Universally Composable Security Analysis of TLS— Secure Sessions with Handshake and Record Layer Protocols

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Abstract

We present a security analysis of the complete TLS protocol in the model of Universal Composability. This analysis evaluates the composition of key exchange functionalities to emulate secure communication sessions and is based on the adaption of the secure channel model from [Canetti and Krawczyk, Crypto 2002] to the setting where peer identities are not necessarily known at the start and may remain undisclosed at the end of the protocol. Our analysis shows that TLS including the Diffie-Hellman and key transport suites in the uni-directional and bi-directional model of authentication securely emulates secure communication sessions.

Keywords: Universal Composability, TLS/SSL, key exchange, secure sessions

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1 Introduction

The protocol framework of *Transport Layer Security (TLS)*[1] [18] serves as fundamental primitive for WWW security and has fostered to the most valuable cryptographic protocol family in practice. The TLS protocol suites enable applications to communicate across a distributed network in a way that endpoint authentication and transmission privacy is guaranteed in order for preventing eavesdropping, tampering, and message forgery. Prominent examples include tunneling to create a virtual private network, protect Internet phone calls, and secure the rich facets of multi-party Internet applications, such as Internet voting, online banking, electronic commerce or federated identity management.

The main goal of this paper is to provide a rigorous and generic analysis of TLS's cryptographically relevant parts of the protocol framework, namely the handshake and record-layer protocols. Given the wide deployment of the TLS protocol and the fact that it has been designed as contemporary cryptography started to explore provable security, it is natural that this analysis work is of high, practical interest. The more basic importance of the analytical work is in contributing to a further development of a theory that supports the analysis of complex protocol frameworks as required in real-world applications. Because TLS has already been investigated with respect to certain cryptographic primitives and protocol abstractions (see Section 2), a general belief is that the framework offers a suite of secure sessions protocols. This is even strengthened by the fact that related protocols have been studied (e.g., STS, IPSec). However, it is a well-known fact that rigorous security proofs under computational assumptions give the desired strong guarantee for the security, but neither partial primitives nor protocol derivates do. Yet, there is no security proof of the entire TLS protocol in a solid framework and a careful observation of TLS's subtleties in the various modes provided by the different cipher suites.

Our analysis is carried out in the Universal Composability (UC) security model [5]. The model guarantees protocol security under general composition with arbitrary other protocols and in presence of probabilistic polynomial time-bounded adversaries. Universal Composability has led to a valuable security property and stimulated the search for universal protocol design techniques and their realizations [8, 10, 27, 30, 14, 15]. On the other hand, there are important impossibility results [8, 31] so that a security proof of TLS in the framework is neither obvious nor trivial. Our work particularly continues the way of Canetti's and Krawczyk's consideration of the Σ-protocol underlying the signature based modes in IPSec [12] and their model to build up secure channels [13] in the UC framework with the exception that we explore Universal Composition from a slightly different perspective. Instead of proving single modes, we utilize UC as technique to prove complete protocol frameworks secure in a single proof. Applied to the analysis of TLS, it includes Diffie-Hellman and encrypted key transport in the uni- or bi-directional model of authentication, and their emulation to build secure communication protocols. This sketch clearly poses significant challenges in terms of modeling and analysis.

The most relevant questions posed are how do we reduce the analytical complexity. Is it possible to unitize the framework in meaningful protocol fragments such that the composition theorem allows for an efficient protocol reformulation in the hybrid model? That means, can we define ideal functionalities that capture the cryptographic task of some framework fragment and simply

[1] TLS is the official name for the more recent protocol versions in the SSL/TLS family. There are slight differences between SSL and TLS, but the framework remains substantially the same. We use the name TLS as an umbrella term, since the differences do not apply to our analysis.

reuse the ideal functionality with the next fragment? Otherwise, a composite analysis would not make sense so that we could switch to stand-alone protocol proofs under potentially less stronger security notions. Fortunately, we answer the questions in the positive. To this end, we introduce two UC primitives, dubbed the universal key exchange and universal secure communication sessions functionality. The functionalities are "universal" in the sense that they emulate different key establishment methods and modes of authentication in a self-contained definition. In contrast with the formulation in the post-specified setting as used for the analysis of the Σ-protocol in [12], where the peer identities are disclosed during the protocol execution, in the responder authenticated setting the server identity is publicly known at the start. However, the client identity may remain undisclosed at the end of the protocol. The latter reflects the anonymous uni-directional model of authentication which is of prime interest for anonymous user authentication [41]. In which case, the client does not implicitly authenticate to the server, but executes a higher-layer protocol to send some password over the anonymous channel (e.g., HTTP). These constructions constitute the layered Internet approach for designing network security protocols. However, the protocols significantly differ from universal composable password-based authenticated channels [10, 21]. Here, the adversary can attack the composed protocol. In fact, many of such constructions turned out to be insecure [39, 22, 32]. We show that the TLS framework including the different modes securely emulates the universal secure sessions functionality in presence of non-adaptive adversaries. This result has broadly based applicability. Now we are able to considerably simplify the analysis of higher-layer protocols by employing the composition theorem. None of the prior work has evaluated the essential composability property of TLS before.

Organization The remainder sections are structured as follows. Section 2 summarizes related work on the analysis of TLS. Section 3 clarifies notation and cryptographic building blocks. Section 4 recalls the framework of Universal Composability. Section 5 shortly introduces the TLS protocol family and describes the compositional proof idea. Section 6 is devoted to the subroutines we use throughout the analysis. Section 7 proves the full framework before Section 8 concludes. Appendix A shows that the native composition of handshake protocols with the record-layer is not UC-secure. Appendix B summarizes the TLS specification.

2 Related Work

Because of its eminent role the TLS framework has been repeatedly peer-reviewed. The first analysis dates back to the work of Schneier and Wagner [40]. These authors provide an informal analysis in the core specification. Bleichenbacher [4] has found some weaknesses in the PKCS#1 standard for RSA encryption as used with some SSL 3.0 handshake protocols.[2] Jonsson and Kaliski [29] show that the encryption in the revised PKCS#1.5 standard for RSA encryption is secure against chosen cipher attacks in the Random Oracle Model. Krawczyk [33] analyzes the composition of symmetric authentication and encryption to establish a secure communication channel with TLS record layer protocols. The author mentions some problems in the case of general composition. However, these do not apply to the standard cipher suites.

Apart from the analysis of some cryptographic primitives, a line of research addresses the analysis of *dedicated* TLS protocols on the basis of cryptographic abstractions to allow automated

[2]Note that the attack exploited weaknesses of the PKCS#1 standard and not the TLS protocol.

proof techniques. Paulson [38] conduct an inductive analysis of a simplified version of TLS, using the theorem proving tool Isabelle. Mitchell, Shmatikov, and Stern [35] model checked TLS, using the finite-state enumeration tool named Murphϕ. Ogata and Futatsugi [37] use the interactive theorem prover OTS/CafeObj to check a simplified version of the key transport handshake protocol by means of equational reasoning. He *et al.* [26] provide a correctness proof of TLS in conjunction with the IEEE 802.11i wireless networking protocol, using the Protocol Composition Logic. The drawback these tool-supported approaches currently share is that proofs are considerably simplified. They follow the *Dolev-Yao model* [19] which represents cryptography as term algebras and abstracts away the comprehensiveness of the adversary. There is no mapping lemma that the proofs hold under computational assumptions and in presence of a much more comprehensive adversary.

We remark that Morrissey *et al.* [36] have very recently analyzed in independent work the modularity of a *TLS-like* handshake protocol in a game-based style. The handshake is not exactly conform with the core TLS specification [18] and considers not all protocol variants. Their work focuses on a generic proof of the iterated session key constructions. Further, the security is proved under a weaker definition of the standard indistinguishability notion for key exchange protocols. By contrast, our work is of independent importance and practical relevance. We investigate TLS's intrinsic compositional property, which is to provide higher-layer protocols with some secure communication functionality. Further, our work addresses the native handshake protocols and additionally the record layer protocols in different authentication models under the stronger security notion of Universal Composability.

3 Preliminaries

3.1 Notations

The protocols run between the players client and server. The players may have different roles. A player may act as *initiator* or *responder*. By I, we denote the role of the initiator and by R the responder. Further, let $\mathtt{A}$ be an algorithm. By $y \leftarrow \mathtt{A}(x)$ we denote that y was obtained by running $\mathtt{A}$ on input x. Let $x|y$ denote the concatenations of two elements x and y. Let $p_i : \mathbb{N} \rightarrow \mathbb{N}$, $i \in \mathbb{Z}$ be a polynomial within the security parameter k. By $x \stackrel{r}{\leftarrow} \{0,1\}^{p_i(k)}$ we denote the random variable by selecting a random element of the set $\{0,1\}^{p_i(k)}$. Further, let $j \in \mathbb{Z}$ be the sequence number of a message. Let $P \in (I,R)$ be a tuple consisting of the players I and R. By $\bar{P}$ we denote the inverse, i.e., (R, I). By $\perp$ we denote an anonymous player, i.e., a party whose identity is not known. We refer to the handshake protocol structure as π and the composition with the record-layer protocols as ρ. Additionally, we use different indexes to capture the modes of authentication in ideal functionalities. We refer to a responder-only authenticated functionality as $\mathcal{F}^1$, i.e., a functionality where the responder authenticates to the initiator, but the initiator's identity remains unknown. Further, we denote an ideal functionality, where both players authenticate by $\mathcal{F}^2$, and a hybrid functionality of $\mathcal{F}^1$ and $\mathcal{F}^2$ by $\mathcal{F}^{(1,2)}$.

3.2 Cryptographic Building Blocks and their Constructions

TLS makes according to specification from [18] use of the following cryptographic primitives:

- An ASYMMETRIC ENCRYPTION SCHEME for transporting the encrypted premaster secret whose encryption operation is denoted by $\mathtt{ENC}_{pk_R}()$ and the corresponding decryption operation

by $\texttt{DEC}_{sk_R}()$. R knows a private key sk_R and the corresponding public key pk_R is signed by a Certification Authority (CA). The specification mandates to employ the RSA-OAEP construction in the key transport mode. It has been proven based on the assumptions of the Random Oracle Model to satisfy indistinguishability under adaptive chosen ciphertext attacks [29].

- A DIGITAL SIGNATURE SCHEME for entity authentication. The signing algorithm is denoted by $\texttt{SIG}_{sk}()$ and the predicate verification algorithm by $\texttt{VER}_{vk}()$. The players own a signing key sk and the respective verification vk is certified by a CA. Instations of the digital signature scheme are the DSA and RSA-PSS signing algorithms. The latter construction has been proven to satisfy weak existential unforgability under chosen message attacks in the Random Oracle Model [28].

- A MESSAGE AUTHENTICATION CODE function and a SYMMETRIC ENCRYPTION SCHEME. The authentication function is denoted by $\texttt{HMAC}_k()$, the symmetric encryption and decryption operations by $\texttt{E}_k()$ and $\texttt{D}_k()$. For the implementation of the encryption scheme, the specification recommends DES and 3DES in different modes and with different key lengths. The construction of symmetric authentication with encryption has been proven to satisfy weak unforgeability under chosen message attacks and indistinguishability under chosen plaintext attacks, respectively [33, 3].

- A PSEUDO-RANDOM FUNCTION for the key derivation and confirmation. We denote the function by $\texttt{PRF}_k()$. It is evaluated with seed k on input string l_i, $i \in [1, 4]$. This string is labeled with different publicly known space delimiters and two independently chosen random values, i.e., the nonces transmitted in the first protocol rounds by the players, or a function thereof. See Appendix B.3. The specification defines a special construction based on HMAC combiners. Recently, the construction has been proved to be a good randomness extractor [20].

4 The Universal Composability Security Framework

We give an overview of the UC security framework, referring the reader to [5] for a comprehensive description.

4.1 System Model

In the UC framework, Interactive Turing Machines (ITM) interact in two worlds. See Fig. 1. The real-world model comprises honest parties and the adversary $\mathcal{A}$. The parties run a protocol π in order to compute a cryptographic task. $\mathcal{A}$ controls the communication and potentially corrupts the parties. The ideal world includes "dummy" parties who interact with an ideal functionality $\mathcal{F}$, running the ideal protocol ϕ. The functionality $\mathcal{F}$ represents a trusted party that carries out the same cryptographic task. It simply obtains the inputs of all players and provides them with the desired outputs. The ideal-world adversary $\mathcal{S}$ (dubbed the *simulator*) is allowed to delay messages. However, is unable to gain knowledge of any inputs/outputs except the functionality $\mathcal{F}$ is willing to grant it. Intuitively, the ideal functionality captures the security requirements of a given cryptographic task we expect from the real-world protocol π and defines the adversarial corruption model we consider in that setting.

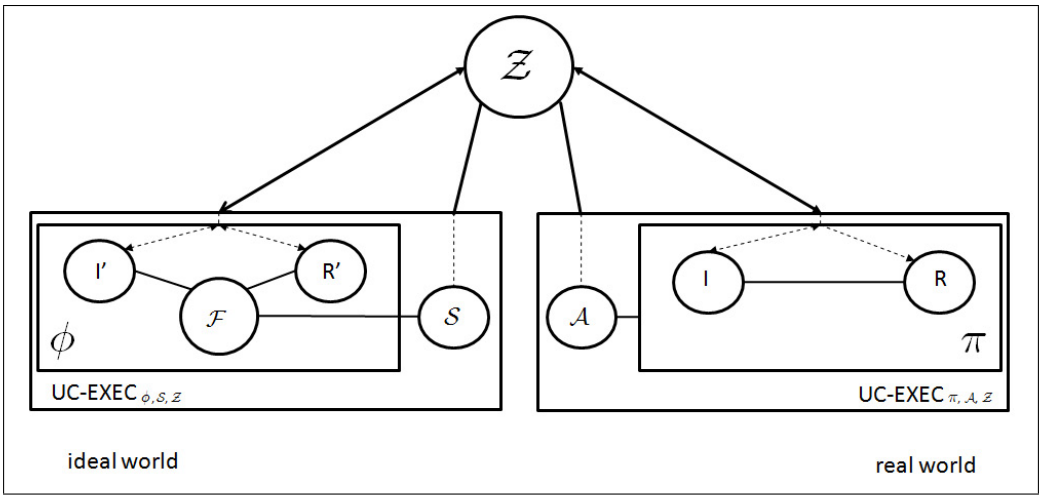

Figure 1: The Real World/Ideal World Paradigm in the UC Framework. In the real world, player I and R execute protocol π in front of adversary $\mathcal{A}$. In the ideal world, the dummy players I' and R' interact with the ideal Functionality $\mathcal{F}$ in presence of the simulator $\mathcal{S}$ to compute the same cryptographic task.

