, non-persistent* continuous NMC (see Corollary 1 of Theorem 1) unconditionally secure against *bit permutations* composed with *bit overwrites*. [AGM+15] gives a *one time* Non-Malleable Code resilient against *bit permutations* composed with *bit-wise tampering*. In [CMTV15] they construct a CNMC secure against *bitwise tampering* (but permutations are not allowed).

Unconditionally secure Continuous Non-Malleable Codes are notoriously hard to construct. Very little progress was made since CNMC were proposed in 2015:

– [CMTV15] authors construct a CNMC secure against *bitwise tampering* which is the simplest variant of split-state model.
– [CGL16] authors achieve a so-called many-many non-malleable code in the 2-split state model. Their construction achieves non-malleability as long as the number of rounds of tampering is at most n^γ for some constant $\gamma < 1$, where n is the length of the codeword.
– [AKO17] authors give the *persistent* continuous NMC construction for $2-$split state.
– [ADN+16] gives Continuous NMC against $8-$split state tampering (optimal number of states would be 3).

This makes our result the first known unconditionally secure construction of CNMC outside of split-state model.

1.3 NMC- Related work

In [DPW10] the authors construct an efficient code which is non-malleable with respect to bit-wise tampering, i.e., tampering functions that modify each bit of

the codeword arbitrarily but independently of the value of the other bits of the codeword. Later works [DKO13, ADL14, CZ14, CG14, Li17] provided stronger results by considering a model where the codeword is split into s parts called *states*, which can each be tampered arbitrarily but independently of the other states. Other works considered tampering via permutations and perturbations [AGM+14], which are not captured in the split-state model. In [BDSKM16] authors show how to construct efficient, unconditionally secure non-malleable codes for bounded output locality (i.e. when every bit of tampering output can depend on at most some n^δ bits of input for $\delta < 1$).

The definition in [DPW10] allows the adversary to be computationally unbounded. We call this an *information theoretic* NMC. Later works considered a notion of *computational* NMC where the adversary and tampering functions are restricted to efficient computations, see for instance [CKM11, LL12, AAnHKM+16, BDSKM17].

The definition in [DPW10] allows the adversary to tamper the codeword *only once*. We call this *one-shot* tampering. Faust *et al.* [FMNV14] consider a stronger model where the adversary can iteratively submit tampering functions Tamper_i and learn $m_i = \mathsf{Dec}(\mathsf{Tamper}_i(c))$. We call this the *continuous tampering model*. This stronger security notion is needed in many setting, for instance when using NMCs to make tamper resilient computations on von Neumann architectures [FMNV15]. Some additional restrictions are, however, necessary in the continuous tampering model. If the adversary was given an unlimited budget of tampering queries, then, given that the class of tampering functions is sufficiently expressive (e.g. it allows to overwrite single bits of the codeword), the adversary can efficiently learn the entire message just by observing whether tampering queries leave the codeword unmodified or lead to decoding errors, see e.g. [GLM+04].

To overcome this general issue, [FMNV14] assume a *self-destruct* mechanism which is triggered by decoding errors. In particular, once the decoder outputs a special symbol $\perp$ the device *self-destructs* and the adversary loses access to his tampering oracle. This model still allows an adversary many tamper attempts, as long as his attack remains covert. Jafargholi and Wichs [JW15] provide a general study of when CNMCs can be built assuming a self-destruct mechanism.

Faust et al. [FMNV14] constructed a CNMC in the 2-state model which is secure against computationally bounded adversaries. It was shown in the same work that it is *impossible* to construct an information theoretic CNMC for the 2-state model.

Information-theoretic results for CNMC. In [CMTV15] authors construct a CNMC secure against *bitwise tampering* which is the simplest variant of split-state model. In [AKO17] authors give the first information theoretic *persistent* continuous NMC construction for 2−split state. Finally in [ADN+16] authors give the first information theoretic construction of CNMC in 8− split state. Before [ADN+16] the only known result that achieves some sort of non-malleable codes secure against *non-persistent* continuous tampering was the result by Chattopadhyay, Goyal, and Li [CGL16]. They achieve this by constructing a so-called many-

many non-malleable code in the 2-split state model. Their construction achieves non-malleability as long as the number of rounds of tampering is at most n^γ for some constant $\gamma < 1$, where n is the length of the codeword.

1.4 Application to Commitment Schemes

Commitment schemes. The notion of commitment is perhaps the most fundamental concept in cryptographic protocol design. The idea is that a sender binds herself to a choice of a message m by exchanging some information with a receiver. The commitment should be *hiding*, i.e., the verifier does not learn the committed message. Later, the sender can choose to open the commitment, i.e., release more information allowing the receiver to determine m. The commitment should be *binding*, i.e., the sender cannot make the receiver output a message different from the one she had in mind at commit time.

The strongest possible security notion for commitment schemes is UC security, which intuitively asks that using the scheme is equivalent to giving m to a trusted party who will only release it on request from the committer. This is much stronger than simply asking for hiding and binding, e.g., we get security under general composition. But unfortunately, we know that UC security cannot be achieved without set-up assumptions. So a long line of research has been aimed at achieving weaker but meaningful security guarantees without set-up.

An important example of this is the notion of non-malleable (NM) commitments [DDN91]. Here we consider an adversarial Man-in-the-middle (MiM), who on side receives a commitment from an honest sender to message m (the "left session") and on the other side sends a commitment to an honest receiver (the "right session"), containing m'. The MiM wins if he succeeds in forming a new commitment on the right such that m' has some non-trivial relation to m. The NM property does not follow from hiding and binding and is very important, for instance in making auctions where committed bid is fair, or towards implementing secure coin-flipping. Technically the NM property is captured by requiring a simulator that will simulate the left session without knowing m and still the MiM wins with essentially the same probability.

The strongest form of NM commitment security is concurrent NM commitments. Here, the MiM is allowed to start any number of left sessions and right sessions and can schedule them as he likes. One can also consider restricted versions of this, for instance a 1-1 NM commitment is secure if only 1 left and 1 right session is allowed. A restriction that we want to consider is self-destruct (SD) concurrent non-malleable commitment. In this version, once the MiM makes a invalid commitment in a right session, all commitment computed after that session are considered invalid and cannot be used to win the game. This notion is close in spirit to the one of the weak non-malleable commitments, which has been applied in multiple works.

An even stronger notion of commitment security is CCA security([CLP10]): we consider again a MiM, but he is now given an oracle that he can query on input a commitment from (one of) the right session(s), as long as it is not a copy of something from a left session. The requirement is that hiding holds for the left

session(s), even in presence of the oracle. Intuitively, a CCA secure commitment is also NM secure, all other things being equal: if the MiM could break NM security and come up with a new commitment on the right side that is related to one from the left, he could submit it to the oracle in the CCA game and use the reply to break hiding on the left side. One restriction on CCA commitments that has been considered is parallel CCA security, where the MiM can ask only one query that may, however, contain an unbounded number of commitments. Another restriction is that of self-destruct (SD)-CCA, where the oracle stops working if the MiM submits an invalid commitment.

Parallel CCA commitments from CNMC. In this second part we investigate possible applications our CNMC. In particular, we will show a bridge between (unconditionally secure) CNMC and (computational) cryptographic primitives secure in the concurrent setting.

For the stand-alone setting the result of [AGM+15] shows how to use a bit parallel CCA commitment[5] to construct a *1-1* string non-malleable commitment relying on stand-alone NM code. In particular, constructing string commitment from the corresponding 1-bit primitive, they first encode the input message with an NM code and then apply a 1-bit commitment scheme.

Following the same approach of [AGM+15] but using a CNMC (resilient to the same class of tampering functions of [AGM+15]) we are asking which flavor of non-malleability w.r.t. commitment we can achieve. In particular, is it enough to plug-in our CNMC in the construction of [AGM+15] to obtain a concurrent NM string commitment? The answer is only partial yes, due to the self-destruct limitation of CNMC. Indeed, a MiM adversary of NM commitments can compute multiple invalid commitments. Then, we show how to bypass this limitations requiring only OWFs.

In more details, we obtain a compiler that takes a CCA bit commitment and constructs an SD concurrent NM commitment. Due to the adaptiveness of our NM code we actually achieve a stronger security notion, namely a string SD-CCA commitment scheme. Furthermore we can relax the requirements on the CCA bit commitment: it just needs to be SD-CCA-secure instead of CCA-secure.

Summarizing, we show a compiler that on input a (non-tag based) SD-CCA bit commitment scheme and a continuous non-malleable code resilient against permutations and bit overwrites, outputs a (non-tag based) SD-CCA string commitment scheme. Our construction, like the one of [AGM+15], preserves the

[5] Note that a particular accent is placed on the fact that the compiler requires as input a possible (non-tag based) n-parallel bounding CCA bit commitment because. The reduction is non-trivial only because they are working in the standard non-tag based setting. Otherwise, in case of tags, one can simply sign the entire transcript using the tags and obtain a non-malleable string commitment. In case of bit commitments, tag-based non-malleability is a stronger requirement than the standard (non-tag-based) non-malleability. Pass and Rosen [PR05] argue that for string commitments, the two notions are equivalent since one can simply commit the tag as part of the string, if there are no tags. Since we only have bit commitments, this does not work.

round complexity of the bit commitment scheme and does not require any additional assumption. Finally , we show that a SD-parallel CCA string commitment scheme can be upgraded to a parallel string commitment scheme without self-destruct, assuming only one-way functions. The construction is non-trivial (it requires very recent developed tecniques) and adds only two rounds of interaction.

Together with our compiler described above, this implies the first construction that exploit the CNMC property to obtain a parallel CCA commitment. Furthermore, parallel CCA commitment founds multiple applications like [Kiy15, BDH+17]. Observe that parallel CCA commitment is not implied by parallel NM commitment (see [BFMR18]).

Previous work on NMCs and NM commitments. The literature presents works that exploit the properties of the non-malleable code to construct non-malleable commitments. Goyal et al. [GPR16] use non-malleable codes in the split-state model to realize a 3-round one-one non-malleable commitment relying on one way permutations secure against a quasi-polynomial time adversary. Chandran et al. [CGM+16] show that block non-malleable codes with t blocks imply non-malleable commitments of $t-1$ rounds. As we discuss above, Agrawal et al. [AGM+15] showed that is possible to construct a one-one non-malleable commitment relying on a non-malleable code and a bounded parallel CCA bit commitment. However, no one before uses non-malleable codes to construct a parallel CCA commitment scheme.

The aim of this second part is to build bridges between different notions of non-malleability, and to not construct a new NM commitment or a CCA commitment that are already available in literature. Indeed, there is a long line of research that tries to reduce the round complexity of NM commitment (e.g. [DDN91, Bar02, PR05, Wee10, PW10, LP11, Goy11, GLOV12, GRRV14, GPR16, COSV16, COSV17b, KS17, Khu17, LPS17]). Several constructions of CCA commitment are also available in literature (e.g. [CLP10, LP12, Kiy14, GLP+15]).

Improve the SD-RCCA PKE scheme of [CMTV15]. As a second application of our CNMC we demonstrate that by plugging in our CNMC in the construction of [CMTV15] we can improve the efficiency of their SD-RCCA PKE scheme. In particular, if we plug our non-malleable code in the construction of [CMTV15] we will obtain a modified SD-RCCA PKE that is improved in the following aspects: 1) the modified SD-RCCA PKE uses a single public/secret key pair, instead of n (where n is the size of the codeword); 2) we obtain a construction with no a priori bound on the length of the string to be encrypted.

1.5 Technical overview of our CNMC secure against permutations-and-overwrites

Construction of Continuous Non-Malleable Code. Our code consists of an amalgamation of two different layers of encoding schemes.

The top layer is a Reed-Solomon code used here as a sharing scheme. We take a message m, append a random suffix and then encode it using Reed-Solomon to receive a codeword consisting of N blocks that may be seen as shares of $\lfloor \frac{N}{3} \rfloor$-out-of-$N$ secret sharing scheme. The intuition behind this scheme is that the adversary needs to learn at least $\frac{N}{3}$ shares to learn anything about the initial message.

The bottom layer is using a Two-Split State Super Strong Non-Malleable Code (instantiated either by [AKO17] or [Li17]). Each share s_i from the above secret sharing scheme is converted into $(s_i||i)$ and then encoded using the two-split state code to get two shares (L_i, R_i) (We also expect the bit-parity of L_i to be 0 and the bit-parity of R_i to be 1). The final code is $(L_1, R_1, ..., R_N, L_N)$.

To prove that the just described code is actually continuous non-malleable code, we first redefine the experiment in the definition of continuous codes. The new definition is obviously stronger, so it is sufficient to work with it. In the new definition, whenever an adversary tampers with a blocks (L_i, R_i) with non-constant functions and succeeds in creating valid (from the point of view of Super Strong NMC decoder) output blocks (L'_i, R'_i) (In particular, the parities of all (L'_i, R'_i) must be correct), we will reveal blocks (L_i, R_i) to the adversary.

As observed earlier, the adversary's necesary task is to learn at least $\lfloor \frac{N}{3} \rfloor$ blocks of the underlying s_i shares.

Since the adversary can only tamper bitwise and permute bits we can prove that if the adversary doesn't *know* $\frac{N}{3}$ blocks and he tries to modify the code he will either get detected with probability exponentially close to 1, or he can attempt to learn some small amount information about the codeword (i.e. tamper with few blocks L_i, R_i with non-constant function). However, using the bottom layer, we show that every attempt to learn even the smallest information about the codeword (i.e. by overwriting all but only few bits) yields some probability of detection which amplifies with amount of information adversary is trying to learn. We will therefore show that adversary can not (i.e. the probability is negligable) breach $\lfloor \frac{N}{3} \rfloor$ blocks threshold.

The argument consists of two main technical lemmata:

- [lemma 8] If the adversary applies any non- constant functions f, g to single block L_i, R_i then, due to combination of super strong nmc properties and parity requirements we have placed on L_i, R_i, adversary risks close to $\frac{1}{2}$ detection probability.
- [lemma 10] If the adversary decides to mix bits between different blocks (L_i, R_i) he has to risk violation of parity requirements on these blocks. This lemma is inspired by similar lemma for unary schemes from [AGM+14].

Using these ideas we can claim that if adversary tampers with k blocks using non-constant functions he also gets detected with a probability $1 - p^{-k}$.The proof of this fact is more involved because we have to deal with minute cases. For example if we prove that mixing of bits will make the parity unpredictable for each block it still may happen that the events of error are correlated so not obviously amplify the error rate.

Example 1. Assume adversary tampers only with L_1 and L_2, if he permutes bits in a way that output L'_1 contains first halfs of vectors L_1, L_2, L'_2 contains second halfs of L_1, L_2. Then parity of L'_1 is correct if and only if parity of L'_2 is correct. We handle this by picking only largest possible subset of independent parity checks (see lemma 9) in this case we would focus only on parity of L'_1 and discard any other checks generated by L_1, L_2, R_1, R_2.

Example 2. Consider a tampering function which takes one bit from some blocks (L_i, R_i) and permutes them to the last block (L'_N, R'_N) while fixing all other (L_i, R_i) to some constants. If (L'_N, R'_N) has a correct parity and valid Super-Strong NMC decoding then we will reveal, to adversary, all blocks that 'donated' bits to (L'_N, R'_N). Notice however that this will not reveal more bits then $|L_N| + |R_N|$ blocks.

Above examples illustrate how we bound number of blocks adversary can learn for each independent validity check he has to create.

Technical overview of our self-destruct CCA commitment and parallel CCA commitment.

The self-destruct CCA commitment scheme. We want to show that given a self-destruct CCA bit commitment scheme (non-tag based), committing to each bit of the codeword individually, results in a self-destruct CCA string commitment scheme. The security proof is based on the following high-level idea: if the adversary of the self-destruct CCA string commitment is mauling, then, the attack on the commitment level can be "translated" into an attack on the non-malleable code. In other word, we can show an adversary $\mathcal{A}_{\mathsf{NMCode}}$ that breaks the security of the non-malleable code using the adversary $\mathcal{A}$ on the commitment level that distinguish a commitment of message m_0 from a commitment of message m_1. $\mathcal{A}_{\mathsf{NMCode}}$ will act as the sender in the left session with $\mathcal{A}$. Instead in the k-th right session (for $k = 1, \ldots, \mathtt{poly}(\lambda)$) $\mathcal{A}_{\mathsf{NMCode}}$ will act as a receiver of the string commitment. Then he needs to emulate the oracle $\mathcal{O}$ of the string commitment computing the following steps: 1) define a tamper function f_k based on value v committed in the right session (note that he can obtain v querying the oracle of the bit commitment $\mathcal{O}^{\mathtt{bit}}$[6]) 2) send back to $\mathcal{A}$ the decoding of $f_k(\mathsf{enc}_{m_b})$, where enc_{m_b} is an encoding of m_b (received from the challenger of the non-malleable code game). At the end, $\mathcal{A}_{\mathsf{NMCode}}$ will output what $\mathcal{A}$ outputs. However we notice that the adversary that we described is not yet an adversary against the non-malleable code since the tamper functions can be dependent on what is committed on the left. We can demonstrate that the hiding of the self-destruct CCA bit commitment ensures that the distribution of the tamper functions is computational independent from the message committed by the sender. Therefore the final adversary against the non-malleable code will simply commits to a random message on the left session. Finally, we crucially need that the non-malleable code is information theoretic secure since we have no guarantee that $\mathcal{O}^{\mathtt{bit}}$ works in polynomial time.

[6] The definition of the tamper function is more complicated, see Section 5 for the details.

Upgrade SD-PCCA commitment scheme to PCCA commitment scheme. At a very high level our PCCA string commitment scheme works as follows. The sender interacts with the receiver in order to compute a commitment τ of m using a self-destruct PCCA string commitment. Furthermore, the receiver engages with the sender a protocol to allow the extraction of a trapdoor. We use the "trapdoor protocol" described in [COSV17b] where the trapdoor is represented by the knowledge of two signatures under a verification key sent by receiver in the 4th last round. In order to allow the extraction of the trapdoor, the receiver sends a signature of a randomly chosen message in the 3rd last round by the sender. Then, the sender executes a special witness-indistinguishable proof of knowledge (WIPoK) with the receiver in order to prove that he computed a valid commitment of m or that he knows a trapdoor.

Observe that if we use a 3-round WIPoK it is not clear how the proof of security will proceed. In particular, in the security proof there are some hybrids were we simulate the oracle of the parallel CCA commitment in polynomial time extracting the committed messages from the WIPoKs. Let us consider the hybrid were we switch the witness in one of the WIPoK. In the reduction to the WI we have to emulate the oracle of the parallel CCA commitments, since the reduction has to work in polynomial time. As we said, our hope to emulate the oracle is to extract the committed messages from the WIPoKs, however the extraction procedure rewinds also the challenger of the WI.

