# Unconditionally-Secure Universally Composable Password-Based Key-Exchange based on One-Time Memory Tokens

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**Abstract.** We continue the recent trend in cryptography to study protocol design in presence of tamper-proof hardware tokens. We present a very efficient protocol for password-based authenticated key exchange based on the weak model of one-time memory tokens, recently introduced by Goldwasser et al. (Crypto 2008). Our protocol only requires four moves, very basic operations, and the sender to send $\ell$ tokens in the first step for passwords of length $\ell$. At the same time we achieve information-theoretic security in Canetti's universal composition framework (FOCS 2001) against adaptive adversaries (assuming reliable erasure), even if the tokens are not guaranteed to be transferred in an authenticated way, i.e., even if the adversary can read or substitute transmitted tokens (as opposed to many previous efforts).

**Keywords.** password based key exchange, information-theoretic, one-time memory tokens

## 1 Introduction

Recently, the area of designing cryptographic protocols in the presence of hardware tokens has gained significant attention [27, 28, 26, 21, 35, 12, 15, 30, 23, 22, 18, 17, 5, 16, 13]. The pleasant aspect of using hardware tokens is that they facilitate the design of efficient protocols while still achieving strong security guarantees. Notably, most of the aforementioned works design protocols in Canetti's universal composition (UC) framework [7]. There are two reasons for this: (a) the derived protocols provide strong security even in "hostile" environments (e.g., in presence of concurrent executions of other protocols), and (b) vice versa, the UC framework allows an easy modeling of such tokens.

### 1.1 Hardware Tokens in Cryptography

The aforementioned works use different types of hardware tokens, ranging from very powerful smartcards as in [28] to very minimalist one-time memory tokens as in [21]. The latter are tamper-proof tokens which allow a sender to store two values and to transfer the token to a receiver, and when one of the two values is read according to a receiver's choice, the other value collapses and is irrevocably lost. Although resembling a hardware-based implementation of an oblivious transfer protocol, Goldwasser et al. [21] point out that these tokens implement a weaker functionality than oblivious transfer. The main difference is that the (single) measurement on the receiver's side is carried out locally such that an adversarial receiver may postpone the measurement instead of performing the read-out when supposed to do. This is in contrast to an execution of an oblivious transfer protocol where the sender knows when the transfer takes place (and can therefore enforce a transfer at a certain point in a protocol run).

There are essentially two types of approaches using tokens: some devised protocols focus on concrete problems [26, 30, 18, 5] like secure computation of set intersections and aim at providing really efficient protocols. Others provide general feasibility results to compute arbitrary functions [28, 21, 22, 23, 17] and are, therefore, more general

but at the same time often less efficient than desired. It seems that, so far, there is a lack of finding truly efficient protocols for very fundamental tasks in cryptography. One such example is (password-based) key exchange and is the starting point of our work.

### 1.2 UC-Secure Password-Based Key Exchange

Roughly, in an authenticated key exchange protocol two parties aim to establish a shared key with a designated partner. It is often the case that the parties only hold a human-memorizable password of limited entropy, and still try to establish a strong cryptographic key. A widely-deployed security notion for such password-based authenticated key exchange protocols has been given by Bellare et al. [4]. Later, Canetti et al. [9] (and subsequently [1]) gave stronger formalizations for password-based key exchange in the UC framework. So far, only a few password-based key exchange protocols meeting this notion have been proposed, usually relying on encryption schemes and hash proof systems [10, 1, 20, 25, 29]. An exception is the recently proposed solution by Canetti et al. [8] which relies on oblivious transfer; we discuss this approach in more detail later.

**An Efficient and Strongly Secure Construction.** In line with previous efforts, our protocol for authenticated password-based key exchange achieves UC security. Remarkably, we use only the weak one-time memory tokens (for strings) and, in contrast to previous efforts, our protocols are information-theoretically secure against adaptive corruptions (if the sender can reliably erase randomness). That is, we formally show UC security for unbounded adaptive adversaries and unbounded environments in the hybrid model where we model one-time memory tokens through ideal functionalities. In particular, we do not rely on any unproven cryptographic assumption. According to a recent result by Unruh [43] our protocols thus also resist quantum adversaries. In addition, from a result by Nielsen [36] about non-committing encryption, it follows there cannot exist key agreement protocols secure against adaptive corruptions in the UC model, in the standard model and assuming a full state reveal (and even if the key agreement protocol is only computationally secure). Otherwise, the key derived in such a protocol could be used to encrypt messages in a non-committing way. Our protocol overcomes this impossibility result by assuming reliable erasures.