4.2 Security Definition

In the UC framework there exists an additional entity called the environment $\mathcal{Z}$. The environment plays the role of a "judge" who has to distinguish between the two worlds. Therefore, the environment feeds all parties with input, retrieves their outputs, and interacts with the adversary in an arbitrary way throughout the computation. The ideal-world adversary $\mathcal{S}$ does not perceive the message exchange between the real-world parties and has to simulate the interaction in order to mimic the behavior of $\mathcal{A}$. Then, security of protocol π is captured by the fact that every attack $\mathcal{A}$ mounts in the real world, $\mathcal{S}$ carries out in the ideal world. The protocol security is implied, since in the ideal world such attacks cannot be mounted. We have then that the outputs $\mathcal{Z}$ retrieved from the execution of ϕ with the dummy players and $\mathcal{S}$ and the execution of π with the real-world players and $\mathcal{A}$ are indistinguishably distributed. Here, indistinguishability means in this case computational indistinguishability ("$\approx$"). Informally, a protocol π is said to *securely emulate* an ideal-world protocol ϕ. In addition, a protocol π is said to *securely realize* a cryptographic task, if for any real-world adversary $\mathcal{A}$ that interacts with $\mathcal{Z}$ and real players running π, there exists an ideal-world simulator $\mathcal{S}$ that interacts with $\mathcal{Z}$, the ideal functionality $\mathcal{F}$, and the dummy players running the ideal protocol ϕ, so that no probabilistic polynomial time environment $\mathcal{Z}$ is able to distinguish whether it is interacting with the real-world $\mathcal{A}$ or the ideal-world adversary $\mathcal{S}$. A more general definition is:

Definition 1 *A protocol π* `UC-emulates` *protocol ϕ if for any adversary $\mathcal{A}$ there exists an adversary $\mathcal{S}$ such that for all environments $\mathcal{Z}$ that output only one bit:*

$$\texttt{UC}-\texttt{EXEC}_{\phi,\mathcal{S},\mathcal{Z}} \approx \texttt{UC}-\texttt{EXEC}_{\pi,\mathcal{A},\mathcal{Z}}$$

A protocol π `UC-realizes` *an ideal functionality $\mathcal{F}$ if π UC-emulates the ideal protocol for $\mathcal{F}$.*

We sometimes abuse the notation and write $\texttt{UC}-\texttt{EXEC}_{\mathcal{F},\mathcal{S},\mathcal{Z}} \approx \texttt{UC}-\texttt{EXEC}_{\pi,\mathcal{A},\mathcal{Z}}$ to say that π UC-realizes an ideal functionality $\mathcal{F}$. It is easy to see that ideal protocol ϕ is the protocol that

defines the communication between $\mathcal{F}$ and the dummy players that simply forward their inputs and outputs. This is equivalent to $\mathcal{F}$ bypassing the dummy players.

Relaxed UC Security The standard notion of UC security is a strong security definition and rules out the simulatability of some important, provably secure protocols. In order to make the restriction clear, we recall the example from [13]. Consider a two-move Diffie-Hellman protocol. Assume that a prime p and a generator g of a large subgroup of $\mathbb{Z}_p^*$ of prime order are given. The initiator fixes $x \xleftarrow{r} \mathbb{Z}_q$ and sends $\alpha = g^x$ to the responder. Upon reception the responder fixes $y \xleftarrow{r} \mathbb{Z}_q$ and sends $\beta = g^y$. Both players locally output g^{xy}. Simulating the two-move Diffie-Hellman protocol with access to a functionality, say $\mathcal{F}_{\mathrm{KE}}^{\mathrm{bad}}$, which independently fixes the shared key μ at random yields a view that allows the environment to distinguish the two worlds. To understand why, assume that the simulator comes up with the values α' and β'. Next, the environment instructs the adversary to corrupt the initiator before receiving the responder's answer. Then, the environment learns the random value fixed by $\mathcal{F}_{\mathrm{KE}}^{\mathrm{bad}}$ due to the output from the responder and the simulator has to come up with a value x' such that $\beta'^{x'} = \mu$. Since the values α' and β' are independent from μ, a value x' exists only with negligible probability.

To mitigate the limitations, a relaxation of the UC security definition has been proposed in [13] by providing the functionality with some help in form of a non-information oracle $\mathcal{N}$. The oracle outputs a value which is indistinguishable from a random value. More formally, let $\mathcal{N}$ be a polynomial time machine interactive Turing machine. Then $\mathcal{N}$ is a non-information oracle if no interactive Turing machine $\mathcal{M}$, having interacted with $\mathcal{N}$ on security parameter k, can distinguish with non-negligible probability between the local output of $\mathcal{N}$ and a value drawn uniformly from $\{0,1\}^k$. The purpose of the non-information oracle is to supply the simulator with auxiliary information to make the output from the simulation conform to the output from the functionality. The simulator interacts with the non-information oracle and receives the input for the simulation. If the adversary corrupts a player, then $\mathcal{N}$ discloses its current session state to the adversary, including the randomness (x, y) and its local output (g^{xy}).

Definition 2 *A protocol π is said to be* `relaxed UC-secure` *if there exists a non-information oracle $\mathcal{N}$ such that π securely realized $\mathcal{F}^{\mathcal{N}}$.*

In particular, realizing a key exchange functionality under the relaxed definition has been shown to be equivalent to the notion of SK-security. Informally, a key exchange protocol is said to be SK-secure, if (i) no adversary can force the partners of the session to output different session keys, and in addition (ii) guesses whether the output was the real session key or a random test value. The proof is given in [13].

4.3 Universal Composition

A key point of the UC framework is the composition theorem. It guarantees composition with arbitrary sets of parties. Consider a protocol ρ that operates in the $\mathcal{F}$-hybrid model. That is, parties interact in the normal way and in addition can invoke an arbitrary number of copies of the functionality $\mathcal{F}$. We call the invocation of $\mathcal{F}$ subroutine-respecting, if only ρ is permitted to receive the inputs and outputs of the ideal functionality. Then, the following holds.

Theorem 3 *Let π and ϕ be two subroutine-respecting polynomial-time protocols such that π UC-emulates ϕ. Then $\rho^{\pi/\phi}$ UC-emulates ρ for any polynomial-time protocol ρ.*

If π UC-emulates ϕ, we have that there is no $\mathcal{Z}$ that can distinguish with non-negligible probability between the players running π and players running ϕ in the presence of the adversary. The subroutine-respecting invocation ensures that the surrounding protocol ρ feeds π and ϕ in the same way so that the outputs are identical distributed. The composition theorem prevails that replacing the instance of π with an instance of ϕ does not change the behavior of ρ with respect to any polynomial-time adversary; we have a symmetry between the case that ρ interacts with π and ϕ in the presence of the adversary. The main attraction of the composition theorem follows from the fact that if ϕ UC-realizes $\mathcal{F}$ then the real-world protocol ρ can replace the invocation of subroutine π by calling the ideal functionality. The full proof is detailed in [7].

5 Transport Layer Security

5.1 TLS in a Nutshell

The standard TLS protocol framework [18] comprises handshake, alert, change cipher spec, and record-layer subprotocols. The *handshake protocol* is used to negotiate key material and cryptographic algorithms. See Appendix B.2. Once this has finished, the record-layer protocol can secure application data. The *change cipher spec protocol* only consists of one message that triggers a change in the cryptographic parameters used by the record layer, while the *alert protocol* communicates error messages, whenever a failure during the handshake or message protection occurs. Thus, the essential cryptographic building blocks for TLS and target to the presented analysis are the handshake and record-layer protocols.

Handshake and Record Layer The TLS handshake provides a comprehensive framework of messages and cryptographic tools to negotiate a common secret called the *master secret* k_m. It is based on a first secret called the *premaster secret* k_p. The protocol structure depends on the specific cipher suite. That means, the modularity of the handshake protocol is captured by the fact that different subroutines are applied to negotiate the premaster secret while the remaining structure of the handshake is unchanged. See Fig. 2. Among the cipher suites, TLS makes use of different well-known mechanisms, including encryption of the premaster secret using the server's public key (dubbed EKT) and static (dubbed DHS) or ephemeral signed Diffie-Hellman key exchange (dubbed DHE). Optionally, upon request the client in turn authenticates via a signature over all received values `trscrpt`. A verification key is carried within the client certificate. The master secret is then used to derive cryptographic keys for the record layer. The TLS handshake has to provide up to four keys for the record layer, namely a symmetric encryption key (including an IV for the case of block-cipher based encryption), and a MAC key either for the client or the server which vary according to the message flow. By k_a^P we denote the encryption keys and by k_a^P the authentication keys for $P \in \{I, R\}$. Finally, client and server compute finished messages from the master key that confirm the negotiated security parameters. Applying the session keys, the record layer achieves authenticated encryption by including a MAC of the plaintext in the input to symmetric encryption. Note that a message may be fragmented and compressed when it passes the record layer. Therefor, the message fragments are encoded with a sequence number and the record layer maintains a counter in order to prevent that messages are scrambled.

Remark. A key feature of TLS is session resumption in order to reduce server-sided performance penalties. The client names an earlier session that it intends to continue; if the server agrees,

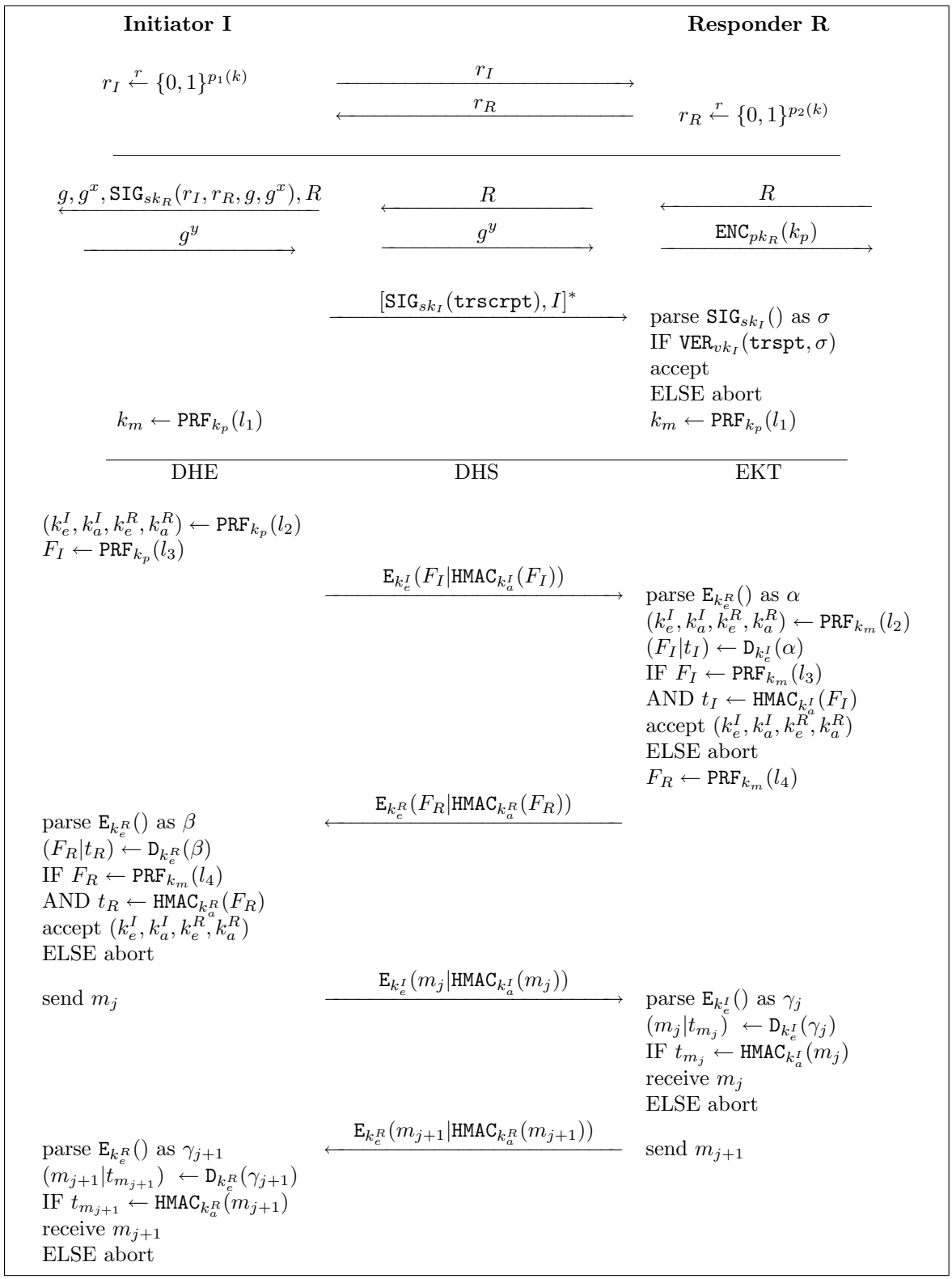

Figure 2: TLS Protocol Specification. Subroutines DHE, DHS, and EKT establish the premaster secret k_p. (*) marks the optional client authentication message. When the protocol outputs an 'abort' message, the alert protocol is invoked with the respective error message. 'send' and 'receive' specify the interfaces to the application layer.

the previous master secret is used with the new nonces to generate new key material for the record layer. We leave out the abbreviated handshake from our analysis in this paper. It is easy to see that the security of the abbreviated handshake follows from the full handshake.

5.2 Proof Idea

The structure of the framework allows conjecture that a composite analysis of TLS is promising. Intuitively, we have the handshake protocols that capture the cryptographic task of key exchange and the composition with record-layer protocols emulates secure transfer of multiple messages. We wish to claim that a modular analysis according to the native TLS composition is feasible, i.e., modeling a handshake protocol as ideal key exchange functionality and composing it with the record-layer protocol in order to instantiate a secure communication session. However, it turns out that the handshake protocol does not securely realize the ideal key exchange functionality. The result applies to all cipher suites in the TLS framework and holds under passive attacks and relaxation of the UC security definition. The reason is that the players commit to the session keys by encrypting and authenticating the finished values. The environment can test the keys on the finished messages and tell the two worlds apart. See Appendix A for a more comprehensive discussion. We avoid the limitation by devising a functionality for the *internal* key exchange within the handshake whose purpose is to negotiate the master secret. It compromises the subroutines DHE, DHS, and EKT. Essentially, they are the heart of the protocol framework and the generic part of TLS. The proof idea is as follows.