To overcome this problem we adopt the approach proposed in [COSV17b] relying on non-interactive primitives instead of 3-rounds WIPoK.

Therefore, similarly to [COSV17b], we construct this WIPoK relying on: instance-dependent trapdoor commitments (IDTC) and special honest-verifier zero knowledge (SHVZK).

In more details, let $(\mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ be the transcript of a 4-round special HVZK delayed-input[7] proof of knowledge (PoK). The transcript $(\mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ is used to prove knowledge of two signatures of two different message w.r.t. a verification key sent by the reaciver. The transcript $(\mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ is used to prove the knowledge of the trapdoor.

At the 4th last round the sender sends an equivocal $\mathtt{com}$ obtained running IDTC. At last round the sender will equivocate $\mathtt{com}$ in order to send as opening $(\mathtt{dec}, \mathsf{ls}^2_{\mathsf{trap}})$. In the last round also $\mathsf{ls}^4_{\mathsf{trap}}$ is sent. The instance used for the IDTC is τ, this means that the commitment $\mathtt{com}$ (computed using IDTC) can be opened to any value because τ is a well-formed commitment.

In the opening phase the sender sends the opening of the self-destruct PCCA string commitment.

Note that the first two rounds of the "trapdoor protocol" can be run with the last two rounds of the self-destruct commitment. Therefore the described construction has $t+2$ rounds (where t is the number of rounds of the self-destruct PCCA string commitment).

[7] By *delayed-input* we mean that the witness and the instance are needed only to play the last round.

Overview of the security proof. In the 1st experiment (the real game RG_0) the sender commits to m_0. We observe that due to the security of the signature scheme we can demonstrate that in the real game $\mathcal{A}$ is committing to a well-formed commitments in all parallel right sessions with non-negligible probability. Symmetrically there is the experiment RG_1 where the sender commits to m_1 and $\mathcal{A}$ is committing to a well-formed commitment in all parallel right sessions. Then we consider a hybrid game $\mathcal{H}_b^0$, for $b \in \{0,1\}$, where the sender commits to m_b and the oracle is emulated extracting the committed values from the special WIPoK. Note that $\mathcal{H}_b^0$ is distributed statistically close to RG_b until $\mathcal{A}$ receives the committed values, therefore we are ensured that we can extract the values committed in the right sessions. The 2nd hybrid game that we consider is $\mathcal{H}_b^1$ in which we switch the witness used to compute the transcript of the special WIPoK in the left sessions (i.e. we are using the trapdoor that is extracted by rewinding $\mathcal{A}$ in the left session). Using techniques that are similar to the one showed in [COSV17b] we are able to demonstrate that also in $\mathcal{H}_b^1$ we can extract the committed values in all parallel right sessions with non-negligible probability. Moreover, we can demonstrate that the distribution of the commitment values along with the view of $\mathcal{A}$ is indistinguishable between $\mathcal{H}_b^0$ and $\mathcal{H}_b^1$, for $b \in \{0,1\}$. Indeed, both in $\mathcal{H}_0^1$ and in $\mathcal{H}_1^1$ we are guaranteed that $\mathcal{A}$ is committing to a well-formed commitment in all parallel right sessions with non-negligible probability. Summing up, a detectable deviation from $\mathcal{H}_0^1$ and $\mathcal{H}_1^1$ implies a contradiction of the self-destruct PCCA security of the underlining commitment. Finally we observe that the extraction procedure of the signatures does not interfere with the reductions since in the parallel right sessions the commitment phase made by $\mathcal{A}$ ends in the third last round. This observation concludes the high-level overview of the security proof.

2 Preliminaries

We denote the security parameter by λ and use "||" as concatenation operator (i.e., if a and b are two strings then by $a||b$ we denote the concatenation of a and b). We use the abbreviation PPT that stands for probabilistic polynomial time. We use $\mathtt{poly}(\cdot)$ to indicate a generic polynomial function and $\mathbb{N}$ to denote the set of positive integer.

A *polynomial-time relation* Rel (or *polynomial relation*, in short) is a subset of $\{0,1\}^* \times \{0,1\}^*$ such that membership of (x,w) in Rel can be decided in time polynomial in $|x|$. For $(x,w) \in \mathsf{Rel}$, we call x the *instance* and w a *witness* for x. For a polynomial-time relation Rel, we define the $\mathcal{NP}$-language L_{Rel} as $L_{\mathsf{Rel}} = \{x | \exists w : (x,w) \in \mathsf{Rel}\}$.Analogously, unless otherwise specified, for an $\mathcal{NP}$-language L we denote by $\mathsf{Rel}_{\mathsf{L}}$ the corresponding polynomial-time relation (that is, $\mathsf{Rel}_{\mathsf{L}}$ is such that $L = L_{\mathsf{Rel}_{\mathsf{L}}}$). We denote by $\hat{L}$ the language that includes both L and all well formed instances that do not have a witness. Moreover we require that membership in $\hat{L}$ can be tested in polynomial time. We implicitly assume that a PPT algorithm that is supposed to receive an instance in $\hat{L}$ will abort immediately if the instance does not belong to $\hat{L}$. Let A and B be two interactive

probabilistic algorithms. We denote by $\langle A(\alpha), B(\beta)\rangle(\gamma)$ the distribution of B's output after running on private input β with A using private input α, both running on common input γ. Typically, one of the two algorithms receives 1^λ as input. A *transcript* of $\langle A(\alpha), B(\beta)\rangle(\gamma)$ consists of the messages exchanged during an execution where A receives a private input α, B receives a private input β and both A and B receive a common input γ. Moreover, we will refer to the *view* of A (resp. B) as the messages it received during the execution of $\langle A(\alpha), B(\beta)\rangle(\gamma)$, along with its randomness and its input. We say that a protocol (A, B) is public coin if B sends to A random bits only.

If $\mathcal{Z}$ is a set then $Z \leftarrow \mathcal{Z}$ will denote a random variable sampled uniformly from $\mathcal{Z}$. We start with some standard definitions and lemmas about the statistical distance. Recall that if X and X' are random variables over the same set $\mathcal{X}$ then the *statistical distance between X and X'* is denoted by $\Delta(X; X')$, and defined as $\Delta(X; X') = \frac{1}{2}\sum_{x\in\mathcal{X}} |\Pr X = x - \Pr X' = x|$. If the variables X and X' are such that $\Delta(X; X') \leq \epsilon$ then we say that X is ϵ-close to X', and write $X \approx_\epsilon X'$. If $\mathcal{E}, \mathcal{E}'$ are some events then by $\Delta(X|\mathcal{E}\ ;\ X'|\mathcal{E}')$ we will denote the distance between variables $\widetilde{X}$ and $\widetilde{X}'$, distributed according to the conditional distributions $P_{X|\mathcal{E}}$ and $P_{X'|\mathcal{E}'}$.

If $U_\mathcal{X}$ is the uniform distribution over $\mathcal{X}$ then $d(X|\mathcal{E}) := \Delta(X|\mathcal{E}; U_\mathcal{X})$ is called *statistical distance of X from uniform given the event $\mathcal{E}$*. Moreover, if Y is independent from X then $d(X|Y) := \Delta((X, Y); (U_\mathcal{X}, Y))$ is called *statistical distance of X from uniform given the variable Y*. More generally, if $\mathcal{E}$ is an event then $d(X|Y, \mathcal{E}) := \Delta((X, Y)|\mathcal{E}; (U_\mathcal{X}, Y)|\mathcal{E})$. It is easy to see that $d(X|Y)$ is equal to the average $\sum_y \Pr(Y = y) \cdot d(X|Y = y) = \mathbb{E}_y(d(X|Y = y))$.

Definition 1 ((Average-) Min-Entropy). *Let X have finite support $\mathcal{X}$. The min-entropy $\mathbf{H}_\infty(X)$ of X is defined by*

$$\mathbf{H}_\infty(X) = -\log \max_{x\in\mathcal{X}} \Pr(X = x).$$

For an event $\mathcal{E}$, the conditional min-entropy $\mathbf{H}_\infty(X|\mathcal{E})$ of X given $\mathcal{E}$ is defined by

$$\mathbf{H}_\infty(X|\mathcal{E}) = -\log \max_{x\in\mathcal{X}} \Pr(X = x|\mathcal{E}).$$

For an event $\mathcal{E}$ and a random variable Y with finite support $\mathcal{Y}$, the average min-entropy $\widetilde{\mathbf{H}}_\infty(X|Y, \mathcal{E})$ of X given Y and $\mathcal{E}$ is defined by

$$\widetilde{\mathbf{H}}_\infty(X|Y, \mathcal{E}) = -\log \mathbb{E}_y \max_{x\in\mathcal{X}} \Pr(X = x|Y = y, \mathcal{E}).$$

Randomness extractors will be the workhorses of our non-malleable code constructions.

Definition 2 (Flexible Two-Source Extractors). *A function $\mathsf{Ext} : \mathcal{X}_1 \times \mathcal{X}_2 \to \mathcal{Z}$ is called a flexible (ϵ, δ)-two-source extractor, if it holds for all tuples $((X_1, Y_1), (X_2.Y_2))$ for which (X_1, Y_1) is independent of (X_2, Y_2) and $\widetilde{\mathbf{H}}_\infty(X_1|Y_1) + \widetilde{\mathbf{H}}_\infty(X_2|Y_2) \geq \log(|\mathcal{X}|) + \log(|\mathcal{Y}|) - \delta$ that*

$$d(\mathsf{Ext}(X_1, X_2)|Y_1, Y_2) \leq \epsilon.$$

A well known example of a flexible two-source extractor is the Hadamard extractor or inner-product-extractor.

Lemma 1 (Hadamard Extractor [ADL14]). *The function* $\mathsf{Ext} : \mathbb{F}_q^n \times \mathbb{F}_q^n \to \mathbb{F}_q$ *given by* $\mathsf{Ext}(x, y) = \langle x, y \rangle$ *is a flexible* (ϵ, δ) *extractor for* $\delta \leq (n-1)\log(q) - 2\log(1/\epsilon)$.

Lemma 2 (Entropy-preservation of inner-product for correlated distributions). *Let* X *be random variable over* $\mathcal{X}^l$*, let* C *be random variable such that for every* c *we have* $\mathbf{H}_\infty(X|C = c) \geq l \cdot \log|\mathcal{X}| - d$*, where* $d < \log|\mathcal{X}|$*. Then for any non-zero* $v \in \mathcal{X}^l$

$$\mathbf{H}_\infty(\langle X, v \rangle_{\mathcal{X}} \mid C = c) \geq \log|\mathcal{X}| - d$$

for every c *in* $\mathsf{supp}(C)$.

We will now assemble a few basic technical lemmata that we will need for our proofs.

Lemma 3 (Bayes' rule for statistical distance [DKO13]). *Let* $(X, Y) \in \mathcal{X} \times \mathcal{Y}$ *be a random variables, such that* $d(X|Y) \leq \epsilon$*. Then for every* $x \in \mathcal{X}$ *we have*

$$\Delta(Y|X = x \ ; \ Y) \leq 2|\mathcal{X}|\epsilon.$$

Also if $\mathcal{A}$ *is a random event such that* $d(X|Y, \mathcal{A}) \leq \epsilon$*, we have:*

$$\Delta(Y|X = x, \mathcal{A} \ ; \ Y|\mathcal{A}) \leq 2|\mathcal{X}|\epsilon.$$

Lemma 4 ([DNO16]). *Let* X, T *be any arbitrarily correlated random variables and let* $\mathcal{E}$ *be random event then*

$$\widetilde{\mathbf{H}}_\infty(X|T, \mathcal{E}) \geq \widetilde{\mathbf{H}}_\infty(X|T) - \log\frac{1}{\Pr(\mathcal{E})}.$$

In the Appendix A the reader can find a series of standard definitions used in the rest of the paper.

2.1 Commitment Schemes

We consider the standard definition of commitment scheme that it is possible to find in Appendix A.

Self-Destruct CCA Secure Commitment Schemes. Let $\Pi = (\mathsf{Sen}, \mathsf{Rec})$ be a commitment scheme. The self-destruct CCA-oracle $\mathcal{O}^{\mathsf{sdcca}}$ for $\Pi = (\mathsf{Sen}, \mathsf{Rec})$ acts as follows in an interaction with an adversary $\mathcal{A}$: it participates with $\mathcal{A}$ in polynomially many sessions of the commit phase of Π as an honest receiver. At the end of each session, if the session is valid, the oracle returns the unique value m committed in the interaction. The oracle outputs $\perp$ and implements the

self-destruct mode, (i.e. the oracle will respond with $\perp$ for all subsequent commitment queries) if one of the following cases happen: 1) a session has multiple valid committed values[8]; 2) the commitment is invalid; 3) if the committed value m is equal to a special self-destruct symbol $\perp$.

More precisely, let us consider the following probabilistic experiment $\mathsf{IND}_b^{\mathsf{sdcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})$ Let $\mathcal{O}^{\mathsf{sdcca}}$ be the SD CCA-oracle for Π. The adversary has access to $\mathcal{O}^{\mathsf{sdcca}}$ during the entire course of the experiment. On input 1^λ, and $z \in \{0,1\}^\star$ the adversary $\mathcal{A}^{\mathcal{O}^{\mathsf{sdcca}}}$ sends two strings m_0 and m_1 with $|m_0| = |m_1|$ to the experiment. The experiment randomly selects a bit $b \leftarrow \{0,1\}$ and commits to m_b to $\mathcal{A}^{\mathcal{O}^{\mathsf{sdcca}}}$. Note that if $\mathcal{A}$ queries the oracle with a commitment of m s.t. $m \in \{m_0, m_1\}$ [9] then, the oracle returns the special symbol `same`. Finally $\mathcal{A}^{\mathcal{O}^{\mathsf{sdcca}}}$ sends a bit y to the experiment. The output of the experiment is replaced by $\perp$ if $\mathcal{A}^{\mathcal{O}^{\mathsf{sdcca}}}$ sends a commitment to $\mathcal{O}^{\mathsf{sdcca}}$ whose transcript is identical to the one computed on the left. Otherwise, the output of the experiment is y. Let $\mathtt{IND}_b^{\mathsf{sdcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})$ denote the output of the experiment described above.

Definition 3 (Self-destruct CCA (SD-CCA) secure string commitment scheme). *Let $\Pi(\mathsf{Sen}, \mathsf{Rec})$ be a commitment scheme and $\mathcal{O}^{\mathsf{sdcca}}$ be the self-destruct CCA-oracle for Π_{sdcca}. We say that Π_{sdcca} is self-destruct CCA-secure (w.r.t. the committed-value oracle), if for every ppt-adversary $\mathcal{A}$ and all $z \in \{0,1\}^\star$ it holds that:*

$$\{\mathtt{IND}_0^{\mathsf{sdcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})\} \approx \{\mathtt{IND}_1^{\mathsf{sdcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})\}$$

Definition 4 (Self-destruct parallel CCA (SD-PCCA) secure string commitment scheme). *The self-destruct parallel CCA oracle is defined like the self-destruct CCA-oracle, except that the adversary is restricted to a parallel query, i.e., the adversary can only send a single query that may contain multiple commitments sent in parallel. Let $\mathtt{IND}_b^{\mathsf{sdpcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})$ define the output of the security game for self-destruct parallel CCA security. The formal definition is then analogous to the definition of SD-CCA security.*

Note that any SD-CCA commitment scheme is also a SD-PCCA commitment scheme.

Definition 5 (Parallel CCA secure (PCCA) string commitment scheme[BFMR18, Kiy15]). *The parallel CCA oracle is defined like self-destruct parallel CCA-oracle, except that the oracle does not implement the self-destruct mode. In more details, when a commitment is not valid, or a session has multiple valid committed values the oracle returns $\perp$, and the committed messages (or the symbol* `same`*)*

[8] The statistical binding property guarantees that this happens with only negligible probability.

[9] As noted in [AGM+15], following [DDN91], this definition allows MIM to commit to the same value. It is easy to prevent MIM from committing the same value generically in case of string commitments: convert the scheme to tag based by appending the tag with v, and then sign the whole transcript using the tag.

in all the other cases. Let $\mathtt{IND}_b^{\mathsf{sdpcca}}(\Pi = (\mathsf{Sen}, \mathsf{Rec}), \lambda, z, \mathcal{A})$ *define the output of the security game for parallel CCA security (PCCA). The formal definition is then analogous to the definition of SD-PCCA security.*

In this paper we also consider a self-destruct (parallel) CCA secure bit commitment scheme that is defined as in Def. 3 (4), except that the message space is $\{0,1\}$ and the oracle never returns `same`.

In all the paper we denote by $\tilde{\delta}$ a value associated with the right session (where the adversary $\mathcal{A}$ plays with the oracle) where δ is the corresponding value in the left session. For example, the sender commits to v in the left session while $\mathcal{A}$ commits to $\tilde{v}$ in the right session.

3 Definitions related to Non-Malleable Codes

Definition 6 (Coding Schemes). *A coding scheme is a pair* $(\mathsf{Enc}, \mathsf{Dec})$*, where* $\mathsf{Enc} : \mathcal{M} \to \mathcal{C}$ *is a randomized function and* $\mathsf{Dec} : \mathcal{C} \to \mathcal{M} \cup \{\perp\}$ *is a deterministic function, such that it holds for all* $M \in \mathcal{M}$ *that* $\mathsf{Dec}(\mathsf{Enc}(M)) = M$*.*

Definition 7 (Two-State Code). *A coding scheme* $(\mathsf{Enc}, \mathsf{Dec})$ *where the counterdomain of* Enc *has the form* $\mathcal{C} = \{0,1\}^k \times \{0,1\}^k$ *is called a* two-state *code.*

Definition 8 (Paritied Two-State Code). *Let* $(\mathsf{Enc}, \mathsf{Dec})$ *(where* $\mathsf{Enc} : \mathcal{M} \to \mathcal{C} = \mathcal{C}_1 \times \mathcal{C}_2 = \{0,1\}^k \times \{0,1\}^k$*) be a two-state code. Now let* $\mathsf{Enc}^{par} : \mathcal{M} \to \mathcal{C}$ *be a randomized function restricted to a condition that* $parity(\mathsf{Enc}(m)_1) = 0$ *and* $parity(\mathsf{Enc}(m)_2) = 1$*, where parity is a function calculating the parity of number of ones in a given vector (i.e.* $parity(0101011) = 0$ *and* $parity(011111) = 1$*).*

More formally, the procedure computing $\mathsf{Enc}^{par}(m)$ *can be described as follows: we run in a loop the encoding procudure* $(c_1, c_2) \leftarrow \mathsf{Enc}(m)$ *until* $parity(c_1) = 0$ *and* $parity(c_2) = 1$*.*

Similarily, let $\mathsf{Dec}^{par} : \mathcal{C} \to \mathcal{M} \cup \{\perp\}$ *be defined as follows: for* $c = (c_1, c_2) \in \mathcal{C}$*, if* $parity(c_1) \neq 0$ *or* $parity(c_2) \neq 1$ *then* $\mathsf{Dec}^{par}(c) := \perp$*, otherwise* $\mathsf{Dec}^{par}(c) := \mathsf{Dec}(c)$*.*

Now, the coding scheme $(\mathsf{Enc}^{par}, \mathsf{Dec}^{par})$ *is called a* paritied *two-state code.*

We will now define the continuous tampering experiment. Our definition is a weaker version of [JW15]: instead of Super Strong Tampering experiment we will use the standard tamper experiment from [DPW10].

Definition 9 ((Continuous-) Tampering Experiment). *We will define continuous non-persistent self-destruct non-malleable codes using [AKO17] experiment, which is equivalent to original [DPW10] experiment. Fix a coding scheme* $(\mathsf{Enc}, \mathsf{Dec})$ *with message space* $\mathcal{M}$ *and codeword space* $\mathcal{C}$*. Also fix a family of functions* $\mathcal{F} : \mathcal{C} \to \mathcal{C}$*. Let* $\mathcal{D} = \{\mathcal{D}_C^f\}_{f \in \mathcal{F}, C \in \mathcal{C}}$ *be some family of distributions over* $\{0,1\}$*, indexed by tampering function* f *and a codeword* C*. We will first define the tampering oracle* $\mathsf{Tamper}_{C,\mathcal{D}}^{\mathsf{state}}(f)$*, for which initially* $\mathsf{state} = \mathtt{alive}$*. For a tampering function* $f \in \mathcal{F}$ *and a codeword* $C \in \mathcal{C}$ *define the tampering oracle by*

$\mathsf{Tamper}^{\mathsf{state}}_{C,\mathcal{D}}(f)$:
If $\mathsf{state} = \mathtt{dead}$ *output* $\perp$
$C' \leftarrow f(C)$
If $\mathsf{Dec}(C') = \mathsf{Dec}(C)$ *and* $\mathcal{D}^f_C = 0$ *output* `same`
$M' \leftarrow \mathsf{Dec}(C')$
If $M' = \perp$ *set* $\mathsf{state} \leftarrow \mathtt{dead}$ *and output* $\perp$
Otherwise output C'

Fix a tampering adversary $\mathcal{A}$ *and a codeword* $C \in \mathcal{C}$*. We define the continuous tampering experiment* $\mathsf{CT}_{C,\mathcal{D}}(\mathcal{A})$ *by*

$\mathsf{CT}_{C,\mathcal{D}}(\mathcal{A})$:
$\mathsf{state} \leftarrow \mathtt{alive}$
$v \leftarrow \mathcal{A}^{\mathsf{Tamper}^{\mathsf{state}}_{C,\mathcal{D}}(\cdot)}$
Output v

Definition 10. *Let* $(\mathsf{Enc}, \mathsf{Dec})$ *be a coding scheme and* CT *be its corresponding continuous tampering experiment for a class* $\mathcal{F}$ *of tampering functions. We say that* $(\mathsf{Enc}, \mathsf{Dec})$ *is an* ϵ*-secure continuously non-malleable code against* $\mathcal{F}$*, if there exists a family of distributions* $\mathcal{D} = \{\mathcal{D}^f_C\}_{f\in\mathcal{F}, C\in\mathcal{C}}$ *over* $\{0,1\}$ *such that for all tampering adversaries* $\mathcal{A}$ *and all pairs of messages* $M_0, M_1 \in \mathcal{M}$ *that*

$$\mathsf{CT}_{C_0,\mathcal{D}}(\mathcal{A}) \approx_\epsilon \mathsf{CT}_{C_1,\mathcal{D}}(\mathcal{A}),$$

where $C_0 \leftarrow \mathsf{Enc}(M_0)$ *and* $C_1 \leftarrow \mathsf{Enc}(M_1)$*.*

4 Continuous Non-Malleable Code against Permutations-With-Overwrites

In this section we define a coding scheme $(\mathsf{Enc}_c, \mathsf{Dec}_c)$ and prove it is a continuous non-malleable code against a class $\mathsf{PermOver}$ of permutations-with-overwrites (the actual definition will follow).

4.1 Coding scheme

Let $\mathcal{M} = \{0,1\}^n$ and $\mathcal{C} = \mathcal{C}_1 \times \cdots \times \mathcal{C}_N$, where each $\mathcal{C}_i = \{0,1\}^{k_1} \times \{0,1\}^{k_1}$. Let also $(\mathsf{Enc}_2, \mathsf{Dec}_2)$ denote a two-state code (actually we need a two-state strong non-malleable code here, however the specific instantiation will be given later) and h_N denote a $\lfloor N/3 \rfloor$-out-of-N secret sharing scheme (again, the specific instantiation will be given later). Now we are ready to introduce the (randomized) function (procedure) $\mathsf{Enc}_c : \mathcal{M} \to \mathcal{C}$:

For $m \in \mathcal{M}$ and a random $r \in \{0,1\}^n$, let $(d_1, \ldots, d_N) \leftarrow h_N(m||r)$ where $(d_1, \ldots, d_N) \in (\{0,1\}^{k_2})^N$ are shares for $(m||r)$. Now, for each d_i let $(L_i, R_i) \leftarrow \mathsf{Enc}^{par}_2(d_i||i)$.

Finally, we state $c_i \leftarrow (L_i, R_i)$ and $\mathsf{Enc}_s(m)$ outputs $(c_1, \ldots, c_N)$.

The definition of Dec_c is simple and straightforward (forced by the definition of a coding scheme).

Remark 1. The above construction is not tight for a given message length n since it also depends on the choice of parameters (N, k_1, k_2) and the specific definitions of both: the two-state code $(\mathsf{Enc}_2, \mathsf{Dec}_2)$ and the secret sharing scheme h_N. However, before we pick adequate parameters and schemes, we need one definition more:

Definition 11. *We call a two-split code* $(\mathsf{Enc}_2, \mathsf{Dec}_2)$ ϵ-admissible *if the scheme* $(\mathsf{Enc}_2^{par}, \mathsf{Dec}_2^{par})$ *fulfills the following requirements:*

1. *[Canonical encoding procedure:]* $\mathsf{Enc}_2^{par}(m)$ *is uniform in* $\{c : \mathsf{Dec}_2^{par}(c) = m\}$.
2. *[Detection of close to bijective tampering:] For any message* m*, if* $\mathsf{Enc}_2^{par}(m) = (X, Y)$ *then for any functions* $f, g : \{0,1\}^{k_1} \to \{0,1\}^{k_1}$ *such that* $\mathbf{H}_\infty(f(X)), \mathbf{H}_\infty(g(Y)) \geq 2/3 \cdot k_1 - 1$ *and (for any* x *or* y*)* $f(x) \neq x$ *or* $g(y) \neq y$ *it holds:*

$$\Pr(\mathsf{Dec}_2^{par}(f(X), g(Y)) = \perp) \geq 1 - \epsilon.$$

3. *[Detection of complete overwrite of one part:] For any constant* $c \in \{0,1\}^{k_1}$*, and any uniform* $X, Y \in \{0,1\}^{k_1}$*, such that parity of* X *is* 0 *and parity of* Y *is* 1 *we get,*

$$\Pr(\mathsf{Dec}_2(X, c) = \perp) \geq 1 - \epsilon,$$
$$\Pr(\mathsf{Dec}_2(c, Y) = \perp) \geq 1 - \epsilon$$

4. *[Leakage resilient storage:] For any message* m*, if* $\mathsf{Enc}_2^{par}(m) = (X, Y)$ *then for any functions* $f, g : \{0,1\}^{k_1} \to \{0,1\}^{k_1}$ *such that* $\widetilde{\mathbf{H}}_\infty(X|f(X)) \geq 1/3 \cdot k_1$ *and* $\widetilde{\mathbf{H}}_\infty(Y|f(Y)) \geq 1/3 \cdot k_1$ *we get*

$$\Delta\left[(f(X), Y); (f(U_0), U_1)\right] \leq \epsilon\ ,$$
$$\Delta\left[(X, g(Y)); (U_0, g(U_1))\right] \leq \epsilon\ ,$$

where U_0, U_1 *are independent uniformly distributed over* $\{0,1\}^{k_1}$*, such that parity of* U_i *is equal* i*.*

An instantiation of the above definition for an appropriate ϵ_c is described in the Appendix B. Through the rest of the paper we always refer to this specific two-state code and the specific error probability when notation $(\mathsf{Enc}_2, \mathsf{Dec}_2)$ and ϵ_c is used.

4.2 Definition of the class of tampering functions

Here we define the class PermOver of tampering functions. Through this paper functions from this class PermOver are called permutations-with-overwrites.

Let us consider a set $\{0,1\}^q$ of vectors of q bits (q-vectors, for short). Now, let denote Π_q the class of permutations of bits of q-vectors. Denote also O_q the class of functions $f : \{0,1\}^q \to \{0,1\}^q$, such that:

for all i, either $f(x)_i = x_i$ or $f(x)_i = b_i$ for a fixed b_i.

Loosly speaking: any function from O_q, independently for each bit, either leaves it unchanged or sets it into a fixed value (i.e. overwrites it).

Now we simply define the class $\mathsf{PermOver}_q = O_q \circ \Pi_q$. For our application we will equate $\mathcal{C} = (\{0,1\}^{k_1} \times \{0,1\}^{k_1})^N$ with $\{0,1\}^{2k_1 N}$ and consider $\mathsf{PermOver} = \mathsf{PermOver}_{2k_1 N}$ as a tampering class for $\mathcal{C}$.

The above description of course finishes the definition of our class of tampering functions, however we want a few further related definitions.

Related definitions. Let us fix a tampering function $t \in \mathsf{PermOver}$. As mentioned above we will think of t as a function from $\mathcal{C}_1 \times \cdots \times \mathcal{C}_N$ to $\mathcal{C}_1 \times \cdots \times \mathcal{C}_N$. Now, for each $i \in \{1, \ldots, N\}$ we say that t either *leaves* or *overwrites* or *modifies* the i-th *block*. These phrases stand for the following:

If $t(c)_i = c_i$ then t leaves the i-th block. If $t(c)_i = a$ for some a independent of c then t overwrites the i-th block. Finally, if none of the previous occurs, then we say that t modifies the i-th block.

If t overwrites i-th block, two cases are possible. Either c_i is independent of $f(c)$ or some bits of c_i are moved to some modified blocks. In the first case we say that t strong-overwrites i-th block and in the second case, it weak-overwrites.

Touched blocks are blocks either modified or weak-overwritten. In that case we say that t *touches* these blocks.

For a function $t \in \mathsf{PermOver}$ and a codeword $c \in \mathcal{C}$ we denote $touch(t, c)$ the set of all touched blocks and its indices, more formally: $touch(t, c) = \{(c_i, i) | \text{t touches } c_i\}$.

Example. The above definitions may look a little bit obscure at first sight, so – to make things clearer – we give an example.

Let $N = 4$ and each $\mathcal{C}_i = \{0,1\}^6$. Now let us consider:

$$t\Big((b_1^1, b_2^1, b_3^1, b_4^1, b_5^1, b_6^1), (b_1^2, b_2^2, b_3^2, b_4^2, b_5^2, b_6^2), (b_1^3, b_2^3, b_3^3, b_4^3, b_5^3, b_6^3), (b_1^4, b_2^4, b_3^4, b_4^4, b_5^4, b_6^4)\Big) =$$
$$\Big((0,0,0,1,1,1), (b_1^2, b_2^2, b_3^2, b_4^2, b_5^2, b_6^2), (0,1,0,1,0,1), (0, b_5^1, b_4^4, 1, b_2^4, b_1^1)\Big).$$

Obviously $t \in \mathsf{PermOver}$ and we have that: t leaves the second block, overwrites the first and the 3-rd block and modifies the 4-th block. The first block is weak-overwritten (because the 5-th block gets one bit from the first block) and the 3-rd block is strongly overwritten. Function t touches the blocks of the indices 1 and 4 so, for exemplary

$$c = ((0,0,1,1,0,0), (0,0,1,1,1,1), (1,1,1,1,0,0), (1,0,0,0,0,1)),$$

we have:

$$touch(t, c) = \{\big((0,0,1,1,0,0), 1\big), \big((1,0,0,0,0,1), 4\big)\}.$$

4.3 Statement and Proof

The main statements for the whole Section 4 are the following:

Theorem 1. *The coding scheme* $(\mathsf{Enc}_c, \mathsf{Dec}_c)$ *is an* $(\alpha + 2\epsilon_c)^{\lfloor N/3 \rfloor}$*-secure continuous non-malleable code against* PermOver *for* $\alpha = (0.5)^{\frac{1}{8 \cdot k_1}}$.

Corollary 1. *Instantiation for the above code with* $(N, k_2, k_1) = (6\lceil n^{2/3} \rceil, \lceil n^{1/3} \rceil, c\lceil n^{1/3} \rceil \log \lceil n^{1/3} \rceil)$, *with* $(\mathsf{Enc}_2, \mathsf{Dec}_2) = (\mathsf{Enc}_{\mathsf{Li}}, \mathsf{Dec}_{\mathsf{Li}})$ *(see Appendix B) and* $h_N = RS_N$ *(see Appendix C) gives us a continuous non-malleable code against* PermOver *such that:*

- *the code rate is* $O(\log n)$*, and*
- *the error rate is* $O(2^{-O(n^{1/3})})$.

Proof. The message length is n and the codeword length is $N \cdot 2 \cdot k_1 \approx 6n^{2/3} \cdot 2 \cdot cn^{1/3}\frac{1}{3}\log n = 4cn \log n$, so the code rate is approximately $4c \log n = O(\log n)$. (Remark: c is a constant from $\mathsf{Enc}_{\mathsf{Li}}$ rate.) The error rate is:

$$(\alpha+2\epsilon_c)^{\lfloor N/3 \rfloor} = ((0.5)^{\frac{1}{8 \cdot k_1}} + 2\epsilon_c)^{\lfloor N/3 \rfloor} \leq (2^{-O\left(\frac{1}{n^{1/3} \log n}\right)} + 2^{-O(n)})^{n^{2/3}+1} = 2^{-O(n^{1/3})}.$$

Before the actual proof of Theorem 1 we want to introduce a slightly modified version of continuous tampering experiment for $(\mathsf{Enc}_c, \mathsf{Dec}_c)$ and PermOver and a definition of a specific type of distribution that we call block-wise distribution.

The described below experiment is obviously stronger (from adversary's point of view) then the original one so it is sufficient to prove that our coding scheme is secure against PermOver for the modified experiment:

Definition 12 ((Modified) Continuous Tampering Experiment). *Let us consider a tampering oracle* $\mathsf{ModTamp}^{\mathsf{state}}_C(t)$*, for which initially* $\mathsf{state} = \texttt{alive}$*. For a tampering function* $t \in$ PermOver *and a codeword* $C \in \mathcal{C}$ *define the tampering oracle by*

$\mathsf{ModTamp}^{\mathsf{state}}_C(t)$:
If $\mathsf{state} = \texttt{dead}$ *output* $\perp$
$C' \leftarrow t(C)$
If $\mathsf{Dec}_c(C') = \mathsf{Dec}_c(C)$ *output* $(\texttt{same}, touch(t, c))$
$M' \leftarrow \mathsf{Dec}_c(C')$
If $M' = \perp$ *set* $\mathsf{state} \leftarrow \texttt{dead}$ *and output* $\perp$
Otherwise output C'

Fix a tampering adversary $\mathcal{A}$ *and a codeword* $C \in \mathcal{C}$*. We define the (modified) continuous tampering experiment* $\mathsf{MCT}_C(\mathcal{A})$ *by*

$\mathsf{MCT}_C(\mathcal{A})$:
$\mathsf{state} \leftarrow \texttt{alive}$
$v \leftarrow \mathcal{A}^{\mathsf{ModTamp}^{\mathsf{state}}_C(\cdot)}$
Output v

Remark 2. The main difference of the above experiment and the original one is the output of the oracle when $\mathsf{Dec}_c(C') = \mathsf{Dec}_c(C)$. In this case in our definition we give the adversary additionally all touched blocks.

Definition 13 (Block-wise Distribution). *For $\mathcal{C} = \mathcal{C}_1 \times \ldots \times \mathcal{C}_N$ the distribution D over $\mathcal{C}$ is a* block-wise distribution *if (informally speaking) each block $\mathcal{C}_i$ is either fixed or uniform and independent of the other blocks.*

Formally, we say that D is a block-wise distribution if there exists a set of indices $I \subset [1, 2, \ldots, N]$ such that for all $i \in I$ there exists $c_i \in \mathcal{C}_i$ such that:

$P_D(\mathcal{C}_i = c_i) = 1$, *and*

the conditional distribution $(D|\mathcal{C}_i = c_i$ for all $i)$ is uniform.

Remark 3. If $|I| = l$ in the above definition, then we will sometimes say that D *has l constant blocks* or that the adversary *knows l blocks.*

Proof skeleton for Theorem 1. Our key observation is that after each oracle call in the tampering experiment, the distribution of the codewords (from the perspective of the adversary) is almost always block-wise. Moreover, to increase the number of known (constant) blocks, the adversary must take a risk of receiving $\perp$. This idea is expressed in the following Lemma 5. Notice, that from basic properties of secret sharing schemes, the tampering experiment is independent from the message m while the number of known blocks is smaller then $\lfloor N/3 \rfloor$. So, the only way for the adversary to distinguish between two different messages is to learn at least $\lfloor N/3 \rfloor$ blocks. However (from Lemma 5) this happens with probability at most $(\alpha + 2\epsilon_c)^{\lfloor N/3 \rfloor}$ (for $\alpha = (0.5)^{\frac{1}{8 \cdot k_1}}$) so this observation finishes the proof for Theorem 1. □

Before the statement of the key Lemma 5, we need one definition more:

Definition 14. *For a block-wise distribution D and a tampering function $t \in$* PermOver *we say that t* freshly-touches *the i-th block if t touches this block and this block is not known in context of D.*

Lemma 5. *Let $\alpha = (0.5)^{\frac{1}{8 \cdot k_1}}$, let $l_1, l_2 \in \mathbb{N}$ such that $l_1 + l_2 < \lfloor N/3 \rfloor$, and let D be a block-wise distribution over C with l_1 constant blocks and let $t \in$ PermOver be a tampering function freshly-touching l_2 blocks. Then, with probability at least $(1 - (\alpha + 2\epsilon_c)^{l_2})$ a call $\mathsf{ModTamp}_C^{\mathsf{state}}(t)$ will return $\perp$. Moreover – with probability at least $(1 - 2^{-n})$ – the distribution D conditioned on the answer from the oracle will be block-wise with $l_1 + l_2$ constant blocks.*

Proof. The proof for the first part of the statement is moved to the Lemma 6. (Notice that in that lemma we do not consider distributions with constant blocks and the concept of freshly-touched blocked, but we only work with uniform distribution of C. However this is not a real limitation since we can easily translate a tampering function $t \in$ PermOver into an equivalent function $t' \in$ PermOver that is independent from known blocks and we can just restrict the domain to not known blocks.)