We first formulate the requirements of the internal master key exchange schemes by defining an ideal key exchange functionality $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$. It captures the fact that two players receive a random key unless either player is corrupt. Next, we demonstrate that the subroutines DHE, DHS, and EKT securely realize the ideal key exchange functionality $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ (Section 6). We carry out the analysis under the assumption that responder-only and mutual authenticated channels exist. Because TLS operates in a setting where the existence of a trusted third party in the sense of a Certificate Authority (CA) is required, we formalize the global setup assumption by formulating the real-world protocols in $\mathcal{F}$-hybrid models. The $\mathcal{F}$-hybrids are comparable with "long-term authentication modules" that are programmed with some additional cryptographic task. They serve as trusted storage and enforce that access to particular cryptographic services is restricted to privileged players. This restriction qualifies a binding between an "identity" of privileged players and "access" to some deposit, and reflects a proof of possession by demonstrating the ownership of a respective secret. Technically, we utilize the certification functionality $\mathcal{F}_{\mathrm{CERT}}$, certified public key encryption functionality $\mathcal{F}_{\mathrm{CPKE}}$ and the certificate authority functionality $\mathcal{F}_{\mathrm{CA}}$ as presented in [6, 11] to mimic the existence of a CA and certified keys. The *universal composability theorem with joint state (JUC)* [16] guarantees protocol composition where the same certified key is used to encrypt or sign multiple messages by multiple parties to a single recipient, and where the adversary may obtain the output on inputs of its choice.[3] Then, we make use of the composition theorem and specify the TLS protocol in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-hybrid model. We analyze the composite protocol and demonstrate the validity that it securely emulates ideal secure communication sessions (Section 7).

[3] Ralf Küsters pointed out that there are some problems with the joint state theorem. The environment can instruct the dummy adversary to exhaust the ITM hosting the functionality and mount a "denial of service" attack such that it tells the two worlds apart. We remark that the problem affects the underlying system model, but not the essence of the simulation paradigm. A solution to the problem is to replace the Turing machine model by inexhaustible interactive Turing machines [34].

6 The Subroutines

TLS builds on various subroutines for the computation of the master secret. We proceed with the specification of an ideal-world functionality, which we henceforth call universal key exchange $\mathcal{F}_{\text{KE}}^{(1,2)}$ that captures the requirements of the subroutines and then analyze its realization.

6.1 Universal Key Exchange Functionality

The key exchange functionality $\mathcal{F}_{\text{KE}}^{(1,2)}$ is illustrated in Fig. 3. It mimics the cryptographic task that the players I and R agree upon a shared secret μ which is indistinguishable from an independently chosen value of the same length as long as a party is uncorrupted. There is a large body of literature that covers ideal key exchange functionalities (e.g., [5, 13, 11]).

Functionality $\mathcal{F}_{\text{KE}}^{(1,2)}$

$\mathcal{F}_{\text{KE}}^{(1,2)}$ proceeds as follows when parameterized with security parameter k.

- Upon receiving an input ("establish-key", SID, ID_I) from some party, where $ID_I \in (\perp, I)$, record ID_I as initiator, and send a message ("establish-session", SID, ID_I) to the adversary. Upon receiving input ("establish-key", SID, R) from some other party, record R as responder, and send the message ("establish-key", SID, R) to the adversary.
- Upon receiving an answer ("impersonate", SID, $\tilde{\mu}$) from the adversary, do:
  - If ID_I=$\perp$, record the adversary as initiator and send message ("Key", SID, $\perp$, $\tilde{\mu}$) to the responder.
  - Else, ignore the message.
- Upon receiving an answer ("Key", SID, P, $\tilde{\mu}$) from the adversary, where P is either the initiator or responder, do:
  - If neither initiator nor responder is corrupted, and there is no recorded key, fix μ uniformly from $\{0,1\}^k$.
  - If either initiator or responder is corrupted, and there is no recorded key, record $\mu \leftarrow \tilde{\mu}$ as the adversary.

  Send message ("Key", SID, $\bar{P}$, μ) to P.

Figure 3: The Universal Key Exchange Functionality

$\mathcal{F}_{\text{KE}}^{(1,2)}$ is similar to the functionalities with the differences:

- The players authenticate in a post-specified fashion. That means, the environment invokes the players with the session identifier SID and (optionally) their own identity. A party learns the peer identity while executing the TLS protocol. This is captured by the fact that the participants receive the peer identity from the functionality. There is no need for a setup. This is an essential difference of TLS to related protocols (e.g, SSH) where the players have already negotiated their public keys before the protocol start.

- The functionality defines different notions of authenticated key exchange. When the initiator is parameterized with an identity (ID_I=I), the functionality assures a matching conversion between the initiator and server. On the other hand, when the initiator is invoked with an anonymous identity (ID_I=$\perp$), the functionality guarantees a matching conversion with the responder. In which case, the functionality uses `SID` to determine the anonymous player. Such technicality is only feasible for a two party functionality. Recall that the `SID`s of all Turing machines in a protocol instance must be identical in the UC framework. Consequently, any player participating in the same session is a potential initiator, including the adversary mounting an impersonation attack. (The corresponding case in the real world is that the environment instructs the adversary to replay the key exchange protocol with the exception that it chooses its own premaster key. The initiator is unable to terminate the session while the responder accepts the session.)

For clarity, we denote the key exchange functionality where the initiator's identity is kept secret and the initiator's identity disclosed by $\mathcal{F}_{\text{KE}}^{1}$ and $\mathcal{F}_{\text{KE}}^{2}$, respectively. Note that $\mathcal{F}_{\text{KE}}^{(1,2)}$ qualifies a hybrid formulation thereof.

6.2 Master Secret Establishment

We now present the basic subroutines that are used to establish the master secret. DHE and DHS are Diffie-Hellman like key exchange, while EKT is a key transport scheme.

6.2.1 Subroutine DHE

Informally, the subroutine is a 2-way Diffie-Hellman key exchange, whereby exponents are randomly chosen and the responder authenticates via a signature; the verification key is certified by a trusted third party and conveyed with the responder's identity. A description of DHE is given in Fig. 4. We capture the fact that the players own certified keys by exploiting the presence of an ideal certification functionality $\mathcal{F}_{\text{CERT}}$. The functionality only permits the owner of the instance to receive a signature on arbitrary messages while any player can verify the signature. We recall $\mathcal{F}_{\text{CERT}}$ in Fig 5.

The certification functionality $\mathcal{F}_{\text{CERT}}$ is realizable given access to the signature functionality $\mathcal{F}_{\text{SIG}}$ in the presence of a certificate authority $\mathcal{F}_{\text{CA}}$. A certificate authority is used to catch the setup assumption that the parties have registered their identities together with some public value. In addition, $\mathcal{F}_{\text{SIG}}$ adds the cryptographic functionalities for signing messages and verifying their validity. Realizing $\mathcal{F}_{\text{SIG}}$ has been shown to be essentially[4] equivalent to the notion of existential unforgeability under chosen message attacks (EU-CMA). The proof appears in [6]. The JUC-theorem ensures that multiple players have access to the same instance $\mathcal{F}_{\text{CERT}}$ and are able to verify arbitrary messages. Consequently, any CMA-secure signature scheme can be employed for signing, and our proof is therefore independent of the model in which the security of the signature scheme is guaranteed (e.g., in the Random Oracle model in the case of RSA-PSS).

Lemma 4 *Subroutine* DHE *in the $\mathcal{F}_{\text{CERT}}$-hybrid model UC-realizes $\mathcal{F}_{\text{KE}}^{1}$.*

[4]Backes and Hofheinz have pointed out that a corrupted signer can claim any signature for any message. This is due to the fact that $\mathcal{F}_{\text{SIG}}$ allows the simulator to opt for an answer when asked for the verification of some message which has not been signed before. See [2] for a workaround and more discussions.

Subroutine DHE

1. Upon reception of an activation query ("Establish-Key", SID, R), the responder chooses primes $p, q, q/p-1$ and g of order q in $\mathbb{Z}_p^*$. The DH exponent g^x is computed with $x \xleftarrow{r} \mathbb{Z}_q$. It calls $\mathcal{F}_{\text{CERT}}$ with a message ("sign", $\text{SID}_{\text{CERT0}}$, (SID, g, g^x)) where $\text{SID}_{\text{CERT0}}$=($R$, SID $\circ$ 0) contains the identity of the owner of the instance. The responder waits for the delivery of ("Signature", $\text{SID}_{\text{CERT0}}$, (SID, g, g^x), σ). Finally, it computes the response message by placing the signature to the DH parameters (g, g^x, σ) and appends its identity R.

2. When receiving an activation query ("Establish-Key", SID, $\perp$), the initiator waits for the delivery of the response message (g, g^x, σ, R). Then, it calls $\mathcal{F}_{\text{CERT}}$ on query ("verify", $\text{SID}_{\text{CERT0}}$, ($g$, g^x), σ) and waits for the verification ("verified", $\text{SID}_{\text{CERT0}}$, ($g$, g^x), f). If σ is an invalid signature ($f = 0$), it aborts. Otherwise, the initiator sends g^y with $y \xleftarrow{r} \mathbb{Z}_q$. Additionally, it computes the master secret $k_m \leftarrow \text{PRF}_{g^{xy}}(l_1)$ and locally outputs ("Key", SID, R, k_m).

3. Upon delivery of message (g^y), the responder derives the master secret $k_m \leftarrow \text{PRF}_{g^{xy}}(l_1)$ and locally outputs ("Key", SID, $\perp$, k_m).

Figure 4: Subroutine DHE, in the $\mathcal{F}_{\text{CERT}}$-Hybrid model

Proof. We construct a simulator $\mathcal{S}$ such that no environment $\mathcal{Z}$ can tell whether it interacts with subroutine DHE in front of adversary $\mathcal{A}$ and $\mathcal{F}_{\text{KE}}^1$ in presence of $\mathcal{S}$. $\mathcal{S}$ runs a simulated copy of $\mathcal{A}$ and mimics an interaction with players executing DHE. Further, the simulator allows the adversary $\mathcal{A}$ to attack the simulated protocol execution in arbitrary way throughout the simulation. Any input from $\mathcal{Z}$ is forwarded to $\mathcal{A}$ and any output from $\mathcal{A}$ is copied to $\mathcal{S}$ in order to be forwarded to $\mathcal{Z}$. $\mathcal{S}$ tries to make the internal protocol simulation consistent with the real protocol execution under the condition that it has no information about the master key k_m. To this end, the simulator proceeds as follows:

1. **Simulating Activation of** R Upon receiving message ("establish-key", SID, R) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ feeds $\mathcal{A}$ with the message (g, g^x, σ, R) from R. Here, x is chosen at random and σ is obtained by handing $\mathcal{A}$ the message ("sign", $\text{SID}_{\text{CERT0}}$, (SID, g, g^x)) on behalf of $\mathcal{F}_{\text{CERT}}$, and setting σ to the value returned by $\mathcal{A}$, where $\text{SID}_{\text{CERT0}}$=($R$, SID $\circ$ 0).

2. **Simulating Activation of** I Upon receiving message ("establish-key", SID, $\perp$) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ waits for delivery of message (g, α, σ, P') from $\mathcal{A}$. It verifies that σ is a valid signature by querying $\mathcal{F}_{\text{CERT}}$ on input ("verify", $\text{SID}_{\text{CERT0}}$, ($g$, α), σ). (The verification succeeds, if P=R and σ is the response to a request by $\mathcal{S}$ of the form ("sign", $\text{SID}_{\text{CERT0}}$, (SID, g, α)) sent to $\mathcal{F}_{\text{CERT}}$. It fails, if $P' \neq R$ is the identity of an uncorrupted party.) $\mathcal{S}$ mimics the response message by choosing y at random and sending g^y to R. In addition, it chooses the master key k_m to be a random value Δ_{k_m} in the same domain and sends ("Key", SID, $\perp$, k_m) to $\mathcal{F}_{\text{KE}}^1$.

3. **Simulating Reception of the Response Message by** R When $\mathcal{A}$ delivers the message β, then two cases exists.

Functionality $\mathcal{F}_{\text{CERT}}$

$\mathcal{F}_{\text{CERT}}$ proceeds as follows, running on security parameter k, with parties P and an adversary $\mathcal{S}$. The SID is assumed to consist of a pair SID=($\text{PID}_{\text{owner}}$, SID$'$), where $\text{PID}_{\text{owner}}$ is the owner of this instance.

Signature Generation: Upon receiving a value ("sign", SID, m) from some the signer $\text{PID}_{\text{owner}}$, send ("sign", SID, m) to the adversary. Upon receiving ("Signature", SID, m, σ) from the adversary, verify that no entry (m, SID, 0) is recorded. If it is, then output an error message to $\mathcal{S}$ and halt. Else, output ("Signature", SID, m, σ) to $\mathcal{S}$, and record the entry (m, σ, 1).

Signature Verification: Upon receiving a value ("verify", SID, m, σ) from some party P, hand ("verify", SID, m, σ) to the adversary. Upon receiving ("verified", SID, m, φ) from the adversary, do:

1. If (m, σ, 1) is recorded then set f=1.
2. Else, if the signer is not corrupted, and no entry (m, σ', 1) for any σ' is recorded, then set f=0 and record the entry (m, σ, 0).
3. Else, if there is an entry (m, σ, f') recorded, then set f=f'.
4. Else, set f=φ, and record the entry (m, σ, φ).

Output ("verified", SID, m, f) to P.

Figure 5: Certification Functionality

- If $\beta \neq g^y$, then the adversary impersonated the client. $\mathcal{S}$ sets the master secret $k_m \leftarrow \text{PRF}_{\beta^x}(l_1)$ and sends message ("impersonate", SID, k_m) followed by a ("Key", SID, R, k_m) message to $\mathcal{F}_{\text{KE}}^1$.
- If β=g^y, then no impersonation occurred. $\mathcal{S}$ fixes the same master key k_m=Δ_{k_m} and answers $\mathcal{F}_{\text{KE}}^1$ with message ("Key", SID, R, k_m).

4. **Simulating Static Corruption** If one of the parties gets corrupted, then $\mathcal{S}$ proceeds by emulating a DHE protocol session, just as a honest party would play it.

We now prove that the simulator $\mathcal{S}$ is such that no environment $\mathcal{Z}$ can distinguish between the ideal execution of $\mathcal{F}_{\text{KE}}^1$ and $\mathcal{S}$, and the real execution of DHE and $\mathcal{A}$ in the $\mathcal{F}_{\text{CERT}}$-hybrid model. This is done in the following way. We define an intermediate hybrid distributions $\mathcal{H}_\mathbf{1}$ and show that the environment notifies the change by contradicting the decisional Diffie-Hellman assumption. The hybrid distribution $\mathcal{H}_\mathbf{1}$ is defined as follows:

- $\mathcal{H}_\mathbf{1}$ takes the distribution of the output of $\mathcal{Z}$ which is identical to a real execution of $\mathcal{A}$ running DHE in the $\mathcal{F}_{\text{CERT}}$-hybrid model with the following exception. When protocol DHE instructs the initiator (respectively the responder) to evaluate the pseudo-random function PRF() with the premaster secret α^y (resp. β^x) in order to compute the master key k_m, we replace the output with an independently chosen random key from the same domain.