To prove the second part we will consider three cases (dependent on the properties of the tampering function t):

case 1: t overwrites at least one block and leaves at least one block. In this case we claim that the adversary gets $\perp$ with probability at least $(1 - 2^{-n})$. Here is the argument:

The total number of either overwritten or left blocks is at least $\frac{2}{3}N$. It means that at least one type of these blocks is a set of size at least $\lfloor N/3 \rfloor$. Without loss of generality we can assume that the number of left blocks is at least $\lfloor N/3 \rfloor$. This means that (from properties of secret sharing), if the adversary does not get $\perp$, all other blocks must be consistent with the shared value $(m||r)$. However for the adversary the value of $r \in \{0,1\}^n$ is completely random so the probability of setting a proper value for a single overwritten block is exactly 2^{-n}.

case 2: t does not overwrite any block. In this case most (at least $\frac{2}{3}N$) of blocks are left. So, if the adversary does not receive bottom then (from properties of secret sharing) she receives `same` so she learns all l_2 freshly modified blocks and the new distribution of the codeword is as described.

case 3: t does not leave any block. In this case most (at least $\frac{2}{3}N$) of blocks are overwritten. If the adversary does not receive bottom she will simply learn all l_2 not overwritten blocks and nothing more. The new distribution of the codeword is obviously as described.

4.4 Technical lemmata

The main result for this section is the following lemma:

Lemma 6. *Let $t \in \mathsf{PermOver}$ touch l blocks. Then, for a random encoding $C \in \mathcal{C}$ of any fixed message m, the probability that the oracle $\mathsf{ModTamp}_C^{\mathsf{alive}}(t)$ will not return $\perp$ is at most $\alpha^{\max\{1,l\}} + 2\epsilon_c$, for $\alpha = (0.5)^{\frac{1}{8 \cdot k_1}}$.*

Before the proof, we need a bunch of auxiliary technical lemmata and definitions:

Lemma 7. *If X is uniform over $\{0,1\}^{k_1}$, such that parity of X is fixed to 0 (or 1), and $h \in \mathsf{PermOver}$ is such that given $h(X)$ at least one bit of X is unknown then, parity of $h(X)$ is unknown (i.e. is uniform).*

Proof of above is trivial and we omit it.

Lemma 8 (Single Block Attack). *For any fixed c let $\mathsf{Enc}_2^{par}(c) = (X, Y)$. For any deterministic function $h \in \mathsf{PermOver}_{2k_1}$ if $h \not\equiv id$ and $h \not\equiv const.$ following is true*

$$\Pr(\mathsf{Dec}_2(h(X,Y)) \neq \perp) \leq 0.5 + \epsilon_c.$$

Proof of above can be found in Appendix D.

For the next lemmata and definitions we carry assumption of Lemma 6:

Symbol t stands for the tamepering function from $\mathsf{PermOver}$ and $(C_1, ..., C_N)$ stands for the N blocks of the encoding of m. Also let $C'_1, ..., C'_N$ denote the values of these blocks after t is applied.

Definition 15. *Let $\mathsf{Transfer}_{\mathsf{i}\to\mathsf{j}}$ be a vector of bits that were permuted from block C_i into block C_j (if bit was permuted and then overwrote we do not include him). Let $S_i = \{j \mid |\mathsf{Transfer}_{\mathsf{i}\to\mathsf{j}}| \neq 0\}$ is a set that keeps track of blocks to which bits from C_i were distributed.*

Definition 16. *We will say that* touched *block C_i is:*

$\mathsf{Touched}_{\mathsf{Spread}}$ *if* $|S_i| \geq 8$
$\mathsf{Touched}_{\mathsf{Permuted}}$ *if* $|S_i| = 1$ *and tampering function preserves block C_i and moves it to $j-$th position, i.e. always $C'_j = C_i$,*
$\mathsf{Touched}_{\mathsf{Concentrated}}$ *it is not* $\mathsf{Touched}_{\mathsf{Spread}}$ *or* $\mathsf{Touched}_{\mathsf{Permuted}}$.

Notice that above cases exhaust all possibilities in which the block can be touched.

Lemma 9. *If C_i is* $\mathsf{Touched}_{\mathsf{Spread}}$ *then there exists set $\tilde{S}_i \subset S_i$ such that:*

1. $|\tilde{S}_i| \geq \lfloor \frac{|S_i|}{4} \rfloor \geq \frac{|S_i|}{8}$,
2. *all together $\{C'_j\}_{j \in \tilde{S}_i}$ contain at most half of bits of L_i and at most half of bits of R_i.*

Lemma 10. *Let C_i be* $\mathsf{Touched}_{\mathsf{Spread}}$ *and let $\tilde{S}_i$ be as in lemma 9. Then for any evaluation of $C_1, \ldots, C_{i-1}, C_{i+1}, \ldots, C_N$, and for any value d_i we get that parities of all blocks $\{C'_j\}_{j \in \tilde{S}_i}$ are ϵ_c-close to uniform under the randomness of C_i.*

Proof. Since $\{C'_j\}_{j \in \tilde{S}_i}$ together contain at most half of bits of L_i and at most half of bits of R_i, by Definition 11.4 we get

$$\mathsf{Transfer}_{\mathsf{i} \to \mathsf{j}_1} || \ldots || \mathsf{Transfer}_{\mathsf{i} \to \mathsf{j}_{|\tilde{S}_i|}} \approx_{\epsilon_c} U$$

where U is uniformly distributed over $\{0,1\}^{|\tilde{S}_i|}$. Notice that once $d_1, ..., d_N$ are all fixed, the random variables $C_1, ..., C_N$ are independent. Thus every block C_j for $j \in \tilde{S}_i$ got some uniformly random, independent bits, which ends the proof that parity of these blocks is independent and unpredictable.

Lemma 11. *Let $d_1, ..., d_N$ be fixed, and let $C = (C_1, ..., C_N)$ be a codeword encoding $(d_1, ..., d_N)$. If tampering function t is such that there is at least one block C_i which is* $\mathsf{Touched}_{\mathsf{Permuted}}$ *then*

$$\mathsf{ModTamp}_C^{\mathsf{alive}}(t) = \perp$$

Proof. Let $j \in S_i$ then $C'_j = C_i$, but after decoding C'_j will hold an position index i instead of j thus decoder will detect this tampering.

Finally, we are for the proof of Lemma 6:

Proof (of Lemma 6). First let us observe that once $d_1, ..., d_N$ are all fixed, the random variables $C_1, ..., C_N$ are independent. Let us notice that if there are any $\mathsf{Touched}_{\mathsf{Permuted}}$ blocks then by lemma 11 we immediately get the thesis of the lemma, thus from now on we will assume there are not $\mathsf{Touched}_{\mathsf{Permuted}}$ blocks. When adversary touches l blocks, he risks getting detected (i.e. creating invalid codeword). He has to fulfill at least l of such checks. If they were independent and probability of detection were 0.4 for each of them, then easily we would get that probability of whole bundle being valid is at most 0.6^l. Sadly the checks

are not independent[10]. However below we argue that many checks are indeed independent and we count their number. We proceed with following procedure, let $r = 0$ let at start $\mathcal{C} = \{1, ..., N\}$ we will pick smallest $i \in \mathcal{C}$ and :

block C_i was $\mathsf{Touched}_{\mathsf{Concentrated}}$: let $j \in S_i$. By lemma 8 we know that probability that block C'_j is valid is at most $0.5 + \epsilon_c$ and this probability is independent of all other blocks.
Remove i from $\mathcal{C}$.
Remove from $\mathcal{C}$ all j such that $|S_j \cap S_i| \neq 0$. By removing all j we obtain independence between all validity checks which we are counting. Notice that we removed at most $k_1 \cdot |S_i| \leq 8 \cdot k_1$ elements from $\mathcal{C}$.
Increase $r = r + 1$.

block C_i was $\mathsf{Touched}_{\mathsf{Spread}}$: by lemma 10 we know that probability that all block $\{C'_j\}_{j \in \tilde{S}_i}$ will be valid is at most $(0.6)^{|\tilde{S}_i|}$, which by lemma 9 can be bounded by $(0.5)^{\frac{|S_i|}{8}} + \epsilon_c$.
Remove i from $\mathcal{C}$, also remove all j such that $|S_j \cap S_i| \neq 0$. By removing all j we obtain independence between all validity checks which we are counting. Notice that we removed at most $k_1 \cdot |S_i|$ elements from $\mathcal{C}$.
Increase $r = r + \lfloor \frac{|S_i|}{4} \rfloor$ (mind that $\lfloor \frac{|S_i|}{4} \rfloor \geq \frac{|S_i|}{8} \geq 1$) .

block C_i was not touched: Remove i from $\mathcal{C}$.

Notice that every time we increase counter r we also remove from $\mathcal{C}$ all blocks that could have created correlated validity checks. Also notice that whenever we removed $k_1 \cdot x$ touched elements from $\mathcal{C}$ we increased the counter r by at least $\max\{1, \frac{x}{8}\}$, thus $r \geq \frac{l}{8 \cdot k_1}$. Notice that we have at least r independent checks, each passing with probability at most $\frac{1}{2} + \epsilon_c$. By simple induction we can show that $(\frac{1}{2} + \epsilon_c)^r \leq (\frac{1}{2})^r + 2\epsilon_c$ as long as $\epsilon_c < \frac{1}{4}$. Thus we get our thesis for $\alpha = (0.5)^{\frac{1}{8 \cdot k_1}}$.

5 SD-CCA Commitment Scheme from NMCode

In this section we describe our $\Pi_{\mathsf{sdcca}} = (\mathsf{Sen}_{\mathsf{sdcca}}, \mathsf{Rec}_{\mathsf{sdcca}})$ a t-round (non-tag based) self-destruct CCA string commitment scheme, that makes use of the following tools.

1. $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}} = (\mathsf{Com}^{\mathtt{bit}}_{\mathsf{sdcca}}, \mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdcca}})$ is a t-round (non-tag based) self-destruct CCA bit commitment scheme.
2. $\Pi_{\mathsf{NMCode}} = (\mathsf{Enc}, \mathsf{Dec})$ is a continuos non-malleable code resilient against $\mathsf{PermOver}$. The procedure Enc outputs a codeword that is n-bits long.

Informal description of Π_{sdcca}. At a very high level our self-destruct CCA string commitment scheme works as follow. The sender $\mathsf{Sen}_{\mathsf{sdcca}}$ encodes the message m to commit using Enc and obtains enc^m. Then he interacts with $\mathsf{Rec}_{\mathsf{sdcca}}$ in order to commit to every single bit of $\mathsf{enc}^m = \mathsf{enc}^m_1 || \ldots || \mathsf{enc}^m_n$ using $\mathsf{Com}^{\mathtt{bit}}_{\mathsf{sdcca}}$.

[10] Imagine that adversary touches only C_1, C_2 puts first half of bits from L_1, L_2 into L'_1 and second half into L'_2 while $R'_1 = R_1, R'_2 = R_2$. Then if first block C'_1 fulfill the parity requirements then C'_2 will also have correct parity

In the opening phase $\mathsf{Sen}_{\mathsf{sdcca}}$ sends the opening of the bit commitments along with m'. $\mathsf{Rec}_{\mathsf{sdcca}}$ accepts the commitment iff $m = m'$, where m is the output of the decoding function Dec on input the codeword obtained from the opening of bit commitment.

This constructions preserves the number of rounds of the bit commitment $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}}$ and does not add any other computational assumptions.

Our SD-CCA commitment scheme is described in more details in Fig 1.

Common input: security parameter λ.
Input to $\mathsf{Sen}_{\mathsf{sdcca}}$: $m \in \{0,1\}^{\lambda}$.
Commitment phase:

1. $\mathsf{Sen}_{\mathsf{sdcca}}$:
   1. Run $\mathsf{enc}^m \leftarrow \mathsf{Enc}(1^{\lambda}, m)$.
2. $\mathsf{Sen}_{\mathsf{sdcca}} \leftrightarrow \mathsf{Rec}_{\mathsf{sdcca}}$:
   1. For $i = 1, \ldots, n$, $\mathsf{Sen}_{\mathsf{sdcca}}$ on input enc^m_i interacts with $\mathsf{Rec}_{\mathsf{sdcca}}$ and computes the commitment phase τ_i of $\mathsf{Com}^{\mathtt{bit}}_{\mathsf{sdcca}}$ obtaining the i-th decommitment information $\mathtt{dec}^{\mathsf{sdcca}}_i$ [a].
3. $\mathsf{Rec}_{\mathsf{sdcca}}$: accepted the commitment iff the following conditions are satisfied.
   1. For $i = 1, \ldots, n$ τ_i is accepting.
   2. For $i, j = 1, \ldots, n$ and $i \neq j$ τ_i is not identical to τ_j.

Decomittment phase:

1. $\mathsf{Sen}_{\mathsf{sdcca}} \rightarrow \mathsf{Rec}_{\mathsf{sdcca}}$
   1. Send $(m, \mathtt{dec}^{\mathsf{sdcca}}_1, \ldots, \mathtt{dec}^{\mathsf{sdcca}}_n)$ to $\mathsf{Rec}_{\mathsf{sdcca}}$.
2. $\mathsf{Rec}_{\mathsf{sdcca}}$: For $i = 1, \ldots, n$ if $\mathtt{dec}^{\mathsf{sdcca}}_i$ is not a valid decommitment of τ_i then abort, otherwise continue as follow.
   1. Run $(\mathsf{enc}^{m'}_i) \leftarrow \mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdcca}}(\tau_i, \mathtt{dec}^{\mathsf{sdcca}}_i)$.
   2. Set $\mathsf{enc}^{m'} = \mathsf{enc}^{m'}_1 || \ldots || \mathsf{enc}^{m'}_n$ and run $m' \leftarrow \mathsf{Dec}(\mathsf{enc}^m)$.
   3. If $m \neq m'$ abort, otherwise output 1.

[a] The n commitment phases are computed in parallel.

Fig. 1: Description of our SD-CCA string commitment scheme.

Theorem 2. *If $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}} = (\mathsf{Com}^{\mathtt{bit}}_{\mathsf{sdcca}}, \mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdcca}})$ is a t-round (non-tag based) self-destruct CCA bit commitment scheme and $\Pi_{\mathsf{NMCode}} = (\mathsf{Enc}, \mathsf{Dec})$ is a continuous non-malleable code resilient against $\mathsf{PermOver}$, then $\Pi_{\mathsf{sdcca}} = (\mathsf{Sen}_{\mathsf{sdcca}}, \mathsf{Rec}_{\mathsf{sdcca}})$ is a a t-round (non-tag based) self-destruct CCA string commitment scheme.*

Proof. **Correctness.** The correctness follows from the correctness of $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}}$ and the definition of code satisfied by Π_{NMCode}.

Statistically Binding. The statistical binding property follows from the statistical binding of $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}}$ and the definition of code satisfied by Π_{NMCode}.

Self-destruct CCA. In order to prove that Π_{sdpcca} is a SD-CCA string commitment scheme we have to show that the distribution of the output of $\mathcal{A}$ in the experiment IND_0 is computational indistinguishable from the distribution of the output of $\mathcal{A}$ in the experiment IND_1 (see Def. 3).

We proceed through a series of hybrid experiments where we will demonstrate that the view of $\mathcal{A}$ combined with the committed values that he queries in the experiment IND_0 along with his output in IND_0 (that we denote with $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_0}(1^\lambda, z)\}_{\lambda\in\mathbb{N},z\in\{0,1\}^*}$) is computational indistinguishable from $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_1}(1^\lambda, z)\}_{\lambda\in\mathbb{N},z\in\{0,1\}^*}$.

We considering the following hybrid experiments.

- IND_b is the real game experiment as defined in the Definition 3, where in the left sessions $\mathsf{Sen}_{\mathsf{sdcca}}$ is committing to enc^{m_b} that is the codeword obtained running $\mathsf{Enc}(1^\lambda, m_b)$, and the adversary $\mathcal{A}$ interacts with the oracle in $\mathtt{poly}(\lambda)$ concurrent right sessions.
- $\mathcal{H}^i_b(1^\lambda, z)$. Let enc^{m_b} be the codeword obtained running $\mathsf{Enc}(1^\lambda, m_b)$.

  In the *left session* interact with $\mathcal{A}$ as the $\mathsf{Sen}_{\mathsf{sdcca}}$ does except that the first i commitments $\tau_1, \ldots, \tau_i$ of the first i bits of enc^{m_b} are replaced by a commitments of a random and independent bits $r_1, \ldots, r_i$.

  In the *right sessions* act as $\mathsf{Rec}_{\mathsf{sdcca}}$ does. Furthermore when a commitment query $\tilde{\tau}_1, \ldots, \tilde{\tau}_n$ in k-th right sessions ($k = 1, \ldots, \mathtt{poly}(\lambda)$) is completed outputs a tampering k-th functions f^b_k defined as follow.

  1. If τ_i is identical to $\tilde{\tau}_j$ set the i-th position of f^b_k to j.
  2. Otherwise forward $\tilde{\tau}_i$ to $\mathcal{O}^{\mathtt{bit}}$ and obtains the committed bit $\hat{b}$. Set the i-th bit of f^b_k to $\mathtt{set}$ if $\hat{b} = 0$ and to $\mathtt{reset}$ otherwise. If $\mathcal{O}^{\mathtt{bit}}$ implements the self-destruct mode write $\perp$ in all the positions of f^b_k.

  After that each $\tilde{\tau}_i$ is processed, the hybrid experiment use the function f^b_k to answer the k-th query made by $\mathcal{A}$. In more details, the hybrid runs $\tilde{m} \leftarrow \mathsf{Dec}(f^b_k(\mathsf{enc}_{m_b}))$, and based on the value of $\tilde{m}$ he takes one of the following choice: a) if $\tilde{m} = m_b$ he outputs $\mathtt{same}$; b) if $\tilde{m} = \perp$ he implements the self-destruct mode; c) if $\tilde{m} \notin \{m_b, \perp\}$ he forwards $\tilde{m}$ to $\mathcal{A}$.

Let $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^i_b}(1^\lambda, z)\}_{\lambda\in\mathbb{N},z\in\{0,1\}^*}$[11] be the random variable describing the view of the adversary $\mathcal{A}$ combined with the committed values that $\mathcal{A}$ queries in the $\mathtt{poly}(\lambda)$ right sessions along with the output of $\mathcal{A}$ in hybrid experiment $\mathcal{H}^i_b(1^\lambda, z)$, $i = 0, \ldots, n$ $b \in \{0, 1\}$.

We are now ready to argue that, due to the self-destruct CCA of $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}}$, $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^{i-1}_b}(1^\lambda, z)\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^i_b}(1^\lambda, z)\}$ for $i = 1, \ldots, n$. In more details we are going to prove the following lemma.

Lemma 12. *For all $m_b \in \{0,1\}^\lambda$ for all $i = 1, \ldots, n$ and $b \in \{0, 1\}$ it holds that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^{i-1}_b(1^\lambda,z)}\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^i_b}(1^\lambda, z)\}$.*

Suppose by contradiction that the lemma does not hold for some i, then it is possible to show an adversary $\mathcal{A}^{\mathcal{O}^{\mathtt{bit}}}_{\mathsf{sdcca}}$ that breaks the self-destruct CCA of $\Pi^{\mathtt{bit}}_{\mathsf{sdcca}}$ as follows. Let $\mathcal{C}_{\mathsf{sdcca}}$ be the challenger of the bit self-destruct CCA game. $\mathcal{A}^{\mathcal{O}^{\mathtt{bit}}}_{\mathsf{sdcca}}$ interacts with $\mathsf{Sen}_{\mathsf{sdcca}}$ and $\mathsf{Rec}_{\mathsf{sdcca}}$ according to the steps described in $\mathcal{H}^{i-1}_b$ (and in $\mathcal{H}^i_b$) except for the i-th commitment τ_i of the left session. For the commitment τ_i he acts as a proxy between $\mathcal{C}_{\mathsf{sdcca}}$ and $\mathsf{Sen}_{\mathsf{sdcca}}$. In the end, $\mathcal{A}^{\mathcal{O}^{\mathtt{bit}}}_{\mathsf{sdcca}}$ runs a distinguisher $\mathcal{D}$ (that exists by contradiction) on input the view

[11] In order to not overburden the notation in the rest of the proof we will write $\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^i_b}(1^\lambda, z)$.

of $\mathcal{A}$ combined with the committed values that $\mathcal{A}$ queries in the the $\mathtt{poly}(\lambda)$ right sessions and the output of $\mathcal{A}$ in the execution with $\mathcal{A}^{\mathcal{O}^{\mathtt{bit}}}_{\mathsf{sdcca}}$ and outputs what $\mathcal{D}$ outputs. W.l.o.g. we assume that the i-th of enc^{m_b} is 1 and the random chosen bit r_i committend in τ_i in $\mathcal{H}^i_b$ is 0, then if $\mathcal{C}_{\mathsf{sdcca}}$ commits to 1 $\mathcal{A}^{\mathcal{O}^{\mathtt{bit}}}_{\mathsf{sdcca}}$ is acting as in $\mathcal{H}^{i-1}_b$, otherwise he is acting as in $\mathcal{H}^i_b$. This observation conclude the proof. We observe that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^0_b}(1^\lambda, z)\} \equiv \{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_b}(1^\lambda, z)\}$, It follows from the definition of $\mathcal{H}^0_b$. Therefore we can conclude that $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_b}(1^\lambda, z)\} \equiv \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^0_b}(1^\lambda, z)\} \approx \cdots \approx \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^n_b}(1^\lambda, z)\}$ for $b \in \{0,1\}$. It remains to argue that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^n_0}(1^\lambda, z)\} \equiv_s \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^n_1}(1^\lambda, z)\}$. For this last part of the proof we will rely on continuos non-malleability of Π_{NMCode}. We notice that both $\mathcal{H}^n_1$ and $\mathcal{H}^n_0$ define the same distribution of tampering functions $F = (F_1, \ldots, F_{\mathtt{poly}(\lambda)})$ and $F \in \mathsf{PermOver}$ since that there are no repeated transcripts. Furthermore, both hybrid experiments $\mathcal{H}^n_1$ and $\mathcal{H}^n_0$ answer to a query of $\mathcal{A}$ with $\mathtt{same}$ when the tampering function does not change the codeword (the codeword can correspond to different messages). We also notice that if the tampering function outputs $\perp$ then both the hybrids implement the self-destruct mode. We can conclude that for all $m_0, m_1 \in \{0,1\}^\lambda$ $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^n_0}(1^\lambda, z)\} \equiv_s \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^n_1}(1^\lambda, z)\}$, which implies that for all $m_0, m_1 \in \{0,1\}^\lambda$ it holds that $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_0}(1^\lambda, z)\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_1}(1^\lambda, z)\}$.

6 Parallel CCA Commitment Scheme from SD-PCCA Commitment Scheme

In this section we describe our $\Pi_{\mathsf{pcca}} = (\mathsf{Sen}_{\mathsf{pcca}}, \mathsf{Rec}_{\mathsf{pcca}})$ a $t+2$-round (non-tag based) PCCA string commitment scheme, that makes use of the following tools.

1. $\Pi_{\mathsf{sdpcca}} = (\mathsf{Sen}_{\mathsf{sdcca}}, \mathsf{Rec}_{\mathsf{sdcca}})$ is a t-round (non-tag based) SD-PCCA string commitment scheme.
2. a 2-round IDTC scheme $\Pi = (\mathsf{Sen}, \mathsf{Rec}, \mathsf{TFake})$ for the following $\mathcal{NP}$-language $L = \{\tau_{\mathsf{sdcca}} : (m, \mathtt{dec}_{\mathsf{sdcca}})$ s.t. $\mathsf{Rec}_{\mathsf{sdcca}}$ on input $(m, \mathtt{dec}_{\mathsf{sdcca}})$ accepts m as a decommitment of $\tau_{\mathsf{sdcca}}\}$.
3. $\Pi_{\mathsf{sign}} = (\mathsf{Gen}, \mathsf{Sign}, \mathsf{Verify})$ is a signature scheme.
4. A 4-round delayed-input public coin $\mathsf{LS}_{\mathsf{trap}} = (\mathcal{P}_{\mathsf{trap}}, \mathcal{V}_{\mathsf{trap}})$ with SHVZK simulator S_{trap}. $\mathsf{LS}_{\mathsf{trap}} = (\mathcal{P}_{\mathsf{trap}}, \mathcal{V}_{\mathsf{trap}})$ is adaptive-input PoK for the $\mathcal{NP}$-relation $\mathsf{Rel}_{L_{\mathsf{trap}}}$ where $L_{\mathsf{trap}} = \{(\mathsf{vk} : \exists\ (\sigma_1, \mathsf{msg}_1, \sigma_2, \mathsf{msg}_2)$ s.t. $\mathsf{Verify}(\mathsf{vk}, \mathsf{msg}_1, \sigma_1) = 1$ AND $\mathsf{Verify}(\mathsf{vk}, \mathsf{msg}_2, \sigma_2) = 1$ AND $\mathsf{msg}_1 \neq \mathsf{msg}_2\}$. We denote with ℓ_{trap} the dimension of the instances belonging to $\mathsf{LS}_{\mathsf{trap}}$.

Informal description of our $\Pi_{\mathsf{pcca}} = (\mathsf{Sen}_{\mathsf{pcca}}, \mathsf{Rec}_{\mathsf{pcca}})$. At a very high level our PCCA string commitment scheme works as follow. The sender $\mathsf{Sen}_{\mathsf{pcca}}$ interacts with the receiver $\mathsf{Rec}_{\mathsf{pcca}}$ in order to compute a commitment τ_{sdcca} of m using Π_{sdpcca}. Furthermore, $\mathsf{Rec}_{\mathsf{pcca}}$ engages with $\mathsf{Sen}_{\mathsf{pcca}}$ a protocol to allow the extraction of a trapdoor. We adopt the one described in [COSV17b] in which the trapdoor is represented by the knowledge of two signatures under a verification key sent by $\mathsf{Rec}_{\mathsf{pcca}}$ in the 4th last round. In order to allow the extraction of the

trapdoor, $\mathsf{Rec}_{\mathsf{pcca}}$ sends a signature of a message randomly chosen in the 3rd last round by the $\mathsf{Sen}_{\mathsf{pcca}}$.

Furthermore $\mathsf{Sen}_{\mathsf{pcca}}$ engages $\mathsf{Rec}_{\mathsf{pcca}}$ a special WIPoK that proves that τ_{sdcca} is a well-formed commitment or the knowledge of two signatures for two different messages w.r.t. a verification key.

In more details, the special WIPoK works as follows. In the 3rd last round $\mathsf{Sen}_{\mathsf{pcca}}$, computes and sends a trapdoor commitment $\mathtt{com}$ w.r.t. the instance τ_{sdcca} using the equivocal commitment procedure of Π.

In the 4th last round and in the 2nd last round $\mathsf{Sen}_{\mathsf{pcca}}$ receives, respectively, from $\mathsf{Rec}_{\mathsf{pcca}}$ 1st and the 3rd round $(\mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}})$ of $\mathsf{LS}_{\mathsf{trap}}$. Then, $\mathsf{Sen}_{\mathsf{pcca}}$ in the last round has to complete the $\mathsf{LS}_{\mathsf{trap}}$ transcript. The $\mathsf{LS}_{\mathsf{trap}}$ transcript proves the knowledge of two signatures for two different messages w.r.t. a verification key vk. Since, the honest committer $\mathsf{Sen}_{\mathsf{pcca}}$ has no such knowledge he runs the SHVZK simulator of $\mathsf{LS}_{\mathsf{trap}}$ on input $\mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}, \mathsf{vk}$ in order to obtain the 2nd and the 4th round $(\mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ of $\mathsf{LS}_{\mathsf{trap}}$.

Moreover, in the last round $\mathsf{Sen}_{\mathsf{pcca}}$ runs the equivocal procedure of Π (using as equivocal secret the decommitment informations of τ_{sdcca}) and obtains an opening $(\mathtt{dec}, \mathsf{ls}^2_{\mathsf{trap}})$ w.r.t. of $\mathtt{com}$. $\mathsf{Sen}_{\mathsf{pcca}}$ sends to $\mathsf{Rec}_{\mathsf{pcca}}$ $(\mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}}, \mathtt{dec})$.

In the opening phase $\mathsf{Sen}_{\mathsf{pcca}}$ sends the opening of Π_{sdpcca}.

Our PCCA commitment scheme is described in more details in Fig 2.

Theorem 3. *If $\Pi_{\mathsf{sdpcca}} = (\mathsf{Sen}_{\mathsf{sdpcca}}, \mathsf{Rec}_{\mathsf{sdpcca}})$ is a t-round (non-tag based) self-destruct PCCA string commitment scheme and OWFs exists, then $\Pi^{\mathsf{sdcca}} = (\mathsf{Sen}_{\mathsf{pcca}}, \mathsf{Rec}_{\mathsf{pcca}})$ is a a $t+2$-round (non-tag based) PCCA string commitment scheme.*

Before we start the security proof we recall that $\mathsf{LS}_{\mathsf{trap}}$ (see [COSV17a, LS90]) and Π (see [COSV17b]) as well as Π_{sign} (using [Rom90]) can be instantiated from OWFs.

Proof. **Correctness.** The correctness follows from the correctness of Π_{sdpcca} and the completeness of Π_{OR}. **Statistically Binding.** The statistical binding property follows from the statistical binding of Π_{sdpcca}. **Parallel CCA.** In order to prove that Π_{pcca} is a PCCA string commitment scheme we have to show that the distribution of the output of $\mathcal{A}$ in IND_0 is computational indistinguishable from the distribution of the output of $\mathcal{A}$ in IND_1 (see Def. 5). We proceed through a series of hybrid experiments where we will demonstrate that the view of $\mathcal{A}$ combined with the committed values that he queries in the experiment IND_0 along with the output of $\mathcal{A}$ in IND_0 (that we will denote with $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_0}(1^\lambda, z)\}_{\lambda\in\mathbb{N}, z\in\{0,1\}^*}$) is computational indistinguishable from $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_1}(1^\lambda, z)\}_{\lambda\in\mathbb{N}, z\in\{0,1\}^*}$.

- $\mathcal{H}^0_b(1^\lambda, z)$ is the real game experiment as defined IND_b in the Definition 5, where in the left session $\mathsf{Sen}_{\mathsf{sdcca}}$ is committing to m_b, and the adversary $\mathcal{A}$ interacts with the oracle $\mathcal{O}^{\mathsf{pcca}}$ in $\mathtt{poly}(\lambda)$ parallel right sessions. We can prove that in $\mathcal{H}^0_b(1^\lambda, z)$ the adversary $\mathcal{A}$, in the the $\mathtt{poly}(\lambda)$ right sessions, computes (except with negligible probability) only well-formed commitments. This proof

Common input: security parameter λ, instances length: $\ell, \ell_{\mathsf{trap}}$.
Input to $\mathsf{Sen}_{\mathsf{pcca}}$: $m \in \{0,1\}^\lambda$.
Commitment phase:

1. $\mathsf{Sen}_{\mathsf{pcca}} \leftrightarrow \mathsf{Rec}_{\mathsf{pcca}}$:
   1. $\mathsf{Sen}_{\mathsf{pcca}}$ on input m interacts with $\mathsf{Rec}_{\mathsf{pcca}}$ and compute the commitment phase τ_{sdpcca} of $\mathsf{Sen}_{\mathsf{sdpcca}}$ obtaining the decommitment information $\mathtt{dec}_{\mathsf{sdpcca}}$.
   2. $\mathsf{Sen}_{\mathsf{pcca}}$ interacts with $\mathsf{Rec}_{\mathsf{pcca}}$ in order to prove that he computes a well-formed commitment of m [a]:
      a. $\mathsf{Rec}_{\mathsf{pcca}}$ runs the following algorithms: $(\mathsf{sk}, \mathsf{vk}) \leftarrow \mathsf{Gen}(1^\lambda)$, $\rho \leftarrow \mathsf{Rec}(1^\lambda, \ell)$, $\mathsf{ls}^1_{\mathsf{trap}} \leftarrow \mathcal{V}_{\mathsf{trap}}(1^\lambda, \ell_{\mathsf{trap}})$. Then, he sends $(\mathsf{vk}, \rho, \mathsf{ls}^1_{\mathsf{trap}})$ to $\mathsf{Rec}_{\mathsf{pcca}}$.
      b. $\mathsf{Sen}_{\mathsf{pcca}}$ picks $\mathsf{msg}, \mathsf{ls}^3_{\mathsf{trap}} \leftarrow \{0,1\}^\lambda$, and runs the following algorithms: $\mathsf{ls}^2_{\mathsf{trap}} \leftarrow \mathcal{S}_{\mathsf{trap}}(1^\lambda,\ \ell_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}; r_{\mathsf{trap}})$, $\mathtt{com}, \mathtt{aux} \leftarrow \mathsf{TFake}(1^\lambda, \rho, \tau_{\mathsf{sdpcca}})$. Then, he sends $(\mathtt{com}, \mathsf{msg})$ to $\mathsf{Rec}_{\mathsf{pcca}}$.
      c. $\mathsf{Rec}_{\mathsf{pcca}}$ picks $\mathsf{ls}^3_{\mathsf{trap}} \leftarrow \{0,1\}^\lambda$ and runs $\sigma \leftarrow \mathsf{Sign}(\mathsf{sk}, \mathsf{msg})$. Then, he sends (c, σ) to $\mathsf{Sen}_{\mathsf{pcca}}$.
      d. $\mathsf{Sen}_{\mathsf{pcca}}$ If $\mathsf{Verify}(\mathsf{vk}, \mathsf{msg}, \sigma) \neq 1$ then aborts, otherwise continues as follow. Runs $\mathsf{ls}^4_{\mathsf{trap}} \leftarrow \mathcal{S}_{\mathsf{trap}}(\mathsf{vk}, r_{\mathsf{trap}})$. Furthermore, he sets $x = (\tau_{\mathsf{sdpcca}})$ and $w = (m, \mathtt{dec}_{\mathsf{sdpcca}})$ then runs $\mathtt{dec} \leftarrow \mathsf{TFake}(x, w, \rho, \mathsf{ls}^2_{\mathsf{trap}}, \mathtt{aux})$. Then, he sends $(\mathtt{dec}, \mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ to $\mathsf{Rec}_{\mathsf{pcca}}$.
1. $\mathsf{Rec}_{\mathsf{sdcca}}$: accept the commitment iff the following conditions are satisfied.
   1. $\mathsf{Rec}(\rho, x, \mathtt{com}, \mathsf{ls}^2_{\mathsf{trap}}, \mathtt{dec}) = 1$.
   2. $\mathcal{V}_{\mathsf{trap}}(\mathsf{vk}, \mathsf{ls}^1_{\mathsf{trap}}, \mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^3_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}}) = 1$
   3. If $\mathsf{Rec}_{\mathsf{sdpcca}}$ accepts the commitment τ_{sdpcca}.

Decomittment phase:

1. $\mathsf{Sen}_{\mathsf{pcca}} \rightarrow \mathsf{Rec}_{\mathsf{pcca}}$: send $(m', \mathtt{dec}_{\mathsf{sdcca}})$ to $\mathsf{Rec}_{\mathsf{pcca}}$.
2. $\mathsf{Rec}_{\mathsf{sdcca}}$: If $\mathtt{dec}_{\mathsf{sdpcca}}$ is not a valid decommitment of τ_{sdpcca} then abort, otherwise runs $\mathsf{Rec}_{\mathsf{sdpcca}}$ on input $(\tau_{\mathsf{sdpcca}}, \mathtt{dec}_{\mathsf{sdpcca}})$ obtaining m. If $m = m'$ output 1 and 0 otherwise.

[a] The rounds a. and b. can be run in parallel with the last two rounds of the commitment phase of Π_{sdpcca}.

Fig. 2: Description of our Parallel CCA string commitment scheme.

follows from the security of the signature scheme. In more details, suppose by contradiction that $\mathcal{A}$ is not computing a well-formed commitments in the i-th right sessions, then we can construct an adversary $\mathcal{A}_{\mathsf{sign}}$ that breaks the security of the signature scheme sign, for some $i \in \{1, \ldots, \mathtt{poly}(\lambda)\}$. Let $\tilde{\mathsf{vk}}$ be the challenge verification key. The adversary $\mathcal{A}_{\mathsf{sign}}$ interacts adversary $\mathcal{A}$ in the left session as a honest sender $\mathsf{Sen}_{\mathsf{pcca}}$ does. In the rights sessions he acts as a honest receiver $\mathsf{Rec}_{\mathsf{pcca}}$ does except for a i-th right session, for which he acts in the following way. In the i-th right session $\mathcal{A}$ uses $\tilde{\mathsf{vk}}$ to compute the first round and the oracle $\mathsf{Sign}(\tilde{\mathsf{sk}}, \cdot)$ to compute a signature $\tilde{\sigma}_1$ of a message $\tilde{\mathsf{msg}}_1$ sent by $\mathcal{A}$ in the second round. At the end of the execution $\mathcal{A}_{\mathsf{sign}}$ extracts using the extractor Ext of $\mathsf{LS}_{\mathsf{trap}}$ two signatures for two different signatures w.r.t. the verification key $\tilde{\mathsf{vk}}$. Observe that the extraction succeeds with non-negligible probability, because by contradiction we are assuming that $\mathcal{A}$ computes a non-valid commitment. This implies that in the i-th right session the instance $\tilde{\tau}_{\mathsf{sdpcca},i}$ of the IDTC scheme Π is false ($\tilde{\tau}_{\mathsf{sdpcca},i} \notin L$) therefore Π is statistically binding. This guarantees that $\mathcal{A}$ is using the knowledge of two different signatures to complete the $\mathsf{LS}_{\mathsf{trap}}$ transcript. The proof ends with the observation that $\mathsf{Sign}(\tilde{\mathsf{sk}}, \cdot)$ is called only once.