CLAIM 5 *Assume the Decisional Diffie-Hellman assumption holds, then* $\texttt{UC}-\texttt{EXEC}^{\mathcal{F}_{\text{CERT}}}_{\text{DHE},\mathcal{A},\mathcal{Z}} \approx \mathcal{H}_1$.

Assume there is an environment $\mathcal{Z}$ and an adversary $\mathcal{A}$ such that $\mathcal{Z}$ distinguishes with non-negligible probability between $\texttt{UC}-\texttt{EXEC}^{\mathcal{F}_{\text{CERT}}}_{\text{DHE},\mathcal{A},\mathcal{Z}}$ and $\mathcal{H}_1$. We construct an adversary $\mathcal{D}$ that contradicts the Diffie-Hellman assumption. That is, given the tuple g^a, g^b, g^z where $a, b \xleftarrow{r} \mathbb{Z}_q$, $\mathcal{D}$ distinguishes between the case where $z{=}ab$ and $z \xleftarrow{r} \mathbb{Z}_q$.

Parameterized with the tuple g^a, g^b, g^z, $\mathcal{D}$ runs a copy of $\mathcal{Z}$ on a simulated interaction with $\mathcal{A}$ and parties running DHE in the $\mathcal{F}_{\text{CERT}}$-hybrid model. $\mathcal{D}$ plays for $\mathcal{Z}$ the roles of $\mathcal{A}$ and the parties with the following exceptions. When the responder fixes g^x, it replaces the value with g^a. Next, if the initiator fixes g^y, it replaces the value with g^b. Further, whenever the players evaluate the pseudo-random function $\texttt{PRF}()$ on seed α^y (resp. β^x) in order to compute the master key, $\mathcal{D}$ replaces the evaluation with $\texttt{PRF}_{g^z}()$. Finally, whenever $\mathcal{Z}$ instructs $\mathcal{A}$ to corrupt a party, $\mathcal{D}$ outputs whatever the simulated $\mathcal{Z}$ outputs.

Consider the case where $z \xleftarrow{r} \mathbb{Z}_q$. We claim that in this case the view of the simulated $\mathcal{Z}$ is identically distributed to the view of $\mathcal{Z}$ in $\mathcal{H}_1$. This is so because g^x, g^y and the seed for the evaluation of the pseudo-random function $\texttt{PRF}_{g^z}()$ are independently and randomly chosen, assuming that $\mathcal{D}$ did not abort. Now, consider the case where $z{=}ab$. We claim that in this case the view of the simulated $\mathcal{Z}$ is identically distributed to the view of $\mathcal{Z}$ in $\texttt{UC}-\texttt{EXEC}^{\mathcal{F}_{\text{CERT}}}_{\text{DHE},\mathcal{A},\mathcal{Z}}$. The only potential mismatch is that the initiator accepts a value α, which is different from the value g^x the responder has chosen. Such a mismatch cannot occur due to the security properties of $\mathcal{F}_{\text{CERT}}$, i.e., the initiator accepts α only when σ is a valid signature under the condition that $\mathcal{Z}$ did not instruct $\mathcal{A}$ to corrupt the responder. And the only way for this to occur is that the responder has registered as owner of the instance $\texttt{SID}_{\text{CERT0}}$, and queried $\mathcal{F}_{\text{CERT}}$ on message ("sign", $\texttt{SID}_{\text{CERT0}}$, ($\texttt{SID}$, g, g^x)) in order to retrieve a signature.

CLAIM 6 *Assume* $\texttt{PRF}()$ *is a secure pseudo-random function, then* $\mathcal{H}_1 \approx \texttt{UC}-\texttt{EXEC}_{\mathcal{F}^1_{\text{KE}},\mathcal{S},\mathcal{Z}}$.

Assume that such environment $\mathcal{Z}$ that distinguishes with non-negligible simulation slice between the interaction $\mathcal{H}_1$and $\texttt{UC}-\texttt{EXEC}_{\mathcal{F}^1_{\text{KE}},\mathcal{S},\mathcal{Z}}$ exists. We construct an adversary $\mathcal{D}$ who has access to the black-box oracle $\mathcal{O}_{\mathcal{P}}$ that contains throughout the whole simulation either $\texttt{PRF}_{k_p}()$ for some randomly chosen secret k_p or a truly random function with the same range. $\mathcal{D}$ runs a copy of $\mathcal{Z}$ and emulates the interaction with parties running DHE in the $\mathcal{F}_{\text{CERT}}$-hybrid model and $\mathcal{A}$ unless it is required to compute the master secret k_m for which it calls $\mathcal{O}_{\mathcal{P}}$. It is easy to see that $\mathcal{D}$ interpolates between $\mathcal{H}_1$ (this is the case where $\mathcal{O}_{\mathcal{P}}$ contains $\texttt{PRF}_{k_p}()$) and $\texttt{UC}-\texttt{EXEC}_{\mathcal{F}^1_{\text{KE}},\mathcal{S},\mathcal{Z}}$ (this is the case where $\mathcal{O}_{\mathcal{P}}$ contains the truly random function). Hence, the output of $\mathcal{Z}$ can be directly used by $\mathcal{D}$ to decide on the content of $\mathcal{O}_{\mathcal{P}}$.

■

6.2.2 Subroutine DHS

This subroutine is identical to protocol DHE with the exception that the responder's DH exponent is certified by a trusted third party and carried within its identity. See Fig. 6. We capture the slight difference by reformulating DHE in the $\mathcal{F}_{\text{CA}}$-hybrid model. Functionality $\mathcal{F}_{\text{CA}}$ serves as trusted registration authority where the responder escrows a static DH exponent g^x and group g of order

q in $\mathbb{Z}_p^*$. $\mathcal{F}_{\text{CA}}$ outputs the values to arbitrary players when it is invoked with the identity of the registered owner. Essentially, $\mathcal{F}_{\text{CA}}$ binds the deposit to a particular identity and captures the setup that a CA has certified static DH parameters. For self-containment, we recall the ideal certificate authority functionality $\mathcal{F}_{\text{CA}}$ in Fig. 7.

Subroutine DHS

1. Upon reception of an activation query ("Establish-Key", SID, R), the responder sends a response messages consisting of its identity R, signaling that it has registered static DH parameters at $\mathcal{F}_{\text{CA}}$ by sending a query ("register", SID_{CA0}, (g, g^x)) where SID_{CA0}=(R, SID $\circ$ 0) includes its identity.

2. When the initiator receives an activation query ("Establish-Key", SID, $\perp$), it waits for the delivery of the response message (R) and invokes $\mathcal{F}_{\text{CA}}$ with query ("retrieve", SID_{CA0}). Then, the initiator waits for the output ("retrieve", SID_{CA0}, (g, g^x)). If no value is retrievable, it aborts. Otherwise, the initiator sends g^y with $y \stackrel{r}{\leftarrow} \mathbb{Z}_q$ and computes the master secret $k_m \leftarrow \text{PRF}_{g^{xy}}(l_1)$. It terminates the protocol with local output ("Key", SID, R, k_m).

3. Upon delivery of message (g^y), the responder computes the master secret $k_m \leftarrow \text{PRF}_{g^{xy}}(l_1)$ and locally outputs ("Key", SID, $\perp$, k_m).

Figure 6: Subroutine DHS, in the $\mathcal{F}_{\text{CA}}$-Hybrid model

Lemma 7 *Subroutine* DHS *in the $\mathcal{F}_{\text{CA}}$-hybrid model UC-realizes $\mathcal{F}_{\text{KE}}^1$.*

Proof. Again, we construct a simulator $\mathcal{S}$ such that no environment $\mathcal{Z}$ distinguishes between the case that $\mathcal{Z}$ interacts with DHS in the $\mathcal{F}_{\text{CA}}$-hybrid model and $\mathcal{A}$, and the case that $\mathcal{Z}$ interacts with $\mathcal{F}_{\text{KE}}^1$ and $\mathcal{S}$. $\mathcal{S}$ runs a simulated copy of the adversary $\mathcal{A}$ and mimics for $\mathcal{A}$ an interaction with players executing DHE under the limitation that it has no clue about the master key k_m. It allows $\mathcal{A}$ to attack the simulated protocol execution in arbitrary way throughout the simulation. Any input from $\mathcal{Z}$ is forwarded to $\mathcal{A}$ and any output from $\mathcal{A}$ is copied to $\mathcal{S}$ in order to be forwarded to $\mathcal{Z}$. In order to make the environment's view consistent with its view of the real-world protocol execution, $\mathcal{S}$ proceeds as follows:

1. **Simulating Activation of** R Upon receiving message ("establish-key", SID, R) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ feeds $\mathcal{A}$ with the message ("register", SID_{CA0}, (g, g^x)), where SID_{CA0}=(R, SID $\circ$ 0) includes the responder's identity. It simulates that the responder has registered at $\mathcal{F}_{\text{CA}}$. In addition, it feeds $\mathcal{A}$ with message (R) from R.

2. **Simulating Activation of** I Upon receiving message ("establish-key", SID, $\perp$) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ waits for delivery of message (P') from $\mathcal{A}$. It retrieves the registered parameters by emulating a request ("retrieve", SID_{CA0}), where SID_{CA0}=(P', SID $\circ$ 0), to $\mathcal{A}$ on behalf of I. It waits for the answer ("retrieve", SID_{CA0}, (g, g^x)) from $\mathcal{A}$. (The retrieval succeeds, if P'=R and (g, g^x) is the response to a the previous registration request by $\mathcal{S}$ of the form ("register", SID_{CA0}, (g, g^x)) sent to $\mathcal{F}_{\text{CERT}}$. It fails, if $P' \neq R$ is the identity of an uncorrupted party.) In addition, $\mathcal{S}$ simulates the response message by choosing y at random and sending g^y to R. Finally, it

Functionality $\mathcal{F}_{\text{CA}}$

$\mathcal{F}_{\text{CERT}}$ proceeds as follows, running on security parameter k, with parties P and an adversary $\mathcal{S}$. The $\texttt{SID}$ is assumed to consist of a pair $\texttt{SID}=(\texttt{PID}_{\text{owner}}, \texttt{SID}')$, where $\texttt{PID}_{\text{owner}}$ is the owner of this instance.

Registration: Upon receiving the first message ("register", $\texttt{SID}$, m) from party P, send ("register", $\texttt{SID}$, m) to the adversary; upon receiving $\texttt{ok}$ from the adversary, and if this is the first request from P, then record the pair (P, m).

Retrieval: Upon receiving a message ("retrieve", $\texttt{SID}$) from other party P', send ("retrieve", $\texttt{SID}$, P') to the adversary, and wait for an $\texttt{ok}$ from the adversary. Then, if there is a recorded pair ($\texttt{SID}$, m) output ("retrieve", $\texttt{SID}$, m) to P'. Else output ("retrieve", $\texttt{SID}$, $\perp$) to P'.

Figure 7: Certificate Authority Functionality

chooses the master key k_m to be a random Δ_{k_m} in the same range and sends ("Key", $\texttt{SID}$, $\perp$, k_m) to $\mathcal{F}_{\text{KE}}^1$.

3. **Simulating Reception of the Response Message by** R When $\mathcal{A}$ delivers the message β, then two cases exists as in the previous simulation.

   - If $\beta \neq g^y$, then the adversary acts as the initiator. $\mathcal{S}$ reacts in the following way to make the simulation consistent with the real-world. It sets the master secret $k_m \leftarrow \texttt{PRF}_{\beta^x}(l_1)$ and sends message ("impersonate", $\texttt{SID}$, k_m) followed by a ("Key", $\texttt{SID}$, R, k_m) message to $\mathcal{F}_{\text{KE}}^1$.
   - If $\beta = g^y$, then the adversary did not intercept the communication between the initiator and responder. $\mathcal{S}$ fixes the same master key $k_m = \Delta_{k_m}$ as in the previous simulation step and queries $\mathcal{F}_{\text{KE}}^1$ with message ("Key", $\texttt{SID}$, R, k_m).

4. **Simulating Static Corruption** If one of the parties gets corrupted, then $\mathcal{S}$ proceeds by emulating a DHE protocol session, just as a honest party would play it.

Analyzing $\mathcal{S}$, it can be seen that the environment's view when interacting with DHS is computational indistinguishable from its view when interacting with DHE. The only difference is that in DHS the responder calls $\mathcal{F}_{\text{CA}}$ to fix g^x (instead of $\mathcal{F}_{\text{CERT}}$). Thus, the potential mismatch is that the initiator accepts a value α, which is different from the value g^x the responder has fixed. However, such a mismatch cannot occur due to the security properties of $\mathcal{F}_{\text{CA}}$. Hence, by the reduction to the Diffie-Hellman assumption and the security of the pseudo-random function, it can be seen that both views of $\mathcal{Z}$ are computational indistinguishable.

Claim 8 *Assume the Diffie-Hellman assumption holds and* $\texttt{PRF}()$ *is a secure pseudo-random function, then* $\texttt{UC}-\texttt{EXEC}^{\mathcal{F}_{\text{CA}}}_{\text{DHS},\mathcal{A},\mathcal{Z}} \approx \texttt{UC}-\texttt{EXEC}_{\mathcal{F}_{\text{KE}}^1,\mathcal{S},\mathcal{Z}}$.

See the proof of subroutine DHE. ■

6.2.3 **Subroutine** EKT

Informally, this subroutine is different in nature from previous subroutines in that the initiator transports the premaster secret encrypted with the responder's public key. The protocol is illustrated in Fig. 8. We formulate EKT in the $\mathcal{F}_{\text{CPKE}}$-hybrid model that provides the players with certified public key functionality. The functionality $\mathcal{F}_{\text{CPKE}}$ maintains a repository where any invoking player deposits plaintexts and retrieves ciphertexts that are only related to the length of the corresponding plaintexts while the owner of the instance is allowed to access the plaintexts. We recall functionality $\mathcal{F}_{\text{CPKE}}$ in Fig. 9.

Subroutine EKT

1. Upon reception of an activation query ("Establish-Key", $\texttt{SID}$, R), the responder sends the response message (R), indicating that it owns an instance of $\mathcal{F}_{\text{CPKE}}$.

2. Upon reception of an activation query ("Establish-Key", $\texttt{SID}$, $\perp$), the initiator waits for the delivery of the response message (R). It fixes a premaster secret $k_p \stackrel{r}{\leftarrow} \{0,1\}^{p_3(k)}$. Next, the initiator sends the value ("encrypt", $\texttt{SID}_{\text{CPKE0}}$, k_p) to $\mathcal{F}_{\text{CPKE}}$ where $\texttt{SID}_{\text{CPKE0}}=(R, \texttt{SID} \circ \texttt{0})$ contains the owner identity, and waits for the delivery of ("Ciphertext", $\texttt{SID}_{\text{CPKE0}}$, c). Then, the initiator derives master secret $k_m \leftarrow \texttt{PRF}_{k_p}(l_1)$, forwards ciphertext value (c) to the responder, and generates local output ("Key", $\texttt{SID}$, R, k_m).