Note that since the right sessions are parallel the reduction does not need to emulate the oracle. In fact it will use the extraction procedure when all the commitment phases of the right sessions are ended and just before the reduction has to send back the committed values to $\mathcal{A}$. Clearly we have that $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_b(1^\lambda,z)}\} \equiv_s \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^0_b(1^\lambda,z)}\}$.

- $\mathcal{H}^1_b(1^\lambda, z)$ differs from $\mathcal{H}^0_b(1^\lambda, z)$ only in the fact that the oracle $\mathcal{O}^{\mathsf{pcca}}$ in $\mathtt{poly}(\lambda)$ parallel right sessions is emulated in polynomial time. Intuitively, in order to extract the value committed in $\tilde{\tau}_{\mathsf{sdcca},i}$ in the i-th right session, the hybrid needs to obtain from $\mathcal{A}$ two valid openings for the commitment $\tilde{\mathtt{com}}_i$ sent in the 2nd last round by $\mathcal{A}$. Indeed, the Ext of Π (that exists from Special Binding of Π) on input two different openings of $\tilde{\mathtt{com}}_i$ and the instance of the $\mathcal{NP}$-language L associated to the IDTC outputs the witness for the instance. In this case, since the instance of the IDTC scheme is $\tilde{\tau}_{\mathsf{sdcca},i}$, the extracted witness corresponds to the opening informations of $\tilde{\tau}_{\mathsf{sdcca},i}$.

  In order to obtain the opening of $\tilde{\mathtt{com}}_i$ the hybrid apply the following extraction procedure $\mathsf{Ext}_{\mathsf{pcca}}$. The hybrid will act as the receiver $\mathsf{Rec}_{\mathsf{pcca}}$ does until the end of the parallel right sessions. Then the hybrid, in i-th right session, rewinds $\mathcal{A}$ up to the 2nd last round and sends multiple different $\tilde{\mathsf{ls}}^{3'}_{\mathsf{trap},i} \leftarrow \{0,1\}^\lambda$ in order to collect a different opening w.r.t. $\tilde{\mathtt{com}}_i$, for $i = 1, \ldots, n$. Finally, the hybrid extracts the committed values using Ext and sends them to $\mathcal{A}$. Observe that in the left session, during the rewinds, $\mathcal{A}$ can send multiple 2nd to last rounds. In particular, $\mathcal{A}$ can send different 3rd rounds of $\mathsf{LS}_{\mathsf{trap}}$. For each new 3rd round of $\mathsf{LS}_{\mathsf{trap}}$, the hybrid in order to complete the left session runs the SHVZK simulator of $\mathsf{LS}_{\mathsf{trap}}$ and

obtains $\mathsf{ls}^{2'}_{\mathsf{trap}}$, $\mathsf{ls}^{4'}_{\mathsf{trap}}$. Then, the hybrid runs TFake in order to obtain opening $(\mathsf{ls}^{2'}_{\mathsf{trap}}, \mathtt{dec})$ w.r.t. $\mathtt{com}$. The hybrid sends $\mathtt{dec}', \mathsf{ls}^{2'}_{\mathsf{trap}}\ \mathsf{ls}^{4'}_{\mathsf{trap}}$ to $\mathcal{A}$.

It is left to argue that during the rewinds $\mathcal{A}$ sends a different openings of $\tilde{\mathtt{com}}_i$, for $i \in \{1, \ldots, \mathtt{poly}(\lambda)\}$. Note that $\mathcal{H}^1_b(1^\lambda, z)$ is distributed statistical close to $\mathcal{H}^0_b(1^\lambda, z)$ until $\mathcal{A}$ receives the committed values. Therefore we are guaranteed that also in $\mathcal{H}^1_b(1^\lambda, z)$ the adversary $\mathcal{A}$ (except with negligible probability) does not use the knowledge of two signatures to compute the $\mathsf{LS}_{\mathsf{trap}}$ transcript. This implies that $\mathcal{A}$ has to equivocate the trapdoor commitment to complete the i-th right session. Therefore, $\mathcal{A}$ in the rewinds is changing the opening of $\tilde{\mathtt{com}}_i$.

For the above reasons the abort probability of the experiment is increasing only by a negligible amount. Moreover $\mathcal{A}$ is using the equivocation procedure of Π, we can claim that in in all the $\mathtt{poly}(\lambda)$ right sessions, computes (except with negligible probability) well-formed commitments. For the above arguments we conclude that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^0_b(1^\lambda,z)}\} \equiv_s \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^1_b(1^\lambda,z)}\}$.

- $\mathcal{H}^2_b(1^\lambda, z)$ differs from $\mathcal{H}^1_b(1^\lambda, z)$ only in the fact that in the left session of $\mathcal{H}^1_b(1^\lambda, z)$ the adversary $\mathcal{A}$ is rewound from the 2nd last round to the 3rd last round, in order to extract two signatures σ_1, σ_2 for two distinct messages $(\mathsf{msg}_1, \mathsf{msg}_2)$ w.r.t. a verification key vk. Note that after p rewinds the probability of not obtaining a valid new signature is less than $1/2$. Therefore the probability that $\mathcal{A}$ does not give a second valid signature for a randomly chosen message after λ/p rewinds is negligible in λ. For the above reasons the procedure of extraction of signatures for different messages in $\mathcal{H}^1_b(1^\lambda, z)$ succeeds except with negligible probability. Observe that the above deviation increases the abort probability of the experiment only by a negligible amount, therefore $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^2_b(1^\lambda,z)}\} \equiv_s \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^1_b(1^\lambda,z)}\}$. Due to the statistical indistinguishability between this two hybrids we can argue that also in $\mathcal{H}^2_b$ the extraction procedure $\mathsf{Ext}_{\mathsf{pcca}}$ succeeds with non-negligible probability. This implies that $\mathcal{A}$ in all the $\mathtt{poly}(\lambda)$ right sessions, computes (except with negligible probability) well-formed commitments.
- $\mathcal{H}^3_b(1^\lambda, z)$ differs from $\mathcal{H}^2_b(1^\lambda, z)$ in the way the transcript of $\mathsf{LS}_{\mathsf{trap}}$ is computed. More precisely, the prover $\mathcal{P}_{\mathsf{trap}}$ of $\mathsf{LS}_{\mathsf{trap}}$ is used to compute the messages $\mathsf{ls}^2_{\mathsf{trap}}$ and $\mathsf{ls}^4_{\mathsf{trap}}$ instead of using the SHVZK simulator. Observe that the procedure of extraction of the signatures succeeds in $\mathcal{H}^3_b(1^\lambda, z)$ with non-negligible probability because before the last round the distribution of $\mathcal{H}^3_b(1^\lambda, z)$ is statistically close to the one of $\mathcal{H}^2_b(1^\lambda, z)$. Note that also in this hybrid we want emulate the oracle in polynomial time. In this hybrid we slightly change the extraction procedure $\mathsf{Ext}_{\mathsf{pcca}}$. The new procedure $\mathsf{Ext}'_{\mathsf{pcca}}$ is almost the same except that for the following modification. In the rewinding threads for the left session the hybrid upon receiving a 3rd round of $\mathsf{LS}_{\mathsf{trap}}$ uses the honest prover $\mathcal{P}$ of $\mathsf{LS}_{\mathsf{trap}}$ to compute $\mathsf{ls}^{2'}_{\mathsf{trap}}, \mathsf{ls}^{4'}_{\mathsf{trap}}$. Moreover, during the rewinding thread the values $\mathtt{dec}, \mathsf{ls}^{2'}_{\mathsf{trap}}$ stay the same. Indeed upon receiving another 3rd round of $\mathsf{ls}_{\mathsf{trap}}$ the hybrid uses the honest prover $\mathcal{P}$ of $\mathsf{LS}_{\mathsf{trap}}$ to compute $\mathsf{ls}^{4'}_{\mathsf{trap}}$ for the same $\mathsf{ls}^{2'}_{\mathsf{trap}}$.

It follows from the SHVZK of $\mathsf{LS}_{\mathsf{trap}}$ that also in this hybrid $\mathsf{Ext}'_{\mathsf{pcca}}$ succeeds with non-negligible probability. In more details, suppose by contradiction that in i-th right session $\mathsf{Ext}'_{\mathsf{pcca}}$ fails to extract the committed messages with non-negligible probability, then it is possible to show an adversary against the SHVZK of $\mathsf{LS}_{\mathsf{trap}}$. The reduction would work as follows. Let $\mathcal{C}^{SHVZK}$ be the challenger for the SHVZK of $\mathsf{LS}_{\mathsf{trap}}$.

1. The reduction interacts with the adversary in the left and in the i-th right session according to $\mathcal{H}^3_b$ (and $\mathcal{H}^2_b$).
2. The reduction, upon receiving $\mathsf{ls}^3_{\mathsf{trap}}$ in the left session forwards this message to $\mathcal{C}^{SHVZK}$ together with $(x_{\mathsf{trap}} = \mathsf{vk}, w_{\mathsf{trap}} = (\mathsf{msg}_1, \mathsf{msg}_2, \sigma_1, \sigma_2))$.
3. The reduction, upon receiving $(\mathsf{ls}^2_{\mathsf{trap}}, \mathsf{ls}^4_{\mathsf{trap}})$ from $\mathcal{C}^{SHVZK}$, uses them to complete the left execution against $\mathcal{A}$.
4. When $\mathcal{A}$ stops, the reduction rewinds the adversary in the i-th right session applying the extraction procedure $\mathsf{Ext}'_{\mathsf{pcca}}$.
5. If the the extraction procedure $\mathsf{Ext}'_{\mathsf{pcca}}$ fails output a random bit, otherwise output 0 (to claim that the challenger has used the Special HVZK simulator).

Note that when the challenger uses the honest procedure to compute the transcript of $\mathsf{LS}_{\mathsf{trap}}$ the view of $\mathcal{A}$ during the reduction is distributed identical to the view of $\mathcal{A}$ in $\mathcal{H}^3_b$. Therefore the probability that the extractor procedure $\mathsf{Ext}'_{\mathsf{pcca}}$ is successful when the transcript of $\mathsf{LS}_{\mathsf{trap}}$ is computed using the honest prover procedure corresponds (except a negligible difference) to the probability that the $\mathsf{Ext}'_{\mathsf{pcca}}$ succeeds in $\mathcal{H}^3_b$. This observation conclude the proof.

Since $\mathsf{Ext}'_{\mathsf{pcca}}$ succeeds with non negligible probability, we can argue that $\mathcal{A}$, in all the $\mathtt{poly}(\lambda)$ right sessions, computes (except with negligible probability) well-formed commitments.

From the SHVZK of $\mathsf{LS}_{\mathsf{trap}}$ it also follows that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^2_b}(1^\lambda, z)\}$ and $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^3_b}(1^\lambda, z)\}$ are computationally indistinguishable.

- The hybrid $\mathcal{H}^4_b(1^\lambda, z)$ differs from $\mathcal{H}^3_b(1^\lambda, z)$ since the honest sender procedure is used to compute $\mathtt{com}, \mathtt{dec}$ in the left session. More precisely, the hybrid computes $\mathsf{ls}^2_{\mathsf{trap}} \leftarrow \mathcal{P}_{\mathsf{trap}}(1^\lambda, \ell_{\mathsf{trap}}, \mathsf{ls}^1_{\mathsf{trap}})$[12] and $(\mathtt{com}, \mathtt{dec}) \leftarrow \mathsf{Sen}(1^\lambda, \mathsf{ls}^2_{\mathsf{trap}})$.

  Note that also in this hybrid we are emulating the oracle $\mathcal{O}^{\mathsf{pcca}}$ extracting the committed values by using the same procedure as in $\mathcal{H}^3_b(1^\lambda, z)$. We can argue that also in this hybrid the procedure $\mathsf{Ext}'_{\mathsf{pcca}}$ succeeds with non negligible probability, otherwise we can show a reduction to the trapdorness property of Π.

  In more details, suppose by contradiction that in i-th right session $\mathsf{Ext}'_{\mathsf{pcca}}$ fails to extract the committed messages with non-negligible probability, it is possible to show an adversary against the trapdoor property of Π. The

[12] Note that due to the delayed-input property of $\mathsf{LS}_{\mathsf{trap}}$ the statement $x_{\mathsf{trap}} = \mathsf{vk}$ and witness $w_{\mathsf{trap}} = (\mathsf{msg}_1, \mathsf{msg}_2, \sigma_1, \sigma_2)$ are required by $\mathcal{P}_{\mathsf{trap}}$ only to compute $\mathsf{ls}^4_{\mathsf{trap}}$ and are not needed to compute $\mathsf{ls}^2_{\mathsf{trap}}$.

reduction would work as follows. Let $\mathcal{C}^{Trap}$ be the challenger for the trapdorness experiment of Π.

1. The reduction interacts with the adversary in the left and in the i-th right session according to $\mathcal{H}_b^4$ (and $\mathcal{H}_b^3$).
2. The reduction, upon receiving $\mathsf{ls}^1_{\mathsf{trap}}, \rho$ in the left session computes $\mathsf{ls}^2_{\mathsf{trap}} \leftarrow \mathcal{P}_{\mathsf{trap}}(1^\lambda, \ell_{\mathsf{trap}}, \mathsf{ls}^1_{\mathsf{trap}})$ and forwards this message to $\mathcal{C}^{Trap}$ together with $\tau_{\mathsf{sdcca}}, \rho$.
3. The reduction, upon receiving $(\mathtt{com}, \mathtt{dec})$ from $\mathcal{C}^{Trap}$, uses them to complete the left execution against $\mathcal{A}$.
4. When $\mathcal{A}$ stops, the reduction rewinds the adversary in the i-th right session applying the extraction procedure $\mathsf{Ext}'_{\mathsf{pcca}}$.
5. If the the extraction procedure $\mathsf{Ext}'_{\mathsf{pcca}}$ fails output a random bit, otherwise output 0 (to claim that the challenger has used equivocation procedure).

We conclude the proof observing that the view of $\mathcal{A}$ in the reduction is distributed identical to $\mathcal{H}_b^3$ when the challenger used the equivocation procedure to compute $(\mathtt{com}, \mathtt{dec})$ and to $\mathcal{H}_b^4$ otherwise.

Since $\mathsf{Ext}'_{\mathsf{pcca}}$ succeeds with non negligible probability, we can argue that $\mathcal{A}$, in all the $\mathtt{poly}(\lambda)$ right sessions, computes (except with negligible probability) well-formed commitments.

The trapdorness of Π ensures that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_b^4}(1^\lambda, z)\}$ and $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_b^3}(1^\lambda, z)\}$ are computationally indistinguishable.

To conclude the proof, it remains left to argue that $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_0^4}(1^\lambda, z)\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_1^4}(1^\lambda, z)\}$. For this last part of the proof we will rely on the self-destruct parallel CCA of Π_{sdpcca}. Suppose by contradiction that the claim does not hold we can use $\mathcal{A}$ to construct an adversary $\mathcal{A}_{\mathsf{sdpcca}}$ that breaks the self-destruct parallel CCA security of Π_{sdpcca}. Let m_0, m_1 be the challenge messages, then $\mathcal{A}_{\mathsf{sdpcca}}$ works as following against the challenger $\mathcal{C}_{\mathsf{sdpcca}}$. In the left session $\mathcal{A}_{\mathsf{sdpcca}}$ acts as a proxy for all the messages of Π_{sdpcca} between $\mathcal{C}_{\mathsf{sdpcca}}$ and $\mathcal{A}$ and computes the other messages of Π_{sdcca} according to the sender procedure described in $\mathcal{H}_0^4(1^\lambda, z)$ ($\mathcal{H}_1^4(1^\lambda, z)$). In the right sessions $\mathcal{A}_{\mathsf{sdpcca}}$ has to emulate the oracle $\mathcal{O}^{\mathsf{pcca}}$, and he will relies on the oracle $\mathcal{O}^{\mathsf{sdpcca}}$ of Π_{sdpcca}. In more details $\mathcal{A}_{\mathsf{sdpcca}}$ acts as a proxy between the queries made by $\mathcal{A}$ and the oracle $\mathcal{O}^{\mathsf{sdpcca}}$. Note that both in the hybrid experiments $\mathcal{H}_0^4(1^\lambda, z)$ and $\mathcal{H}_1^4(1^\lambda, z)$ $\mathcal{A}$ always (except with negligible probability) queries the oracle $\mathcal{O}^{\mathsf{pcca}}$ on well-formed commitments, which implies that the oracle $\mathcal{O}^{\mathsf{sdpcca}}$ never implements the self-destruct mode. Therefore, $\mathcal{A}_{\mathsf{hiding}}$ can always (except with negligible probability) uses $\mathcal{O}^{\mathsf{sdpcca}}$ to handle the queries made by $\mathcal{A}$. As regarding the other messages of Π_{pcca} he will acts as the hones receiver of $\mathcal{H}_0^4(1^\lambda, z)$ ($\mathcal{H}_1^4(1^\lambda, z)$). At the end $\mathcal{A}_{\mathsf{sdpcca}}$ runs the distinguisher $\mathcal{D}$ (that exists by contradiction) that distinguishes $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_0^4}(1^\lambda, z)\}$ from $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}_1^4}(1^\lambda, z)\}$, and he outputs what $\mathcal{D}$ outputs. Observe that if $\mathcal{C}_{\mathsf{sdpcca}}$ commits to m_0 then $\mathcal{A}_{\mathsf{sdpcca}}$ acts as in $\mathcal{H}_0^4(1^\lambda, z)$ otherwise he acts as in $\mathcal{H}_1^4(1^\lambda, z)$. Observe also that the rewinds made to extract the signatures in the left session do not interfere with the reduction since the parallel commitment queries made

by $\mathcal{A}$ ends in the third last round. We can conclude that for all $m_0, m_1 \in \{0,1\}^\lambda$ $\{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^4_0}\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathcal{H}^4_1}\}$, which implies that for all $m_0, m_1 \in \{0,1\}^\lambda$ it holds that $\{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_0}(1^\lambda, z)\} \approx \{\mathtt{view}^{\mathcal{A}}_{\mathsf{IND}_1}(1^\lambda, z)\}$.