3. Upon delivery, the responder first decrypts c by sending ("decrypt", $\texttt{SID}_{\text{CPKE0}}$, c) to $\mathcal{F}_{\text{CPKE}}$ and waits for the output ("Plaintext", $\texttt{SID}_{\text{CPKE0}}$, k_p). If the decryption succeeds, the responder computes the master secret $k_m \leftarrow \texttt{PRF}_{k_p}(l_1)$ and terminates with local output ("Key", $\texttt{SID}$, $\perp$, k_m). Otherwise, the responder aborts.

Figure 8: Subroutine EKT, in the $\mathcal{F}_{\text{CPKE}}$-Hybrid model

The ideal certified public key encryption functionality is realizable through a protocol that emulates the ideal public key encryption functionality in the $\mathcal{F}_{\text{CA}}$-hybrid model. The proof is detailed in [11]. Realizing the public key functionality in the presence of non-adaptive[5] adversaries has been proved to be essentially equivalent to the notion of indistinguishability under adaptive chosen ciphertext attacks (CCA2) [14]. Consequently, any CCA2-secure encryption scheme can be employed for key transportation, and our proof is therefore independent of the model in which the security of the encryption scheme is guaranteed (e.g., in the Random Oracle model in the case of RSA-OAEP, which is the recommended in the π specification).

The JUC-theorem guarantees secrecy of encrypted messages in a protocol composition where the same certified public key is used to encrypt multiple messages by multiple parties to a single recipient, and where the adversary may obtain the output of the decryption algorithm on inputs of its choice. It is sufficient to analyze the security of a single instance of $\mathcal{F}_{\text{CPKE}}$ and then apply the JUC theorem to deduce the security of the composite case, where multiple sessions use the same

[5]Obviously, realizing $\mathcal{F}_{\text{CPKE}}$ in the presence of adaptive adversaries so that secrecy of multiple messages is preserved under the condition that the adversary has corrupted the decryptor and gained access to the secret key is a considerably stronger requirement and demands for additional techniques, such as forward secure or non-committing encryption [17, 9].

Functionality $\mathcal{F}_{\text{CPKE}}$

$\mathcal{F}_{\text{CPKE}}$ proceeds as follows, when parameterized by message domain M, a function E with domain M and range $\{0,1\}^*$, and function D of domain $\{0,1\}^*$ and range $M \cup \texttt{error}$. The $\texttt{SID}$ is assumed to consist of a pair $\texttt{SID}=(\texttt{PID}_{\text{owner}}, \texttt{SID}')$, where $\texttt{PID}_{\text{owner}}$ is the owner of this instance.

Encryption: Upon receiving a value ("encrypt", $\texttt{SID}$, m) from a party P, proceed as follows:

1. If $m \notin M$ then return an error message to P.
2. If $m \in M$ then:
   - If the owner of this instance of $\mathcal{F}_{\text{CPKE}}$ is corrupted, then hand also the entire value m to the adversary and receive tag c.
   - Else, calculate a value c by choosing a random r of the same length as m, and selecting $c \leftarrow E(r)$.

   Record the pair (c, m), and output ("Ciphertext", $\texttt{SID}$, c). (If c already appears in a previously recorded pair then return an error message to P.)

Decryption: Upon receiving a value ("decrypt", $\texttt{SID}$, c) from the owner of this instance, proceed as follows. (If the input is received from another party then ignore.)

1. If there is a recorded pair (c, m), then hand ("Plaintext", $\texttt{SID}$, m) to P.
2. Otherwise, compute $m \leftarrow D(c)$, and hand ("Plaintext", $\texttt{SID}$, m) to P.

Figure 9: Certified Public Key Encryption Functionality

instance of $\mathcal{F}_{\text{CPKE}}$.

Lemma 9 *Subroutine* EKT *in the $\mathcal{F}_{\text{CPKE}}$-hybrid model UC-realizes $\mathcal{F}_{\text{KE}}^1$.*

Proof. Let $\mathcal{A}$ be a real-world adversary that operates against EKT. As in the prior simulations, we construct an ideal-world adversary $\mathcal{S}$ such that no environment $\mathcal{Z}$ can distinguish between the case that it interacts with $\mathcal{A}$ and parties running EKT or with $\mathcal{S}$ in the ideal world for $\mathcal{F}_{\text{KE}}^1$. $\mathcal{S}$ runs a simulated copy of $\mathcal{A}$ and mimics an interaction with players executing EKT. It tries to make the internal protocol simulation consistent with the real protocol execution. The simulator allows the adversary $\mathcal{A}$ to attack the simulated protocol execution in arbitrary way throughout the simulation. Any input from $\mathcal{Z}$ is forwarded to $\mathcal{A}$ and any output from $\mathcal{A}$ is copied to $\mathcal{S}$ in order to be forwarded to $\mathcal{Z}$. In detail, $\mathcal{S}$ proceeds in the following way:

1. **Simulating Activation of** R Upon receiving message ("establish-key", $\texttt{SID}$, R) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ feeds $\mathcal{A}$ with the message (R), indicating that the responder has registered at $\mathcal{F}_{\text{CPKE}}$.
2. **Simulating Activation of** I Upon receiving message ("establish-key", $\texttt{SID}$, $\perp$) from $\mathcal{F}_{\text{KE}}^1$, $\mathcal{S}$ waits for delivery of message (P') from $\mathcal{A}$. It chooses a premaster secret k_p at random and queries $\mathcal{A}$ on behalf of $\mathcal{F}_{\text{CPKE}}$ with message ("encrypt", $\texttt{SID}_{\text{CPKE0}}$, k_p), where $\texttt{SID}_{\text{CPKE0}}=(P'$,

$\texttt{SID} \circ 0$). (The request succeeds unless P' is an uncorrupted party and P'=R. Otherwise, the simulator aborts.) Upon reception of message ("Ciphertext", $\texttt{SID}_{\text{CPKE0}}$, c), $\mathcal{S}$ sets the master key k_m to be a value Δ_{k_m} in the same range, forwards the ciphertext c to the responder and outputs ("Key", SID, $\perp$, k_m) to $\mathcal{F}_{\text{KE}}^1$.

3. **Simulating Reception of the Response Message by** R When $\mathcal{A}$ delivers the message α, $\mathcal{S}$ mimics the decryption of α by sending ("decrypt", $\texttt{SID}_{\text{CPKE0}}$, α) to $\mathcal{A}$ on behalf of $\mathcal{F}_{\text{CPKE}}$ and waiting for the response ("Plaintext", $\texttt{SID}_{\text{CPKE0}}$, k'_p). Now, two cases exists agai.

   - If k'_p=k_p, then $\mathcal{S}$ fixes the same master key k_m=Δ_{k_m} as in the previous simulation step and outputs ("Key", SID, $\perp$, k_m) to $\mathcal{F}_{\text{KE}}^1$.
   - Otherwise, $\mathcal{S}$ sets the master secret $k_m \leftarrow \texttt{PRF}_{k'_p}(l_1)$ and outputs message ("impersonate", SID, k_m) followed by a ("Key", SID, R, k_m) message to $\mathcal{F}_{\text{KE}}^1$.

4. **Simulating Static Corruption** If one of the parties gets corrupted, then $\mathcal{S}$ proceeds by emulating a DHE protocol session, just as a honest party would play it.

We now prove that no environment $\mathcal{Z}$ exists that distinguishes with non-negligible simulation slice between the case that $\mathcal{Z}$ interacts with EKT and $\mathcal{A}$, and the case that $\mathcal{Z}$ interacts with $\mathcal{F}_{\text{KE}}^1$ and $\mathcal{S}$. This is done by reduction to the security of the pseudo-random function.

CLAIM 10 *Assume* PRF() *is a secure pseudo-random function, then* $\texttt{UC}{-}\texttt{EXEC}^{\mathcal{F}_{\text{CPKE}}}_{\text{EKT},\mathcal{A},\mathcal{Z}} \approx \texttt{UC}{-}\texttt{EXEC}_{\mathcal{F}_{\text{KE}}^1,\mathcal{S},\mathcal{Z}}$.

Assume that such environment $\mathcal{Z}$ that distinguishes with non-negligible simulation slice between the interaction $\texttt{UC}{-}\texttt{EXEC}^{\mathcal{F}_{\text{CPKE}}}_{\text{EKT},\mathcal{A},\mathcal{Z}}$ and $\texttt{UC}{-}\texttt{EXEC}_{\mathcal{F}_{\text{KE}}^1,\mathcal{S},\mathcal{Z}}$ exists. We construct an adversary $\mathcal{D}$ who has access to the black-box oracle $\mathcal{O}_{\mathcal{P}}$. $\mathcal{D}$ plays for $\mathcal{Z}$ the roles of $\mathcal{A}$ and the parties running EKT. The black-box oracle $\mathcal{O}_{\mathcal{P}}$ contains throughout the whole simulation either $\texttt{PRF}_{k_p}()$ for some randomly chosen secret k_p or a truly random function with the same range. $\mathcal{D}$ runs a copy of $\mathcal{Z}$ and emulates the interaction with parties running EKT in the $\mathcal{F}_{\text{CPKE}}$-hybrid model and $\mathcal{A}$ unless it is required to compute the master secret k_m for which it calls $\mathcal{O}_{\mathcal{P}}$.

Consider the case where $\mathcal{O}_{\mathcal{P}}$ calls the truly random function. We claim that in this case the environment's view is identically distributed to $\texttt{UC}{-}\texttt{EXEC}_{\mathcal{F}_{\text{KE}}^1,\mathcal{S},\mathcal{Z}}$. This is so because the master key is chosen randomly and independently from the initiator's state under the condition that $\mathcal{S}$ did not abort. Now, consider the case where $\mathcal{O}_{\mathcal{P}}$ calls $\texttt{PRF}_{k_p}()$. We claim that in this case the environment's view is identically distributed to $\texttt{UC}{-}\texttt{EXEC}^{\mathcal{F}_{\text{CPKE}}}_{\text{EKT},\mathcal{A},\mathcal{Z}}$. Indeed, the only potential mismatch is that the responder accepts a premaster key k_p that has not been fixed by the initiator, assuming $\mathcal{Z}$ neither instructed $\mathcal{A}$ to corrupt nor impersonate the initiator. Such a mismatch cannot occur due to the security of $\mathcal{F}_{\text{CPKE}}$. The responder accepts the premaster key k_p, when it receives output ("Plaintext", $\texttt{SID}_{\text{CPKE0}}$, k_p) from $\mathcal{F}_{\text{CPKE}}$. And the only way for this to occur is that the initiator has uploaded the premaster key by sending ("encrypt", $\texttt{SID}_{\text{CPKE0}}$, k_p) to $\mathcal{F}_{\text{CPKE}}$ before. ■

6.3 On Realizing Mutual Authentication

The framework enables the responder to opt for initiator authentication. Then, the initiator proves its identity by signing the transcript of incoming and outgoing messages whereby the verification key is certified by a trusted third party and appended to the signature. Employing the composition theorem, we capture the model by reformulating the subroutines in the $\mathcal{F}_{\text{CERT}}$-hybrid model, assuming the registration of the initiator as owner of the instance. The subroutines are extended in the following way:

Before the initiator submits the response message, it stores the message transcript in value `trscrpt`. Next, it feeds $\mathcal{F}_{\text{CERT}}$ with message ("sign", $\texttt{SID}_{\text{CERT1}}$, `trscrpt`) where $\texttt{SID}_{\text{CERT1}}=(I,$ $\texttt{SID} \circ 1)$ includes its identity and waits for the answer ("Signature", $\texttt{SID}_{\text{CERT1}}$, `trscrpt`, ς). Then, the initiator places the signature ς and its identity I to the response message. When the responder receives the message, it first computes its own `trscrpt` and checks that ς is a valid signature by calling $\mathcal{F}_{\text{CERT}}$ on input ("verify", $\texttt{SID}_{\text{CERT1}}$, `trscrpt`, ς). If the verification fails, it aborts. Otherwise, the responder continues to process the subroutine in the normal way. If the subroutine terminates, then the responder generates local output ("Key", `SID`, I, k_m).

Theorem 11 *Subroutines* DHE *in the* $(\mathcal{F}_{\text{CERT}}, \mathcal{F}_{\text{CERT}})$*-hybrid model,* DHS *in the* $(\mathcal{F}_{\text{CA}}, \mathcal{F}_{\text{CERT}})$*-hybrid model and* EKT *in the* $(\mathcal{F}_{\text{CPKE}}, \mathcal{F}_{\text{CERT}})$*-hybrid model UC-realize* $\mathcal{F}_{\text{KE}}^2$.

Proof. The proof follows from the composition theorem. Lemma 4, 7, and 9 imply that the subroutines DHE, DHS and EKT UC-realize $\mathcal{F}_{\text{KE}}^1$. It remains to show that $\mathcal{F}_{\text{KE}}^1$ in the $\mathcal{F}_{\text{CERT}}$-hybrid model securely realizes $\mathcal{F}_{\text{KE}}^2$. It is easy to see that $\mathcal{Z}$'s distribution when it interacts with the dummy players calling functionality $\mathcal{F}_{\text{KE}}^1$ is identical to the its distribution when it communicates with the dummy players invoking $\mathcal{F}_{\text{KE}}^2$ except that the responder's output includes the initiator identity. However, by calling $\mathcal{F}_{\text{CERT}}$ the initiator commits to its identity. Hence, the output distributions of $\mathcal{F}_{\text{KE}}^1$ in the $\mathcal{F}_{\text{CERT}}$-hybrid are indistinguishable from an emulation of $\mathcal{F}_{\text{KE}}^2$. ■

7 TLS UC-Realizes Secure Communication Sessions

The natural abstraction of TLS is to allow players secure exchange of multiple messages in a single instance of the protocol. The sessions are normally invoked by higher-order protocols that enforce different policies to aggregate messages from distinct sources. While key exchange aims at securely sharing a key which is indistinguishable from a randomly chosen one, the requirement of secure communication is authenticity and secrecy of the messages sent in the session.

7.1 Universal Secure Communication Sessions

Secure communication sessions have been discussed in [5, 13] for the general case in which all players are authenticated. We refine our functionality and relax the requirements to the universal model of authentication in the post-specified setting. In that model the players learn their peer identity during the execution of the protocol and must cope with impersonation attacks against the initiator, provided the environment keeps the initiator's identity secret. In which case, we have to expect a real-world adversary that plays the role of the initiator by intercepting the first two protocol rounds, choosing own premaster secret, and completing the protocol in the normal way. The initiator will be unable to terminate the session. Nevertheless, the responder accepts

Functionality $\mathcal{F}_{\text{SCS}}^{(1,2)}$

$\mathcal{F}_{\text{SCS}}^{(1,2)}$ proceeds as follows, when parameterized by a leakage function $l : \{0,1\}^* \to \{0,1\}^*$.