Comparison to [CMTV15]. In [CMTV15] authors show that it is possible to construct a string SD-RCCA PKE scheme $\Pi_{\mathsf{sdrcca}} = (\mathsf{Gen}_{\mathsf{sdrcca}}, \mathsf{Enc}_{\mathsf{sdrcca}}, \mathsf{Dec}_{\mathsf{sdrcca}})$ starting from a continuos non malleable code and a 1-bit SD-RCCA PKE[13] $\Pi^{\mathtt{bit}}_{\mathsf{sdrcca}} = (\mathsf{Gen}^{\mathtt{bit}}_{\mathsf{sdrcca}}, \mathsf{Enc}^{\mathtt{bit}}_{\mathsf{sdrcca}}, \mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdrcca}})$ scheme. At a very high level their construction works as follow.

- The generation algorithm $\mathsf{Gen}_{\mathsf{sdrcca}}$ generates l couple of public key secret key running $\mathsf{Gen}^{\mathtt{bit}}_{\mathsf{sdrcca}}$, where l is the length of the codeword.
- The encryption algorithm $\mathsf{Enc}_{\mathsf{sdrcca}}$ encodes the message m to encrypt using the the non-malleable code. Then, it encrypts the l bits of the codewords using the encryption algorithm $\mathsf{Enc}^{\mathtt{bit}}_{\mathsf{sdrcca}}$ of the 1-bit CCA PKE scheme on the l different secret keys.
- The decryption algorithm $\mathsf{Dec}_{\mathsf{sdrcca}}$ decrypts each bit using $\mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdrcca}}$, obtaining the codeword. The encoded message is obtained running the decode algorithm of the non-malleable code.

Replacing their non-malleable code with our non-malleable it is possible to modify their construction to use just one couple of public key secret key. The proposed SD-RCCA PKE scheme changes the previous scheme in the following three aspects: 1) the generation algorithm runs $\mathsf{Gen}^{\mathtt{bit}}$ once generating a single couple of public key secret key; 2) the encryption/decryption algorithms encrypts/decrypts the bits of the codeword using the same secret key/public key; 3) $\mathsf{Dec}_{\mathsf{sdrcca}}$ will output $\perp$ if for $i, j = 1, \ldots, n$ and $i \neq j$ enc_i is identical to enc_j, where enc_i and enc_j are bit encryption. Finally, we note that we obtain a construction with no a priori bound on the length of the string to be encrypted. In more details, we have the following theorem.

Theorem 4. *If* $\Pi^{\mathtt{bit}}_{\mathsf{sdrcca}} = (\mathsf{Gen}^{\mathtt{bit}}_{\mathsf{sdrcca}}, \mathsf{Enc}^{\mathtt{bit}}_{\mathsf{sdrcca}}, \mathsf{Dec}^{\mathtt{bit}}_{\mathsf{sdrcca}})$ *is a 1-bit SD-RCCA PKE and* $\Pi_{\mathsf{NMCode}} = (\mathsf{Enc}, \mathsf{Dec})$ *is a continuos non-malleable code resilient against* $\mathsf{PermOver}$*, then* $\Pi_{\mathsf{sdrcca}} = (\mathsf{Gen}_{\mathsf{sdrcca}}, \mathsf{Enc}_{\mathsf{sdrcca}}, \mathsf{Dec}_{\mathsf{sdrcca}})$ *is a string SD-RCCA PKE scheme.*

The proof is almost the same as the one of Theorem 2. Intuitively, in the modified construction it is possible to use a single couple of public key secret key because our code is resilient against permutation functions.

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[13] The formal definition can be found in [CMTV15]

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A Definition and Tools

Definition 17 (One-way function (OWF)). *A function $f : \{0,1\}^\star \to \{0,1\}^\star$ is called one way if the following two conditions hold:*

- *there exists a deterministic polynomial-time algorithm that on input y in the domain of f outputs $f(y)$;*
- *for every* PPT *algorithm $\mathcal{A}$ there exists a negligible function ν, such that for every auxiliary input $z \in \{0,1\}^{\mathtt{poly}(\lambda)}$:*

$$\text{Prob}\ [y \leftarrow \{0,1\}^\star : \mathcal{A}(f(y), z) \in f^{-1}(f(y))] < \nu(\lambda).$$

Definition 18 (Following the notation of [CPS13]). *A triple of* PPT *algorithms* $(\mathsf{Gen}, \mathsf{Sign}, \mathsf{Verify})$ *is called a* signature scheme *if it satisfies the following properties.*

Validity: *For every pair $(s, v) \leftarrow \mathsf{Gen}(1^\lambda)$, and every $m \in \{0,1\}^\lambda$, we have that*

$$\mathsf{Verify}(v, m, \mathsf{Sign}(s, m)) = 1.$$

Security: *For every* PPT *$\mathcal{A}$, there exists a negligible function ν, such that for all auxiliary input $z \in \{0,1\}^\star$ it holds that:*

$$\Pr[(s, v) \leftarrow \mathsf{Gen}(1^\lambda); (m, \sigma) \leftarrow \mathcal{A}^{\mathsf{Sign}(s,\cdot)}(z, v) \wedge \mathsf{Verify}(v, m, \sigma) = 1 \wedge m \notin Q] < \nu(\lambda)$$

where Q denotes the set of messages whose signatures were requested by $\mathcal{A}$ to the oracle $\mathsf{Sign}(s, \cdot)$.

Definition 19 (Computational indistinguishability). *Let $X = \{X_\lambda\}_{\lambda\in\mathbb{N}}$ and $Y = \{Y_\lambda\}_{\lambda\in\mathbb{N}}$ be ensembles, where X_λ's and Y_λ's are probability distribution over $\{0,1\}^l$, for same $l = \mathtt{poly}(\lambda)$. We say that $X = \{X_\lambda\}_{\lambda\in\mathbb{N}}$ and $Y = \{Y_\lambda\}_{\lambda\in\mathbb{N}}$ are* computationally indistinguishable*, denoted $X \approx Y$, if for every* PPT *distinguisher $\mathcal{D}$ there exists a negligible function ν such that for sufficiently large $\lambda \in \mathbb{N}$,*

$$\left|\Pr\left[\ t \leftarrow X_\lambda : \mathcal{D}(1^\lambda, t) = 1\ \right] - \Pr\left[\ t \leftarrow Y_\lambda : \mathcal{D}(1^\lambda, t) = 1\ \right]\right| < \nu(\lambda).$$

We note that in the usual case where $|X_\lambda| = \Omega(\lambda)$ and λ can be derived from a sample of X_λ, it is possible to omit the auxiliary input 1^λ. In this paper we also use the definition of *Statistical Indistinguishability*. This definition is the same as Definition 19 with the only difference that the distinguisher $\mathcal{D}$ is unbounded. In this case use $X \equiv_s Y$ to denote that two ensembles are statistically indistinguishable.

Definition 20 (Witness Indistinguishable (WI)). *An argument/proof system $\Pi = (\mathcal{P}, \mathcal{V})$, is* Witness Indistinguishable (WI) *for a relation* Rel *if, for every malicious* PPT *verifier $\mathcal{V}^\star$, there exists a negligible function ν such that for all x, w, w' such that $(x, w) \in \mathsf{Rel}$ and $(x, w') \in \mathsf{Rel}$ it holds that:*

$$\left|\Pr\left[\,\langle \mathcal{P}(w), \mathcal{V}^\star\rangle(x) = 1\,\right] - \Pr\left[\,\langle \mathcal{P}(w'), \mathcal{V}^\star\rangle(x) = 1\,\right]\right| < \nu(|x|).$$

Obviously one can generalize the above definitions of WI to their natural adaptive-input variants, where the adversarial verifier can select the statement and the witnesses adaptively, before the prover plays the last round.

Definition 21 (Proof/argument system). *A pair of* PPT *interactive algorithms $\Pi = (\mathcal{P}, \mathcal{V})$ constitute a* proof system *(resp., an* argument system*) for an $\mathcal{NP}$-language L, if the following conditions hold:*

Completeness: *For every $x \in L$ and w such that $(x, w) \in \mathsf{Rel}_{\mathsf{L}}$, it holds that:*

$$\Pr\left[\,\langle \mathcal{P}(w), \mathcal{V}\rangle(x) = 1\,\right] = 1.$$

Soundness*: For every interactive (resp.,* PPT *interactive) algorithm $\mathcal{P}^\star$, there exists a negligible function ν such that for every $x \notin L$ and every z:*

$$\Pr\left[\,\langle \mathcal{P}^\star(z), \mathcal{V}\rangle(x) = 1\,\right] < \nu(|x|).$$

A proof/argument system $\Pi = (\mathcal{P}, \mathcal{V})$ for an $\mathcal{NP}$-language L, enjoys *delayed-input* completeness if $\mathcal{P}$ needs x and w only to compute the last round and $\mathcal{V}$ needs x only to compute the output. Before that, $\mathcal{P}$ and $\mathcal{V}$ run having as input only the size of x. The notion of delayed-input completeness was defined in [CPS+16]. An interactive protocol $\Pi = (\mathcal{P}, \mathcal{V})$ is *public coin* if, at every round, $\mathcal{V}$ simply tosses a predetermined number of coins (i.e. a random challenge) and sends the outcome to the prover. Moreover we say that the transcript τ of an execution $b = \langle \mathcal{P}(z), \mathcal{V}\rangle(x)$ is *accepting* if $b = 1$.

Definition 22 (Proof of Knowledge [LP11]). *A protocol $\Pi = (\mathcal{P}, \mathcal{V})$ that enjoys completeness is a* proof of knowledge *(PoK) for the relation $\mathsf{Rel}_{\mathsf{L}}$ if there exists a probabilistic expected polynomial-time machine Ext, called the extractor, such that for every algorithm $\mathcal{P}^\star$, there exists a negligible function ν, every statement $x \in \{0,1\}^\lambda$, every randomness $r \in \{0,1\}^\star$ and every auxiliary input $z \in \{0,1\}^\star$,*

$$\Pr\left[\,\langle \mathcal{P}^\star_r(z), \mathcal{V}\rangle(x) = 1\,\right] \leq \Pr\left[\, w \leftarrow \mathsf{Ext}^{\mathcal{P}^\star_r(z)}(x) : (x, w) \in \mathsf{Rel}_{\mathsf{L}}\,\right] + \nu(\lambda).$$

We also say that an argument system Π is a argument of knowledge *(AoK) if the above condition holds w.r.t. any* PPT *$\mathcal{P}^\star$.*

In this paper we also consider the *adaptive-input* PoK/AoK property for all the protocols that enjoy delayed-input completeness. Adaptive-input PoK/AoK ensures that the PoK/AoK property still holds when a malicious prover can choose the statement adaptively at the last round.

A *3-round protocol $\Pi = (\mathcal{P}, \mathcal{V})$* for a relation Rel_L is an interactive protocol played between a prover $\mathcal{P}$ and a verifier $\mathcal{V}$ on common input x and private input

w of $\mathcal{P}$ s.t. $(x, w) \in \mathsf{Rel}_L$. In a 3-round protocol the first message a and the third message z are sent by $\mathcal{P}$ and the second messages c is played by $\mathcal{V}$. At the end of the protocol $\mathcal{V}$ decides to accept or reject based on the data that he has seen, i.e. x, a, c, z.

We usually denote the message c sent by $\mathcal{V}$ as a *challenge*, and as *challenge length* the number of bit of c.

Definition 23 (Σ-Protocol). *A 3-round public-coin protocol $\Pi = (\mathcal{P}, \mathcal{V})$ for a relation Rel_L is a Σ-Protocol if the following properties hold:*

- *Completeness: if $(\mathcal{P}, \mathcal{V})$ follow the protocol on input x and private input w to $\mathcal{P}$ s.t. $(x, w) \in \mathsf{Rel}_L$, $\mathcal{V}$ always accepts.*
- *Special soundness: if there exists a polynomial time algorithm such that, for any pair of accepting transcripts on input x, $(a, c_1, z_1), (a, c_2, z_2)$ where $c_1 \neq c_2$, outputs witness w such that $(x, w) \in \mathsf{Rel}_L$.*
- *Special Honest Verifier Zero-knowledge (Special HVZK): there exists a* PPT *simulator algorithm $\mathcal{S}$ that for any $x \in L$, security parameter λ and any challenge c works as follow: $(a, z) \leftarrow \mathcal{S}(1^\lambda, x, c)$. Furthermore, the distribution of the output of $\mathcal{S}$ is computationally indistinguishable from the distribution of a transcript obtained when $\mathcal{V}$ sends c as challenge and $\mathcal{P}$ runs on common input x and any w such that $(x, w) \in \mathsf{Rel}_L$*[14].

Definition 24 (adaptive-input Special Honest Verifier Zero-knowledge [COSV17a]). *A delayed-input 3-round protocol $\Pi = (\mathcal{P}, \mathcal{V})$ for relation Rel_L enjoys* adaptive-input Special Honest Verifier Zero-knowledge (SHVZK) *if there exists a two phases* PPT *simulator algorithm $\mathcal{S}$ that works as follow:*

1. *$a \leftarrow \mathcal{S}(1^\lambda, c, \ell; \rho)$, where 1^λ is the security parameter, c is the challenge ℓ is the size of the instance to be proved and the randomness ρ;*
2. *$z \leftarrow \mathcal{S}(x, \rho)$*[15]*, where x is the instance to be proved.*

Π is SHVZK if any $x \in L$ and for any $c \in \{0,1\}^\lambda$, the distribution of the transcripts (a, c, z), computed by $\mathcal{S}$, is computationally indistinguishable from the distribution of a transcript obtained when $\mathcal{V}$ sends c as challenge and $\mathcal{P}$ runs on common input x and any w (available only in the third round) such that $(x, w) \in \mathsf{Rel}_L$.

Definition 25 (Commitment Scheme). *Given a security parameter 1^λ, a commitment scheme $\Pi = (\mathsf{Sen}, \mathsf{Rec})$ is a two-phase protocol between two* PPT *interactive algorithms, a sender Sen and a receiver Rec. In the commitment phase Sen on input message m interacts with Rec to produce a commitment $\mathtt{com}$, and the private output $\mathtt{d}$ of Sen.*

In the decommitment phase, Sen sends to Rec a decommitment information $(m, \mathtt{d})$ such that Rec accepts m as the decommitment of $\mathtt{com}$.

[14] Note that we require that the two transcripts are computationally indistinguishable as in [GMY06], instead of following [CDS94] that requires the perfect indistinguishability between the two transcripts.

[15] To not overburden the notation we omit the randomness when we use the adaptive-input Special HVZK simulator

Formally, we say that $\Pi = (\mathsf{Sen}, \mathsf{Rec})$ *is a perfectly binding commitment scheme if the following properties hold:*

Correctness:

- *Commitment phase. Let* $\mathtt{com}$ *be the commitment of the message* m *given as output of an execution of* $\Pi = (\mathsf{Sen}, \mathsf{Rec})$ *where* Sen*runs on input message* m*. Let* $\mathtt{d}$ *be the private output of* Sen *in this phase.*
- *Decommitment phase*[16]*.* Rec *on input* m *and* $\mathtt{d}$ *accepts* m *as decommitment of* $\mathtt{com}$*.*

Statistical (resp. Computational) Hiding: *for any adversary (resp.* PPT *adversary)* $\mathcal{A}$ *and a randomly chosen bit* $b \in \{0,1\}$*, consider the following hiding experiment* $\mathsf{exp}^b_{\mathcal{A},\Pi}(\lambda)$*:*

- *Upon input* 1^λ*, the adversary* $\mathcal{A}$ *outputs a pair of messages* m_0, m_1 *that are of the same length.*
- Sen*on input the message* m_b *interacts with* $\mathcal{A}$ *to produce a commitment of* m_b*.*
- $\mathcal{A}$ *outputs a bit* b' *and this is the output of the experiment.*

For any adversary (resp. PPT *adversary)* $\mathcal{A}$*, there exist a negligible function* ν*, s.t.:*

$$\left|\Pr\left[\ \mathsf{exp}^0_{\mathcal{A},\Pi}(\lambda) = 1\ \right] - \Pr\left[\ \mathsf{exp}^1_{\mathcal{A},\Pi}(\lambda) = 1\ \right]\right| < \nu(\lambda).$$

Statistical (resp. Computational) Binding: *for every commitment* $\mathtt{com}$ *generated during the commitment phase by a possibly malicious unbounded (resp. malicious* PPT*) sender* $\mathsf{Sen}^\star$ *there exists a negligible function* ν *such that* $\mathsf{Sen}^\star$*, with probability at most* $\nu(\lambda)$*, outputs two decommitments* $(m_0, \mathtt{d}_0)$ *and* $(m_1, \mathtt{d}_1)$*, with* $m_0 \neq m_1$*, such that* Rec *accepts both decommitments.*

We also say that a commitment scheme is perfectly binding *iff* $\nu(\lambda) = 0$*.*

A.1 2-Round Instance-Dependent Trapdoor Commitments

Here we define a special commitment scheme based on an $\mathcal{NP}$-language L where sender and receiver also receive as input an instance x. While correctness and computational hiding hold for any x, we require that statistical binding holds for $x \notin L$ and knowledge of a witness for $x \in L$ allows to equivocate. Finally, we require that a commitment along with two different openings allows to compute the witness for $x \in L$. We recall that $\hat{L}$ denotes the language that includes L and all well formed instances that are not in L.

Definition 26. *Let* 1^λ *be the security parameter,* L *be an* $\mathcal{NP}$*-language and* Rel_L *be the corresponding* $\mathcal{NP}$*-relation. A triple of* PPT *algorithms* $\Pi = (\mathsf{Sen}, \mathsf{Rec}, \mathsf{Sen})$ *is a 2-Round Instance-Dependent Trapdoor Commitment scheme if the following properties hold.*

[16] In this paper we consider a non-interactive decommitment phase only.