- Upon receiving an input ("establish-session", `SID`, ID_I) from some party, where $ID_I \in (\perp, I)$, record ID_I as initiator, and send the message to the adversary. Upon receiving input ("establish-session", `SID`, R) from some party, record R as responder, and forward the message to the adversary.
- Upon receiving a value ("impersonate", `SID`) from the adversary, do:
  - If (ID_I=$\perp$), check that no `ready` entry exists, and record the adversary as initiator.
  - Else ignore the message.
- Upon receiving a value ("send", `SID`, m, $\bar{P}$) from party P, which is either initiator or responder, check that a record (`SID`, P, $\bar{P}$) exists, record `ready` (if there is no such entry) and send ("sent", `SID`, $l(m)$) to the adversary and a private delayed value ("receive", `SID`, m, P) to $\bar{P}$. Else ignore the message. If the sender is corrupted, then disclose m to the adversary. Next, if the adversary provides m' and no output has been written to the receiver, then send ("send", `SID`, m', P') to the receiver unless P' is an identity of an uncorrupted party.
- Upon receiving a query ("expire-session", `SID`) from either party, un-set the variable `ready`.

Figure 10: The Universal Secure Communication Sessions Functionality

the session and answers to the adversary, mimicking arbitrary party. We capture the requirements by formulating a universal secure communication sessions functionality $\mathcal{F}_{\text{SCS}}^{(1,2)}$ in Fig. 10. Let us highlight some characteristics of $\mathcal{F}_{\text{SCS}}^{(1,2)}$ in the following:

- The functionality handles a uni- and bi-directional model of authentication. The latter is accomplished by invoking the players with their own identity. The first is realized by an environment that invokes the initiator with an empty identity value $\perp$. Then, the adversary is able to mount an impersonation attack. The functionality proceeds in the usual way except that a secure session is established between the adversary and the responder.
- When the adversary has neither impersonated nor corrupted a player, the functionality guarantees that the adversary gains no information other than some leakage information of the message sent while intercepting the session. A leakage function $l(m)$ discloses side channel information about the transmitted plaintext m. In particular, the information leakage includes the length of message m and some message flow information that allow to determine the ideal-world adversary the transmitted messages' source and destination. A pendant to the real world is that the adversary gains some network information about the TLS-protected channel from lower-layers protocols. Further, the leakage reveals the error messages provided by the TLS alert protocol, when either party fails to complete the protocol.

- TLS runs above transport-layer protocols which provide the players with routing information (e.g., IP address, domain) and a communication channel for the session. Furthermore, these protocols typically ensure that the channel is locally fresh by exchanging a pair of nonces. Activating players with session identifier `SID`, the environment mimics these surrounding processes and lower-layers protocols, thus it guarantees in fact a globally unique invocation of the functionality. The session identifier assures also that the functionality may address the initiator though its identity is undisclosed (because it knows the responder's identity and the underlying system model permits a party, i.e., the initiator, to interact with the functionality with an identical session identifier).

- Upon corruption, the functionality allows to fix a message unless the message has been delivered to an uncorrupted party. This requirement makes sure that the adversary cannot mimic other uncorrupted players.

- Messages are provided by the application. The environment mimics these surrounding processes. We assume that the environment "prepares" the messages, i.e., compresses, fragments and adds a sequence number into the encoding. Otherwise, the functionality must provide the technicalities. This would unnecessarily complicate the formulation of the ideal functionality.

7.2 Protocol ρ realizes $\mathcal{F}_{\text{SCS}}^{(1,2)}$

We start with a description the complete TLS protocol for secure communication ρ. It illustrated in Fig. 11. In order to abstract away the comprehensiveness of the TLS protocol and its various cipher suites, we apply Theorem 11 and reformulate protocol ρ in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model, in which the ideal key exchange functionality securely emulates the subroutines. Theorem 3 guarantees that no probabilistic polynomial time-bounded environment distinguishes between the case that it observes an instance of TLS executing the subroutines DHE, DHS and EKT and the case that it interacts with a TLS instance where the subroutines are replaced by the ideal key exchange functionality. We now proceed to prove the main theorem.

Theorem 12 *Protocol ρ in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model securely realizes $\mathcal{F}_{\text{SCS}}^{(1,2)}$.*

Proof. Let $\mathcal{A}$ be a real-world adversary that operates against ρ. We construct an ideal-world adversary $\mathcal{S}$ such that no environment $\mathcal{Z}$ can distinguish between the case that it interacts with $\mathcal{A}$ and parties running ρ in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model or with $\mathcal{S}$ in the ideal world for $\mathcal{F}_{\text{SCS}}^{(1,2)}$. $\mathcal{S}$ runs a simulated copy of $\mathcal{A}$ and mimics an interaction with players executing ρ. It tries to make the internal protocol simulation consistent with the real protocol execution and the limitation that it has no information about the transmitted message m other than its length $l(m)$. The simulator allows the adversary $\mathcal{A}$ to attack the simulated protocol execution in arbitrary way throughout the simulation. $\mathcal{S}$ emulates the protocol execution in such a way that $\mathcal{A}$ thinks that it intercepts a real-world execution of ρ, and such that its interaction with $\mathcal{Z}$ is distributed computationally indistinguishable from that observed by the environment in the real-world execution.

In detail, the simulator proceeds in the following way:

1. **Simulating invocation of I.** Upon receiving ("establish-session", `SID`, ID_I) from $\mathcal{F}_{\text{SCS}}^{(1,2)}$, $\mathcal{S}$ feeds $\mathcal{A}$ with the init message (r_I) where $r_I \xleftarrow{r} \{0,1\}^{p_1(k)}$.

Protocol ρ

1. Upon activation with query ("establish-session", $\mathtt{SID}$, ID_I) by $\mathcal{Z}$, where $ID_I \in (\perp, I)$, the initiator sends the init message (r_I) where $r_I \xleftarrow{r} \{0,1\}^{p_1(k)}$ is a nonce. Upon activation with query ("establish-key", $\mathtt{SID}$, R) by $\mathcal{Z}$, the responder waits for the receipt of the init message. It responds with own nonce $r_R \xleftarrow{r} \{0,1\}^{p_2(k)}$ and initializes a copy of $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ with session identifier $\mathtt{SID}_{\mathrm{KE}}$=$(r_I|r_R)$ by sending query ("establish-key", $\mathtt{SID}_{\mathrm{KE}}$, R) to $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$.

2. Upon receiving the response message, the initiator calls $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ with session identifier $\mathtt{SID}_{\mathrm{KE}}$=$(r_I|r_R)$ on query ("establish-key", $\mathtt{SID}_{\mathrm{KE}}$, ID_I) and waits for the delivery of output ("Key", $\mathtt{SID}_{\mathrm{KE}}$, R, μ). It then computes the session keys $(k_e^I, k_a^I, k_e^R, k_a^R) \leftarrow \mathtt{PRF}_\mu(l_2)$ and the finished value $F_I \leftarrow \mathtt{PRF}_\mu(l_3)$. Additionally, the initiator sends the final initiator message $(\mathtt{E}_{k_e^I}(F_I|\mathtt{HMAC}_{k_a^I}(F_I)))$.

3. When the responder receives the final initiator message (α), it first waits for the delivery of ("Key", $\mathtt{SID}_{\mathrm{KE}}$, ID_I, μ) from $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$. Then, the responder computes in the same way the session keys $(k_e^I, k_a^I, k_e^R, k_a^R) \leftarrow \mathtt{PRF}_\mu(l_2)$ for the players. It decrypts the final initiator message $(F_I|t_I) \leftarrow \mathtt{D}_{k_e^I}(\alpha)$ and verifies that $F_I \leftarrow \mathtt{PRF}_\mu(l_3)$ and $t_I \leftarrow \mathtt{HMAC}_{k_a^I}(F_I)$. If the verification fails, it aborts. Otherwise, it computes the finished value $F_R \leftarrow \mathtt{PRF}_\mu(l_4)$ and sends the final responder message $(\mathtt{E}_{k_e^R}(F_R|\mathtt{HMAC}_{k_a^R}(F_R)))$.

4. Upon delivery of the final responder message (β), the initiator decrypts the message $(F_R|t_R) \leftarrow \mathtt{D}_{k_e^R}(\beta)$. Then, it verifies that $F_R \leftarrow \mathtt{PRF}_\mu(l_4)$ and $t_R \leftarrow \mathtt{HMAC}_{k_a^R}(F_R)$. If the verification fails, it aborts.

5. Once the session keys are agreed upon, the sender $P \in (I, R)$ waits for the transmission notification ("send", $\mathtt{SID}$, m, $\bar{P}$) from $\mathcal{Z}$. It then sends $\mathtt{E}_{k_a^P}(m|t_m)$ whereby message m is authenticated through the tag $t_m \leftarrow \mathtt{HMAC}_{k_a^P}(m)$. Upon receiving the message γ, the receiver $\bar{P}$ decrypts the message $(m|t_m) \leftarrow \mathtt{D}_{k_a^P}(\gamma)$ and verifies that $t_m \leftarrow \mathtt{HMAC}_{k_a^P}(m)$. If the verification fails, it aborts. Otherwise, the receiver accepts the message and makes the local output ("receive", $\mathtt{SID}$, m, P) to $\mathcal{Z}$.

Figure 11: The full TLS Framework Structure, in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-hybrid Model

2. **Simulating invocation of R.** Upon receiving ("establish-session", $\mathtt{SID}$, R) from $\mathcal{F}_{\mathrm{SCS}}^{(1,2)}$, $\mathcal{S}$ waits for receipt of an init message (r'_I) from $\mathcal{A}$. Then, it chooses a nonce $r_R \xleftarrow{r} \{0,1\}^{p_2(k)}$ and feeds $\mathcal{A}$ with the response message (r_R, R). Finally, it calls $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ on query ("establish-key", $\mathtt{SID}'_{\mathrm{KE}}$, R), where $\mathtt{SID}'_{\mathrm{KE}}$=$(r'_I|r_R)$.

3. **Simulating receipt of a response message by I.** Upon $\mathcal{A}$ delivers the message (r'_R, P') to I, $\mathcal{S}$ proceeds as follows:

   (a) $\mathcal{S}$ verifies that I has previously sent the init message (r_I).

   (b) $\mathcal{S}$ mimics on behalf of I the master key generation by invoking a copy of $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$. The

master key is obtained by handing $\mathcal{F}_{\text{KE}}^{(1,2)}$ the message ("establish-key", $\texttt{SID}_{\text{KE}}$, ID_I), where $\texttt{SID}_{\text{KE}}=(r_I|r_R)$ and waiting for the delivery of the response message ("Key", $\texttt{SID}_{\text{KE}}$, R, μ). Otherwise, $\mathcal{S}$ terminates with an internal error message.

(c) $\mathcal{S}$ defines the master key μ, the session keys $(k_e^I, k_a^I, k_e^R, k_a^R)$, and the finished value F_I to be random values Δ_{k_m}, $(\Delta_{k_e^I}, \Delta_{k_a^I}, \Delta_{k_e^R}, \Delta_{k_a^R})$, and Δ_{F_I} chosen from the appropriate spaces, respectively.

(d) $\mathcal{S}$ feeds $\mathcal{A}$ with the final initiator message $(\texttt{E}_{\Delta_{k_e^I}}(\Delta_{F_I}|t_I))$, where $t_I \leftarrow \texttt{HMAC}_{\Delta_{k_a^I}}(\Delta_{F_I})$.

4. **Simulating receipt of a final initiator message by R.** When $\mathcal{A}$ delivers the message (α) to R, $\mathcal{S}$ proceeds as follows:

(a) $\mathcal{S}$ verifies that it has previously received an init message (r_I') and sent a response message (r_R, R).

(b) $\mathcal{S}$ waits for the master key by mimicking the key establishment process of $\mathcal{F}_{\text{KE}}^{(1,2)}$. Now we distinguish between the following two distinct cases.

Case 1 ($ID_I = I$): If $\mathcal{S}$ receives an answer ("Key", $\texttt{SID}_{\text{KE}}$, ID_I, μ) from $\mathcal{F}_{\text{KE}}^{(1,2)}$, then no impersonation attack has occurred. In this case $\mathcal{S}$ uses for k_m, $(k_e^I, k_a^I, k_e^R, k_a^R)$, and F_I exactly the same values Δ_{k_m}, $(\Delta_{k_e^I}, \Delta_{k_a^I}, \Delta_{k_e^R}, \Delta_{k_a^R})$, and Δ_{F_I} that it has chosen on behalf of the initiator before. Then, it waits for the delivery of the final initiator message and applies the session keys to decrypt $(F_I'|t_I') \leftarrow \texttt{D}_{\Delta_{k_e^I}}(\alpha)$. $\mathcal{S}$ compares whether $F_I' = \Delta_{F_I}$ and $t_I' = t_I$. If the verification fails, it aborts the simulation. Otherwise, it chooses F_R to be a random value Δ_{F_R} from the same space and feeds $\mathcal{A}$ with the final responder message $(\texttt{E}_{\Delta_{k_e^R}}(\Delta_{F_R}|t_R))$ where $t_R \leftarrow \texttt{HMAC}_{\Delta_{k_a^R}}(\Delta_{F_R})$. Then, $\mathcal{S}$ prepares for the secure message exchange on behalf of R.

Case 2 ($ID_I = \perp$): If $\mathcal{S}$ receives an answer ("Key", $\texttt{SID}_{\text{KE}}'$, P', $\tilde{\mu}$), then the original master key has been modified by the adversary implying the impersonation attack framing the initiator. In this case $\mathcal{S}$ computes $(k_e^I, k_a^I, k_e^R, k_a^R)$, and F_I as specified in the protocol, i.e., $(k_e^I, k_a^I, k_e^R, k_a^R) \leftarrow \texttt{PRF}_{\tilde{\mu}}(l_2)$, and $F_I \leftarrow \texttt{PRF}_{\tilde{\mu}}(l_3)$. Then, it waits for the delivery of the final initiator message and decrypts $(F_I'|t_I') \leftarrow \texttt{D}_{k_e^I}(\alpha)$. $\mathcal{S}$ compares whether $F_I' = F_I$ and $t_I' = \texttt{HMAC}_{k_a^I}(F_I)$. If the verification fails, it aborts the simulation. Otherwise, $\mathcal{S}$ computes $F_R \leftarrow \texttt{PRF}_{\tilde{\mu}}(l_4)$, and feeds $\mathcal{A}$ with the final responder message $(\texttt{E}_{k_e^R}(F_R|\texttt{HMAC}_{k_a^R}(F_R)))$. Finally, $\mathcal{S}$ sends ("impersonate", $\texttt{SID}$) to $\mathcal{F}_{\text{SCS}}^{(1,2)}$. This is exactly the point in the simulation where the adversary has impersonated the unauthenticated party. Then, $\mathcal{S}$ continues the simulation with the exception that the interaction proceeds with $\mathcal{A}$ and I aborts the protocol.