Correctness. *In the 1st round,* Rec *on input* 1^λ *and* $x \in \hat{L}$ *outputs* ρ*. In the 2nd round* Sen *on input the message* m*,* 1^λ*,* ρ *and* $x \in L$ *outputs* $(\mathtt{com}, \mathtt{dec})$*. We will refer to the pair* $(\rho, \mathtt{com})$ *as the commitment of* m*. Moreover we will refer to the execution of the above two rounds including the exchange of the corresponding two messages as the commitment phase. Then* Rec *on input* m*,* x*,* $\mathtt{com}$*,* $\mathtt{dec}$ *and the private coins used to generate* ρ *in the commitment phase outputs 1. We will refer to the execution of this last round including the exchange of* $\mathtt{dec}$ *as the decommitment phase. Notice that an adversarial sender* $\mathsf{Sen}^\star$ *could deviate from the behavior of* Sen *when computing and sending* $\mathtt{com}$ *and* $\mathtt{dec}$ *for an instance* $x \in \hat{L}$*. As a consequence* Rec *could output 0 in the decommitment phase. We will say that* $\mathtt{dec}$ *is a valid decommitment of* $(\rho, \mathtt{com})$ *to* m *for an instance* $x \in \hat{L}$*, if* Rec *outputs 1.*

Hiding. *Given a* PPT *adversary* $\mathcal{A}$*, consider the following hiding experiment* $\mathsf{ExpHiding}^b_{\mathcal{A},\Pi}(\lambda, x)$ *for* $b = 0, 1$ *and* $x \in \hat{L}_R$*:*

- *On input* 1^λ *and* x*,* $\mathcal{A}$ *outputs a message* m*, along with* ρ*.*
- *The challenger on input* x, m, ρ, b *works as follows: if* $b = 0$ *then it runs* Sen *on input* m*,* x *and* ρ*, obtaining a pair* $(\mathtt{com}, \mathtt{dec})$*, otherwise it runs* TFake *on input* x *and* ρ*, obtaining a pair* $(\mathtt{com}, \mathtt{aux})$*. The challenger outputs* $\mathtt{com}$*.*
- $\mathcal{A}$ *on input* $\mathtt{com}$ *outputs a bit* b' *and this is the output of the experiment.*

We say that hiding *holds if for any* PPT *adversary* $\mathcal{A}$ *there exist a negligible function* ν*, s.t.:*

$$\left|\text{Prob}\left[\ \mathsf{ExpHiding}^0_{\mathcal{A},\Pi}(\lambda, x) = 1\ \right] - \text{Prob}\left[\ \mathsf{ExpHiding}^1_{\mathcal{A},\Pi}(\lambda, x) = 1\ \right]\right| < \nu(\lambda).$$

Special Binding. *There exists a* PPT *algorithm* Ext *that on input a commitment* $(\rho, \mathtt{com})$*, the private coins used by* Rec *to compute* ρ*, and two valid decommitments* $(\mathtt{dec}, \mathtt{dec}')$ *of* $(\rho, \mathtt{com})$ *to two different messages* m *and* m' *w.r.t. an instance* $x \in L$*, outputs* w *s.t.* $(x, w) \in \mathsf{Rel}_\mathsf{L}$ *with overwhelming probability.*

Trapdoorness. *For any* PPT *adversary* $\mathcal{A}$ *there exist a negligible function* ν*, s.t. for all* $x \in L$ *it holds that:*

$$\left|\text{Prob}\left[\ \mathsf{ExpCom}_{\mathcal{A},\Pi}(\lambda, x) = 1\ \right] - \text{Prob}\left[\ \mathsf{ExpTrapdoor}_{\mathcal{A},\Pi}(\lambda, x) = 1\ \right]\right| < \nu(\lambda)$$

where $\mathsf{ExpCom}_{\mathcal{A},\Pi}(\lambda, x)$ *and* $\mathsf{ExpTrapdoor}_{\mathcal{A},\Pi}(\lambda, x)$ *are defined below*[17]*.*

[17] We assume w.l.o.g. that $\mathcal{A}$ is stateful.

$\mathsf{ExpCom}_{\mathcal{A},\Pi}(\lambda, x)$:	$\mathsf{ExpTrapdoor}_{\mathcal{A},\Pi}(\lambda, x)$:
-*On input* 1^λ *and* x, $\mathcal{A}$ *outputs* (ρ, m).	-*On input* 1^λ *and* x, $\mathcal{A}$ *outputs* (ρ, m).
-Sen*on input* 1^λ, x, m *and* ρ, *outputs* $(\mathtt{com}, \mathtt{dec})$.	-TFake *on input* 1^λ, x *and* ρ, *outputs* $(\mathtt{com}, \mathtt{aux})$.
	-TFake *on input* $\mathtt{tk}$ *s.t.* $(x, \mathtt{tk}) \in \mathsf{Rel}_{\mathsf{L}}$, x, ρ, $\mathtt{com}$, $\mathtt{aux}$ *and* m *outputs* $\mathtt{dec}$.
-$\mathcal{A}$ *on input* $(\mathtt{com}, \mathtt{dec})$ *outputs a bit* b *and this is the output of the experiment.*	-$\mathcal{A}$ *on input* $(\mathtt{com}, \mathtt{dec})$ *outputs a bit* b *and this is the output of the experiment.*

B Instantiation of Definition 11

In this section we discuss possible instantiations of definition 11:

with [AKO17]: $\mathsf{Enc}_{\mathsf{AKO}} : \{0,1\}^m \to \left(\{0,1\}^{O(m^6)}\right)^2$ is $2^{-O(m)}-$ admissible,

with [Li17]: $\mathsf{Enc}_{\mathsf{Li}} : \{0,1\}^m \to \left(\{0,1\}^{O(m \cdot \log m)}\right)^2$ is $2^{-O(m)}-$ admissible.

Of course the second code of the above gives better parameters. However we will argue for both above statements.

The most straight forward instantiation would be using the super-strong NMC from [AKO17] (with the improved affine evasive function from [Agg15]). Let us briefly recall the construction:

For any $m \in \mathcal{M}$, $\mathsf{Enc}_{\mathsf{AKO}}(m) = \mathsf{Enc}_1 \circ \mathsf{Enc}_2(m)$, where for any $m \in \mathcal{M}$, $\mathsf{Enc}_2(m) \leftarrow \{X \in \mathbb{F} \mid h(X) = a||b||m\}$, where h is affine-evasive function, and a, b were chosen uniformly at random. For any $x \in \mathbb{F}$, $\mathsf{Enc}_1(x) = (L, R)$, where $L, R \in \mathbb{F}^N$ are uniform such that $\langle L, R\rangle = x$, $\phi(L, a) =$ valid and $\phi(R, b) =$ valid where ϕ is a check function based on Reed-Salomon codes, such that for every a we get $\Pr_x(\phi(x, a) = \text{valid}) = \frac{1}{|a|}$. The construction is secure for any $N \geq C \cdot \log^4 |\mathbb{F}|$, where C is some constant, which makes the whole codeword length equal to $C \cdot \log^5(|\mathbb{F}|)$.

Another instantiation option is with nm-Extractor (f.e. [Li17]), we need to modify the extractor in generic way:

Definition 27 (Encoding scheme based on nm-Ext). *Let* $r \in N$ *be any constant*

$\mathsf{Enc}_{\mathsf{Li}}(M)$:	$\mathsf{Dec}_{\mathsf{Li}}(L, R)$:
$L, R \leftarrow \{l, r \mid \mathsf{Ext}(l, r) = 0^r||M\}$	*Check whether:* $\mathsf{Dec}(L, R) = 0^r||M$,
Output (L, R)	*If the check fails output* $\perp$
	Otherwise, output $M \in \mathcal{M}$

When we encode m bits with modified [Li17] we obtain $O(m \log m)$ bits of codeword.

Let us make some observations about $(\mathsf{Enc}_{\mathsf{AKO}}, \mathsf{Dec}_{\mathsf{AKO}})$ and $(\mathsf{Enc}_{\mathsf{Li}}, \mathsf{Dec}_{\mathsf{Li}})$ schemes, Since the proofs are identical in all but one cases we will refer to $(\mathsf{Enc}, \mathsf{Dec})$ as any of the mentioned schemes. Also for simplicity we will not get into exact parameters instead of we will denote negligible factors as $\mathsf{negligible}$, reader can think of $\mathsf{negligible} = 2^{-O(m)}$.

[Canonical encoder] Fulfilled trivially.

[Balanced code] There is the same number of codewords decoding to each message. Thus distribution of $\mathsf{Enc}(U)$ is the same as $(U, U' | \mathsf{Dec}(U, U') \neq \perp)$ for U, U' uniform and independent.

[Detection of close to bijective tampering] Since Enc is Super-Strong NMC that means that whenever adversary's tampering outputs valid codeword i.e. $\mathsf{Dec}(f(L), g(R)) \neq \perp$ and $(f(L), g(R)) \neq (L, R)$ then adversary will learn whole $f(L), g(R)$, that can not give him any information about encoded message m else adversary broke the NMC-game, thus:

$$d(Dec(m) | f(L), g(R)) < \mathsf{negligible},$$

by lemma 3 :

$$\Delta\left([f(L), g(R) | \mathsf{Dec}(L, R) = m] \; ; \; [f(\tilde{L}), g(\tilde{R})]\right) < \mathsf{negligible},$$

where $\tilde{L}, \tilde{R}$ are uniform and independent. If $f(\tilde{L}), g(\tilde{R})$ have a lot of entropy then:

$$d(\langle f(\tilde{L}), g(\tilde{R}) \rangle) < \mathsf{negligible},$$

and thus

$$\Pr[\mathsf{Dec}(f(\tilde{L}), g(\tilde{R})) = \perp] > 1 - \mathsf{negligible},$$

$$\Delta\left([\mathsf{Dec}(f(L), g(R)) | \mathsf{Dec}(L, R) = m] \; ; \; [\perp]\right) < \mathsf{negligible}$$

[Leakage resilient storage] Since all of the constructions are based on the randomness extractor by the identical argument as above we will obtain the leakage resilience.

Now let us make some observations about $(\mathsf{Enc}^{par}, \mathsf{Dec}^{par})$ (recall Definition 8). Let us go through all conditions from Definition 11:

[Canonical encoder] Fulfilled trivially.

[Detection of close to bijective tampering] Notice that for uniformly random L, R, parities are respectively $0, 1$ with probability $1/4$. If $(\mathsf{Enc}^{par}, \mathsf{Dec}^{par})$ would not fulfill detection property with some non-negligible probability δ then $(\mathsf{Enc}, \mathsf{Dec})$ would not fulfill it with probability at least $\delta/4$ which is a contradiction.

[Leakage resilient storage] For $\mathsf{Enc}(m) = L, R$ we know that

$$f(L), g(R) \approx_{\mathsf{negligible}} f(U), g(U') \tag{1}$$

where U, U' are independent, uniformly distributed.
Let $\mathsf{Enc}^{par}(m) = (L_0, R_1)$. We want to prove that

$$f(L_0), g(R_1) \approx_{\text{negligible}} f(U_0), g(U_1)$$

where U_0, U_1 are independent, uniformly distributed with parities respectively $0, 1$. Let us define following function:

$$f'(x) := \begin{cases} f(x) \text{ if parity of } x \text{ is } 0 \\ c, \text{ otherwise.} \end{cases}$$

Since f', g' reveal only one extra bit of information compare to f, g, by equation 1 we get

$$f'(L), g'(R) \approx_{\text{negligible}} f'(U), g'(U')$$

which implies the Leakage resilience of $(\mathsf{Enc}^{par}, \mathsf{Dec}^{par})$.

[Detection of complete overwrite of one part, for [AKO17]] If $c = 0^k$ we can easly exclude it from the domain simply assuming $\mathsf{Dec}_{\mathsf{AKO}}(0^k) = \perp$, it is only one point thus such change will affect parameters in negligible way. For $c \neq 0^k$ by lemma 2 we get that for U_0 uniform with parity 0: $\mathbf{H}_\infty(\langle U_0, c\rangle)$ misses only 1 bit to full entropy. Thus $\langle U_0, c\rangle$ will not hit correct affine evasive set, and decoding will result in $\perp$.

[Detection of complete overwrite of one part, for [Li17]] Argument is very similar, since most outer layer of [Li17] is based on inner product, as long as $c \neq 0^k$ we get that $\mathbf{H}_\infty(\langle U_0, c\rangle)$ misses only 1 bit to full entropy. By the 0^r padding on message we added to nm-extractor we get that probability of hitting the $0^r||m$ format is at most $(\frac{1}{2})^{r-1}$ which concludes the proof.

Remark 4. Keep in mind that entropy rate threshold $1/3$ in definition 11 is completely arbitrary and all proofs can be easily adopted to any other threshold.

C Instantiation of a secret sharing scheme

In this section we aim for a coding scheme $RS_N : \{0,1\}^{2n} \to (\{0,1\}^{k_2})^N$ that holds the $\lfloor N/3 \rfloor$-out-of-N secret sharing property. We show such construction for all parameters such that $\lfloor N/3 \rfloor \cdot k_2 \geq 2n$.

It turns out that the only we need for this purpose is the Reed-Solomon error correcting code c with following parameters:

- alphabet size $= 2^{k_2}$,
- block length $= N$,
- message length $M = 2 \cdot \lceil \frac{2n}{k_2} \rceil$.

Now our coding scheme may be defined as: $RS_N(m) = c(m||x)$, where x is a randomness of the same size as m.

We omit the simple proof that the above code actually holds the $\lfloor N/3 \rfloor$-out-of-N secret sharing property.

D Proof of Lemma 8

Proof. Let us denote $h(X,Y) = X', Y'$. Also let us proceed with following definitions, let $\mathsf{Overwrites}_X$ denote number of overwritten, by function h, bits in X. Let $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{Y}'}$ denote the vector of bits that were permuted from X into Y' (if a bit was permuted and then overwritten such bit does not count), and let $\mathsf{Swaps} = |\mathsf{Transfer}_{\mathsf{X}\to\mathsf{Y}'}| + |\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}|$.

Case 1: $\mathsf{Swaps} \neq 0, 2k_1$.
We know that $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{Y}'}| + |\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}| > 0$, without loss of generality let us assume that $|\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}| > 0$, and thus X' has some entropy. Trivially from bounded size of X' we get $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}| + |\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}| \leq n$.

Subcase 1.1: $\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'} = n$*:*
Then $\mathcal{D}$ will output $\perp$ since the parity of X' will be equal 1 instead of 0.

Subcase 1.2: $0 < \mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'} < n$*:*
Then by Definition 11.4, we get

$$\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}, \mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'} \approx_\epsilon \mathsf{Transfer}_{\mathsf{U}_0\to\mathsf{X}'}, \mathsf{Transfer}_{\mathsf{U}_1\to\mathsf{X}'},$$

where U_0, U_1 are independent uniformly distributed over $\{0,1\}^{k_1}$, such that parity of U_i is equal i.
Then by lemma 7 we get that parity of $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'} || \mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}$ is equal 0 with probability $\frac{1}{2} \pm \epsilon_c$. Thus parity of X' will be equal 0 with probability at most $\frac{1}{2} + \epsilon_c$, if parity of X' is equal 1 then $\mathcal{D}$ will output $\perp$.

Case 2: $\mathsf{Swaps} = 2n$.
Then parity of X' is 1 and parity of Y' is 0, $\mathcal{D}$ will output $\perp$.

Case 3: $\mathsf{Swaps} = 0$.
We know that $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{Y}'}| = |\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{X}'}| = 0$, we will denote $h(X,Y) = f(X), g(Y)$. We partition this case in following subcases:

Subcase 3.1: $\mathsf{Overwrites}_X \leq 1/3 \cdot k_1$ *and* $\mathsf{Overwrites}_Y \leq 1/3 \cdot k_1$.
From this subcase assumptions we get

$$\mathbf{H}_\infty(f(X)), \mathbf{H}_\infty(g(Y)) \geq 2/3 \cdot k_1.$$

If $h \not\equiv \mathrm{id}$ then at least one f, g is not id, without loss of generality let us assume that $f \not\equiv \mathrm{id}$. If $f \not\equiv \mathrm{id}$ and $f \in \mathsf{PermOver}$ then there is at least $1/2$ probability [18]

[18] If $f \not\equiv \mathrm{id}$ and $f \in \mathsf{PermOver}$ then there must be at least one overwrite or at least one permutation cycle in f. If there is an overwrite then probability that overwritten bit was equal original is $\frac{1}{2}$. If there was a permutation then it had to have at least one cycle of length at least 2, then $f(X) = X$ only if all elements in a cycle are equal, which again happens with probability at most $\frac{1}{2}$.

that $f(X) \neq X$. Let $Z_X = \{x \in \{0,1\}^{k_1} \mid f(x) = x\}$. Since $\Pr(X \in Z_X) \geq 1/2$ by lemma 4 we get

$$\mathbf{H}_\infty(f(X)|X \in Z_X) \geq \mathbf{H}_\infty(f(X)) - \log \frac{1}{\Pr(X \in Z_X)} \geq \mathbf{H}_\infty(f(X)) - 1$$

Thus by Definition 11.2 we get $\mathcal{D}(f(X), g(Y)) = \perp$ with probability at least $1/2$.

Subcase 3.2: $\mathsf{Overwrites}_X > 1/3 \cdot k_1$ *or* $\mathsf{Overwrites}_Y > 1/3 \cdot k_1$.
Then by Def 11.4

$$\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}, \mathsf{Transfer}_{\mathsf{Y}\to\mathsf{Y}'} \approx_\epsilon \mathsf{Transfer}_{\mathsf{U}_0\to\mathsf{X}'}, \mathsf{Transfer}_{\mathsf{U}_1\to\mathsf{Y}'},$$

where U_0, U_1 are independent uniformly distributed over $\{0,1\}^{k_1}$, such that parity of U_i is equal i.
Let us consider following cases:

Subcase 3.2.a: $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}| = |\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{Y}'}| = 0$.
That means both X' and Y' are constant (i.e. completely overwritten), it contradicts assumption that $h \not\equiv$ const..

Subcase 3.2.b: $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}| = k$ *and* $|\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{Y}'}| = 0$.
By Definition 11.3 we get that $\mathcal{D}(X', Y') = \perp$ with probability at least $1 - \epsilon_c$.
The same holds if $|\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}| = 0$ and $|\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{Y}'}| = k_1$.

Subcase 3.2.c: $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'} \neq 0, k_1$ *or* $\mathsf{Transfer}_{\mathsf{Y}\to\mathsf{Y}'} \neq 0, k_1$.
Without loss of generality assume that $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'} \neq 0, k_1$.

If $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'} \leq 2/3k_1$ then by Definition 11.4 we get

$$\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}, Y \approx_\epsilon \mathsf{Transfer}_{\mathsf{U}_0\to\mathsf{X}'}, U_1$$

By lemma 7 we get that parity of $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}||Y$ is equal 0 with probability $\frac{1}{2} \pm \epsilon_c$. Again parity requirements of $\mathcal{D}$ will be fulfilled with probability at most $\frac{1}{2} \pm \epsilon_c$.

If $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'} > 2/3k_1$ then let us consider complement of $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}$, i.e. the bits of X that were not permuted into X', let us denote that complement as T_X. In a identical fashion as above we obtain

$$T_X, Y \approx_{\epsilon_c} T_{U_0}, U_1$$

We obtain that parity of $T_X||Y$ is unknown and equal to parity of $\mathsf{Transfer}_{\mathsf{X}\to\mathsf{X}'}||Y$.
Again parity requirements of $\mathcal{D}$ will be fulfilled with probability at most $\frac{1}{2} \pm \epsilon_c$.