Note that in all subsequent simulation steps, $\mathcal{S}$ uses session keys (k_a^P, k_a^P) for $P \in (I, R)$ obtained from one of the above two cases.

5. **Simulating receipt of a final responder message by I.** When $\mathcal{A}$ delivers the message (β) to an uncorrupted I, $\mathcal{S}$ proceeds as follows:

(a) $\mathcal{S}$ verifies that it has previously sent an init message (r_I), received a response message (r_R', P'), and sent a final initiator message $(\texttt{E}_{k_e^I}(F_I|t_I))$.

(b) $\mathcal{S}$ uses its own session keys (k_e^R, k_a^R) to decrypt β obtaining $F'_R|t'_R$. Since no responder impersonation attacks may occur it aborts the simulation if $F'_R \neq F_R$ or $t'_R \neq t_R$ whereby F_R and t_R are the values used by $\mathcal{S}$ on behalf of R in the previous simulation step. If the simulation does not abort then $\mathcal{S}$ prepares for the secure message exchange on behalf of I.

6. **Simulating Message Transmission.** Upon receiving ("sent", SID, $l(m)$) from $\mathcal{F}_{\text{SCS}}^{(1,2)}$, $\mathcal{S}$ extracts from $l(m)$ the sender and receiver identities. It then chooses a random message $\Delta_m \xleftarrow{r} \{0,1\}^{l(m)}$ and feeds $\mathcal{A}$ with message $\mathtt{E}_{k_a^P}(\Delta_m|t_{\Delta_m})$ where $t_{\Delta_m} \leftarrow \mathtt{HMAC}_{k_a^P}(\Delta_m)$.

7. **Simulating Message Reception.** Upon receiving the message (γ), the receiver decrypts the message $(\Delta'_{m'},t'_{\Delta_{m'}}) \leftarrow \mathtt{D}_{k_a^P}(\gamma)$ and then verifies that $t'_{\Delta_{m'}} \leftarrow \mathtt{HMAC}_{k_a^P}(\Delta_m)$ using its own keys. If the verification fails, it aborts. Otherwise, $\mathcal{S}$ signals $\mathcal{F}_{\text{SCS}}^{(1,2)}$ to send the message.

8. **Simulating Static Corruption** If one of the parties gets corrupted, then $\mathcal{S}$ proceeds by emulating a ρ protocol session, just as a honest party would play it. In particular, $\mathcal{S}$ uses the message m transmitted by $\mathcal{F}_{\text{SCS}}^{(1,2)}$ in the emulation of the last protocol round.

Analysis of $\mathcal{S}$ We now prove that the simulator $\mathcal{S}$ is such that no environment $\mathcal{Z}$ can distinguish between the ideal execution of $\mathcal{F}_{\text{SCS}}^{(1,2)}$ and $\mathcal{S}$, and the real execution of ρ and $\mathcal{A}$ in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model. To prove that the environment's view is indistinguishable in the two worlds, we define a sequence of hybrid distributions, $\mathcal{H}_\mathbf{1}$ and $\mathcal{H}_\mathbf{2}$, where we make some modifications, starting from the real-world protocol execution and ending at the ideal-world execution. We show

$$\mathtt{UC{-}EXEC}^{\mathcal{F}_{\text{KE}}^{(1,2)}}_{\rho,\mathcal{A},\mathcal{Z}} \approx \mathcal{H}_\mathbf{1} \approx \mathcal{H}_\mathbf{2} \approx \mathtt{UC{-}EXEC}_{\mathcal{F}_{\text{SCS}}^{(1,2)},\mathcal{S},\mathcal{Z}} \tag{1}$$

The distributions are defined as follows:

- $\mathcal{H}_\mathbf{1}$ takes the distribution of the output of $\mathcal{Z}$ which is identical to a real execution of $\mathcal{A}$ running ρ in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model with the following exception. When the protocol ρ instructs the initiator (respectively the responder) to evaluate the pseudo-random function $\mathtt{PRF}()$ with some random key Δ_{k_m} in the range of the master secret in order to compute the session keys (k_a^P, k_a^P) for $P \in (I, R)$ and the finished values F_I and F_R, we replace the outputs with the independently chosen random values $(\Delta_{k_a^P}, \Delta_{k_a^P})$, Δ_{F_I}, and Δ_{F_R}, respectively, all in the range of $\mathtt{PRF}()$.

- $\mathcal{H}_\mathbf{2}$ is identical to $\mathcal{H}_\mathbf{1}$ with the following exception. The simulation fails if on behalf of the responder (resp. initiator) it receives a message which it decrypts with k_a^P and obtains a valid authentication tag which has not been generated by the simulator on behalf of the initiator (resp. responder) before. Since the authentication tags are generated via the $\mathtt{HMAC}()$ function the simulation failure occurs whenever the adversary forges a corresponding tag.

CLAIM 13 *If* $\mathtt{PRF}()$ *is a secure pseudo-random function, then* $\mathtt{UC{-}EXEC}^{\mathcal{F}_{\text{KE}}^{(1,2)}}_{\rho,\mathcal{A},\mathcal{Z}} \approx \mathcal{H}_\mathbf{1}$.

PROOF. Assume that such environment $\mathcal{Z}$ that distinguishes with non-negligible simulation slice between the interaction $\mathtt{UC{-}EXEC}^{\mathcal{F}_{\text{KE}}^{(1,2)}}_{\rho,\mathcal{A},\mathcal{Z}}$ and $\mathcal{H}_\mathbf{1}$ exists. We construct an adversary $\mathcal{D}$ that has

access to the black-box oracle $\mathcal{O}_{\mathcal{P}}$. The oracle computes throughout the whole simulation either $\mathtt{PRF}_{\Delta_{k_m}}()$ for some randomly chosen secret Δ_{k_m} or a truly random function with the same range.

$\mathcal{D}$ runs a copy of $\mathcal{Z}$ and mimics the roles of $\mathcal{A}$ and the parties. It emulates the interaction with parties running ρ in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-hybrid model until it is required to compute the session keys (k_a^P, k_a^P) or the finished values F_I or F_R which it obtains through the corresponding call to $\mathcal{O}_{\mathcal{P}}$. Similar to the previous claims, $\mathcal{D}$ interpolates between $\mathtt{UC{-}EXEC}^{\mathcal{F}_{\mathrm{KE}}^{(1,2)}}_{\rho,\mathcal{A},\mathcal{Z}}$ (this is the case where $\mathcal{O}_{\mathcal{P}}$ contains $\mathtt{PRF}_{\Delta_{k_m}}()$) and $\mathcal{H}_{\mathbf{1}}$ (this is the case where $\mathcal{O}_{\mathcal{P}}$ contains the truly random function) so that based on the output of $\mathcal{O}_{\mathcal{P}}$, $\mathcal{D}$ can distinguish with non-negligible probability between $\mathtt{UC{-}EXEC}^{\mathcal{F}_{\mathrm{KE}}^{(1,2)}}_{\rho,\mathcal{A},\mathcal{Z}}$ and $\mathcal{H}_{\mathbf{a}}$ for any $\mathcal{Z}$. □

Claim 14 *If* $\mathtt{HMAC}()$ *is a WUF-CMA secure message authentication function, then* $\mathcal{H}_{\mathbf{1}} \approx \mathcal{H}_{\mathbf{2}}$.

Proof. Assume that there exists an environment $\mathcal{Z}$ that distinguishes with non-negligible simulation slice between the interaction $\mathcal{H}_{\mathbf{1}}$ and $\mathcal{H}_{\mathbf{2}}$. We specify an adversary $\mathcal{F}$ that has access to an authentication oracle $\mathcal{O}_{\mathcal{M}}$ which on input some message m outputs the corresponding authentication tag $\mathtt{HMAC}_{\Delta_{k_a^P}}(m)$ using some initially fixed random key $\Delta_{k_a^P}$ unknown to $\mathcal{F}$ and the corresponding verification oracle $\mathcal{V}_{\mathcal{M}}$ that on input a message m and a candidate authentication tag t outputs $\mathtt{valid}$ if $t \leftarrow \mathtt{HMAC}_{\Delta_{k_a^P}}(m)$, otherwise it returns $\mathtt{invalid}$. We show how $\mathcal{F}$ can use $\mathcal{Z}$ in order to output a valid tag $\tilde{t}$ for some message $\tilde{m}$ which has not been previously queried to $\mathcal{O}_{\mathcal{M}}$. In fact $\mathcal{F}$ emulates the execution of ρ in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-hybrid model and mimics for $\mathcal{Z}$ the role of the adversary and the players with the exception that the required authentication tags for the computed protocol messages $\mathcal{F}$ obtains from $\mathcal{O}_{\mathcal{M}}$. Thus, the environment can easily recognize when some forgery is received during the emulation of the protocol with respect to the adversary. This forgery is also the output of $\mathcal{F}$. Thus, unless the simulation fails both hybrids $\mathcal{H}_{\mathbf{1}}$ and $\mathcal{H}_{\mathbf{2}}$ proceed identical. By assumption the simulation fails only with negligible probability. □

Claim 15 *If* $\mathtt{E}()$, $\mathtt{D}()$ *is an IND-CPA secure encryption scheme, then* $\mathcal{H}_{\mathbf{2}} \approx \mathtt{UC{-}EXEC}_{\mathcal{F}_{\mathrm{SCS}}^{(1,2)},\mathcal{S},\mathcal{Z}}$.

Proof. Assume that such environment $\mathcal{Z}$ that distinguishes with non-negligible simulation slice between the interaction $\mathcal{H}_{\mathbf{2}}$ and $\mathtt{UC{-}EXEC}_{\mathcal{F}_{\mathrm{SCS}}^{(1,2)},\mathcal{S},\mathcal{Z}}$ exists. We construct an adversary $\mathcal{D}$ that has access to the *Real-Or-Random* encryption oracle $\mathcal{O}_{\mathcal{E}}$ that on input a message m outputs $\mathtt{E}_{\Delta_{k_a^P}}(m_b)$ using some initially fixed random unknown key $\Delta_{k_a^P}$ and bit b such that in case $b = 0$ the oracle uses $m_0 \leftarrow m$ and in case $b = 1$ it uses $m_1 \leftarrow \tilde{m}$ for some $\tilde{m} \neq m$. We show how $\mathcal{S}_{\mathcal{E}}$ breaks the security of the symmetric encryption scheme, i.e., decides on b with a probability non-negligibly larger than 1/2.

$\mathcal{D}$ runs a copy of $\mathcal{A}$ and emulates the interaction with parties running ρ in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-hybrid model except that whenever $\mathcal{D}$ is required to obtain a ciphertext on some message m' (which can be one of the finished messages $F_I|t_{F_I}$ or $F_R|t_{F_R}$, or the securely transmitted message $m|t_m$) it queries $\mathcal{O}_{\mathcal{E}}$ on m'. In this way $\mathcal{D}$ interpolates between $\mathcal{H}_{\mathbf{2}}$ (this is the case where the oracle encrypts a real message, i.e., $b = 0$) and $\mathtt{UC{-}EXEC}_{\mathcal{F}_{\mathrm{SCS}}^{(1,2)},\mathcal{S},\mathcal{Z}}$ (this is the case where the oracle encrypts a different message, i.e., $b = 1$). Based on the output of $\mathcal{Z}$ $\mathcal{D}$ can decide on the bit b used by $\mathcal{O}_{\mathcal{E}}$. Thus, by assumption $\mathcal{H}_{\mathbf{2}} \approx \mathtt{UC{-}EXEC}_{\mathcal{F}_{\mathrm{SCS}}^{(1,2)},\mathcal{S},\mathcal{Z}}$ for any $\mathcal{Z}$. □

This completes the simulation and the proof of the theorem. ■

8 Conclusion

We have analyzed the TLS protocol family in the framework of Universal Composition. We have shown that the complete TLS protocol framework securely realizes secure communication sessions. Thus, future analysis of composite protocols can be considerably simplified by calling the secure communication functionality in the hybrid-model reformulation. The composition theorem preserves that security holds under general composition with arbitrary players. Furthermore, since the ideal functionality is free from any probabilism, it may be expressed in an abstract term algebra, enabling Dolev-Yao style proofs. It allows applying automated proof tools while preserving soundness and composition. In many analysis of Internet protocols TLS has been abstracted away in a Dolev-Yao-style and assumed to provide secure communication channels [25, 23, 24]. Our analysis is a first step towards showing whether this claim is founded.

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A Impossibility of UC-secure TLS Handshakes

The main pillar of the TLS framework are the various handshake protocols that negotiate cryptographic key material for the instantiation of secure communication channels. We would like to claim that the analysis can be carried out in the spirit of a hybrid-model reformulation. We wish to show that the native TLS composition, i.e., establish the session keys by the handshake protocols and then use some cryptography to build secure channels by the record-layer protocols, is secure under the universal composable notion. That means, we would like to show that the handshake protocols securely emulate ideal key exchange and the record layer simply syncs the keys to establish the secure session as suggested in [13]. A separated consideration of the handshake protocols would have wide applicability and is of independent interest because there exist also protocols that take advantage of the functionality provided by the handshake protocol (e.g., the IEEE 802.11 EAP protocol family [1]). Unfortunately, it turns out that the handshake protocols are neither secure under the strong nor relaxed notion of UC security. This is due to a commitment problem in the confirmation of the session keys. The environment can test whether the session keys origin from the interaction with the real protocol in presence of $\mathcal{A}$ or ideal protocol in presence of $\mathcal{S}$.

A.1 Protocol π does not realize $\mathcal{F}_{\text{KE+}}^{(1,2)}$

We would like to claim that the handshake protocol UC-realizes an ideal key exchange functionality. Unfortunately, such a strong claim does not hold. To understand why, we formulate the handshake protocols π in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model. Protocol π in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model is illustrated in Fig. 12. To capture the ideal-world security requirements of π, we make use of functionality $\mathcal{F}_{\text{KE+}}^{(1,2)}$. Essentially, it is a copy of $\mathcal{F}_{\text{KE}}^{(1,2)}$ with the exception that $\mathcal{F}_{\text{KE+}}^{(1,2)}$ outputs the session keys κ (instead of the master secret μ).

Theorem 16 *Protocol π in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model does not securely realize $\mathcal{F}_{\text{KE+}}^{(1,2)}$.*

Proof.(Sketch) There exists an environment $\mathcal{Z}$ that distinguishes with non-negligible probability whether it communicates with the players interacting with the ideal key exchange functionality $\mathcal{F}_{\text{KE+}}^{(1,2)}$ in the presence of adversary $\mathcal{S}$ and players interacting with protocol π in the $\mathcal{F}_{\text{KE}}^{(1,2)}$-hybrid model in front of the adversary $\mathcal{A}$. We construct such $\mathcal{Z}$ as follows:

$\mathcal{Z}$ activates first I with input ("establish-key, SID, ID_I), instructs the adversary to deliver the init message (r_I), and records the value. Similarly, it activates R with input ("establish-key, SID, R), instructs the adversary to deliver the response message (r_R), and records the value. It then instructs the adversary to send the final initiator message ($\mathtt{E}_{k_e^I}(F_I|\mathtt{HMAC}_{k_a^I}(F_I))$), records the

Protocol π

1. Upon activation with query ("establish-key", $\mathtt{SID}$, ID_I), where ID_I=$(\perp, I)$, the initiator sends the init message (r_I) where $r_I \xleftarrow{r} \{0,1\}^{p_1(k)}$ is a nonce. Upon activation with query ("establish-key", $\mathtt{SID}$, R), the responder waits for the receipt of the init message. It responds with own nonce $r_R \xleftarrow{r} \{0,1\}^{p_2(k)}$ and initializes a copy of $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ with session identifier $\mathtt{SID}_{\mathrm{KE}}$=$(r_I|r_R)$ by sending query ("establish-key", $\mathtt{SID}_{\mathrm{KE}}$, R) to $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$.

2. Upon receiving the response message, the initiator calls $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$ with session identifier $\mathtt{SID}_{\mathrm{KE}}$=$(r_I|r_R)$ on query ("establish-key", $\mathtt{SID}_{\mathrm{KE}}$, ID_I) and waits for the delivery of output ("Key", $\mathtt{SID}_{\mathrm{KE}}$, R, μ). It then computes the session keys $(k_e^I, k_a^I, k_e^R, k_a^R) \leftarrow \mathtt{PRF}_\mu(l_2)$ and the finished value $F_I \leftarrow \mathtt{PRF}_\mu(l_3)$. Additionally, the initiator sends the final initiator message $(\mathtt{E}_{k_e^I}(F_I|\mathtt{HMAC}_{k_a^I}()))$.

3. When the responder receives the final initiator message (α), it first waits for the delivery of ("Key", $\mathtt{SID}_{\mathrm{KE}}$, ID_I, μ) from $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$. Then, the responder computes in the same way the session keys $(k_e^I, k_a^I, k_e^R, k_a^R) \leftarrow \mathtt{PRF}_\mu(l_2)$. It decrypts the final initiator message $(F_I|t_I) \leftarrow \mathtt{D}_{k_e^I}(\alpha)$ and verifies that $F_I \leftarrow \mathtt{PRF}_\mu(l_3)$ and $t_I \leftarrow \mathtt{HMAC}_{k_a^I}(F_I)$. If the verification fails, it aborts. Otherwise, it computes the finished value $F_R \leftarrow \mathtt{PRF}_\mu(l_4)$ and sends the final responder message $(\mathtt{E}_{k_e^R}(F_R|\mathtt{HMAC}_{k_a^R}(F_R)))$. Finally, the responder terminates with local output ("Key", $\mathtt{SID}$, ID_I, κ) for κ=$(k_e^I, k_a^I, k_e^R, k_a^R)$.

4. Upon delivery of the final responder message (β), the initiator decrypts the message $(F_R|t_R) \leftarrow \mathtt{D}_{k_e^R}(\beta)$. Then, it verifies that $F_R \leftarrow \mathtt{PRF}_\mu(l_4)$ and $t_R \leftarrow \mathtt{HMAC}_{k_a^R}(F_R)$. If the verification fails, it aborts. Otherwise, the initiator locally outputs ("Key", $\mathtt{SID}$, R, κ) for κ=$(k_e^I, k_a^I, k_e^R, k_a^R)$.

Figure 12: The TLS Handshake Protocol Structure, in the $\mathcal{F}_{\mathrm{KE}}^{(1,2)}$-Hybrid model. **Note**, the finished values are the first messages, which are encrypted and authenticated with the session keys [18, Sect. 7.4.9.].

message, and waits for the output ("Key", $\mathtt{SID}$, ID_I, κ) from R. Recall that κ=$(k_e^I, k_a^I, k_e^R, k_a^R)$ are the session keys. Finally, it decrypts the final initiator message $(F_I|t_I)$ and verifies that $t_I \leftarrow \mathtt{HMAC}_{k_a^I}(F_I)$. If the verification fails, $\mathcal{Z}$ outputs "ideal world", otherwise it outputs "real world".

Obviously, when $\mathcal{Z}$ communicates with protocol π in the presence of adversary $\mathcal{A}$, it outputs "real world". By contrast, when interacting with the ideal key exchange functionality $\mathcal{F}_{\mathrm{KE+}}^{(1,2)}$ in the presence of adversary $\mathcal{S}$, the session keys, say $\tilde{\kappa}$=$(\tilde{k}_e^I, \tilde{k}_a^I, \tilde{k}_e^R, \tilde{k}_a^R)$, are independently and randomly chosen by the functionality. The keys are potentially different from κ. $\mathcal{S}$ has to come up with the session keys κ without ever seeing the master secret μ. Otherwise $\mathcal{Z}$ decrypts the final initiator message using $\tilde{k}_e^I$ and notices that $t_I \nleftarrow \mathtt{HMAC}_{\tilde{k}_a^I}(F_I)$. Consequently the probability that $\mathcal{S}$ simulates the required session keys is $1/k$ and $\mathcal{Z}$ distinguishes between the ideal and real process with overwhelming probability. ■

Relaxed UC-Security We wish to formulate a relaxed definition of the key exchange functionality $\mathcal{F}_{\text{KE+}}^{\mathcal{N},(1,2)}$ and would like to claim that protocol π relaxed UC-realizes $\mathcal{F}_{\text{KE+}}^{\mathcal{N},(1,2)}$. The functionality differs from the previous definition in that it sets the session key to the local output of the non-information oracle $\mathcal{N}$ unless the initiator is uncorrupted. Otherwise, it fixes the session key as the adversary. See [27, p. 117] for a definition. Unfortunately, such a claim does not hold either.

CLAIM 17 *Protocol π does not relaxed UC-secure realize $\mathcal{F}_{\text{KE+}}^{\mathcal{N},(1,2)}$.*

PROOF.(Sketch) We use the equivalence between relaxed-UC security and SK-security, and show that π is not an SK-secure key exchange protocol, even in front of a passive adversary. Adversary $\mathcal{A}$ instructs honest parties to run a session of the π protocol, of which $\mathcal{A}$ stores all messages. Then, $\mathcal{A}$ makes a test query for that session, and receives a pair of keys $(k_e^I, k_a^I, k_e^R, k_a^R)$ that are either the keys corresponding to the executed session or random keys, with probability one half. In order to decide whether it sees real or random keys, $\mathcal{A}$ simply uses those keys to decrypt the third and fourth messages of the π session it monitored, checks the MACs, and decides that it received the real keys if the decryption and MAC verification succeeded. This adversary makes the correct guess with overwhelming probability, contradicting the SK-security definition. ■

Remark. The impossibility of simulating the key exchange functionality is due to the fact that the players commit to the finished value by tagging before encrypting the message. In fact, the problem is similar to the general commitment problem presented in [8]. Since the finished values are identically computed, the impossibility result applies to all cipher suites. However, the result does not indicate that the handshake protocols are insecure. The implication is that the desired native composition is infeasible and one has to search for a composition, which is close to the native TLS framework structure. A promising approach seems to be to naturally relax the key exchange functionality in the form that it outputs the the session keys and two random values in the space of the finished values and analyze whether a TLS-like handshake protocol without the commitment to the finished values securely emulates the functionality.

B TLS Protocol Framework

B.1 Handshake Protocol Structure

The TLS handshake provides a framework of messages and cryptographic tools and an abbreviated handshake can be used to continue an earlier *session,* using its master secret to derive new record-layer keys. The full SSL/TLS handshake is performed as follows:

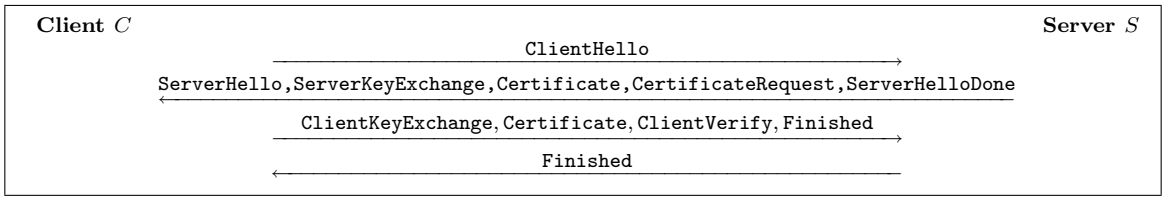

Figure 13: TLS handshake protocol skeleton

1. The client initiates the handshake with a `ClientHello` message, naming the cipher suites that it supports, and optionally compression algorithms to be used. It also chooses a client random nonce.

2. The server responds with a `ServerHello` message, in which it selects one of the proposed cipher suites. It defines a session identifier and chooses a server random nonce. In most cipher suites, a `Certificate` message is sent for server authentication, providing a certified public key of the server. Depending on the cipher suite, a `ServerKeyExchange` message is sent (e.g., for ephemeral Diffie-Hellman). If the server opts for client authentication, it transmits a `CertificateRequest` message. A `ServerHelloDone` message denotes the end of this handshake message flow.

3. The client responds with a `ClientKeyExchange` message containing either the encrypted premaster secret, or other key exchange information. If client authentication has been requested, the client can provide it by sending a `Certificate` message, and adding a `CertificateVerify` message signing the digest of all previous handshake messages. Then, the client sends a message in the change cipher spec subprotocol to indicate that it is going to start using the negotiated cipher suite and key material. Finally, the client signals the end of handshake with the `Finished` message, which contains a digest value over all previous handshake messages; this is the first time for the SSL/TLS record-layer protocol to use the new cryptographic parameters.

4. The server finishes the handshake by sending a change cipher spec message of its own (indicating that subsequent records will use the new cryptographic parameters), followed by a `Finished` message containing the digest of all previous messages (including the client's `Finished`).

B.2 Cipher Suites

Client and server determine the algorithms to be used to secure the communication by agreeing on a cipher suite. Each cipher suite consists of an algorithm for key exchange algorithm, a symmetric encryption algorithm, and a MAC algorithm. For example, the cipher suite TLS_RSA_WITH_AES_256_CBC_SHA-1 specifies the key transport algorithm RSA, the symmetric encryption algorithm AES-256 in CBC mode, and the MAC algorithm SHA-1. In order to negotiate the premaster secret, TLS provides a number of mechanisms including the following illustrated in Tab. 1.

B.3 Key Derivation

After establishing the premaster secret, both client and server compute the master secret as follows:

$$master\ secret = \texttt{PRF}_{premaster\ secret}(\text{“}\texttt{master secret}\text{”}, client\ nonce \,||\, server\ nonce)$$

The session keys are derived from:

$$cryptographic\ keys = \texttt{PRF}_{master\ secret}(\text{“}\texttt{key expansion}\text{”}, server\ nonce \,||\, client\ nonce)$$

`PRF` denotes a special TLS construction for a pseudo random function taking three inputs: a secret key, a string *label*, and a seed (where the latter two are not assumed to be secret). The construction

CipherSuite	Key Exchange	Cipher	Hash
TLS_NULL_WITH_NULL_NULL	NULL	NULL	NULL
TLS_RSA_WITH_NULL_MD5	RSA	NULL	MD5
TLS_RSA_WITH_NULL_SHA	RSA	NULL	SHA
TLS_RSA_WITH_RC4_128_MD5	RSA	RC4_128	MD5
TLS_RSA_WITH_RC4_128_SHA	RSA	RC4_128	SHA
TLS_RSA_WITH_IDEA_CBC_SHA	RSA	IDEA_CBC	SHA
TLS_RSA_WITH_DES_CBC_SHA	RSA	DES_CBC	SHA
TLS_RSA_WITH_3DES_EDE_CBC_SHA	RSA	3DES_EDE_CBC	SHA
TLS_DH_DSS_WITH_DES_CBC_SHA	DH_DSS	DES_CBC	SHA
TLS_DH_DSS_WITH_3DES_EDE_CBC_SHA	DH_DSS	3DES_EDE_CBC	SHA
TLS_DH_RSA_WITH_DES_CBC_SHA	DH_RSA	DES_CBC	SHA
TLS_DH_RSA_WITH_3DES_EDE_CBC_SHA	DH_RSA	3DES_EDE_CBC	SHA
TLS_DHE_DSS_WITH_DES_CBC_SHA	DHE_DSS	DES_CBC	SHA
TLS_DHE_DSS_WITH_3DES_EDE_CBC_SHA	DHE_DSS	3DES_EDE_CBC	SHA
TLS_DHE_RSA_WITH_DES_CBC_SHA	DHE_RSA	DES_CBC	SHA
TLS_DHE_RSA_WITH_3DES_EDE_CBC_SHA	DHE_RSA	3DES_EDE_CBC	SHA
TLS_DH_anon_WITH_RC4_128_MD5	DH_anon	RC4_128	MD5
TLS_DH_anon_WITH_DES_CBC_SHA	DH_anon	DES_CBC	SHA
TLS_DH_anon_WITH_3DES_EDE_CBC_SHA	DH_anon	3DES_EDE_CBC	SHA

Table 1: TLS Cipher Suites

of PRF is as follows (see [18, Sect. 5]).

$$\texttt{PRF}_{secret}(label, seed) = P_{MD5}(S_1, label \,||\, seed) \oplus P_{SHA\text{-}1}(S_2, label \,||\, seed),$$

where S_1 and S_2 are the halves of *secret* and where $\texttt{P}_{hash}$ is a HMAC-based data expansion function that expands a secret and a seed into an arbitrary quantity of output as follows:

$$\begin{aligned} P_{hash}(secret, seed) = & HMAC_{hash}(secret, A_1 \,||\, seed) \,|| \\ & HMAC_{hash}(secret, A_2 \,||\, seed) \,|| \\ & HMAC_{hash}(secret, A_3 \,||\, seed) \,|| \ldots \end{aligned}$$

Here || indicates concatenation, *hash* denotes the name of a specific hash function, and we define

$$\begin{aligned} A_0 &= seed, \\ A_i &= HMAC_{hash}(secret, A_{i-1}). \end{aligned}$$