Note that a UC secure key exchange protocol would also follow from the general feasibility result in [23], showing that one can realize any functionality with a non-interactive, statistically UC-secure protocol based on one-time (bit) memory tokens. What distinguishes our solution from this general result based on circuit evaluation is that, while providing the same security, our protocol only requires very basic operations, and for $\ell$-bit passwords requires only $\ell$ tokens, sent in the first step from the sender to the receiver. The overall protocol requires four moves and $\mathcal{O}(\ell\lambda)$ bits communication for security parameter $\lambda$, determining the statistical security.

At a superficial glance, it may seem easy to design key exchange protocols when one-time memory tokens are available. For instance, simple approaches like having the sender encode random pairs in $\ell$ tokens, asking the receiver to read the $i$-th token according to the $i$-th password bit and to send parts of the value back to authenticate and use the other part to compute the key, are imaginable. This construction, however, does not withstand very basic attacks. Namely, a malicious sender could prepare such tokens and compute the password from the receiver's answer. In fact, the main obstacle in designing UC secure protocols is to solve the "chicken-and-egg" problem of authentication, i.e., one party needs to authenticate first without revealing anything useful to the adversary about the password. Our solution shows that this is still possible with the approach of having the receiver read out the $i$-th of the $\ell$ tokens according to the $i$-th password bit. But we let the sender first authenticate by having the receiver send a random challenge vector after having received the tokens, with the guarantee that the sender's subsequent authentication reply looks random if the receiver reads out a single token at the "wrong" position. In the fourth round of our protocol the receiver then authenticates.

**Authenticated vs. Unauthenticated Token Transmissions.** Another distinct feature is our treatment of reliable token transmissions. To this end, we present our protocol in two steps. First, we discuss a version in the authenticated setting in which both the tokens and the protocol messages are transmitted in an authenticated way, i.e., the adversary can neither change the tokens nor modify any messages, but is still able to read tokens (and

thus inactivate one-time memory tokens). We then discuss that this protocol is insecure if the adversary can also substitute tokens, and thus show how to secure our protocol against these more sophisticated attacks. The common key-confirmation solution to thwart such attacks in key exchange protocols is to authenticate the transcript and then refresh the key. However, since it is unclear how to do this securely in the information-theoretic setting, we use a tailor-made solution which authenticates only one of the message transmissions via an information-theoretically secure one-time MAC.

The reason for presenting our protocol in two steps is to (a) simplify the proof by splitting it, and more importantly, (b) to point out the importance of considering unauthenticated transmission of tokens. Previous efforts often assumed that tokens are eventually faithfully delivered to the intended receiver, and that the adversary could not intercept *and* replace tokens *on behalf of the sender*. This is testified by the common definition of token functionalities (e.g., [28, 21, 12, 23]) in which the receiver is informed by the functionality about a token creation from the sender. In particular, the adversary thus cannot send a different token on behalf of the sender. This is even the case for the tokens considered in [12] where the adversary can basically place tokens inside other tokens, but creation of such tokens again reveals the sender's identity to the receiver. Similarly, in the setting discussed, for example, in [5], the adversary may only access the physically uncloneable function (PUF) token temporarily, before the receiver eventually holds the token, but the adversary again cannot replace the token. Finally, the works by Kolesnikov [30], Fischlin et al. [18] and Ostrovsky et al. [37] consider the dual problem of untrustworthy tokens itself, but not trustworthy transmission of tokens.

It is of course desirable to take substitution attacks into account in order to thwart man-in-the-middle attacks. Indeed, as our key exchange example shows, protocols may become insecure if the adversary is able to carry out such attacks. We therefore present both an ideal functionality for one-time memory tokens which are transmitted authentically and one for unauthentically transferred tokens. The authenticated version serves as a base for our simple protocol, and we use the unauthenticated version for our stronger protocol.

**On Efficient Simulators.** Both in our basic as well in our advanced protocol the simulator for showing UC security runs in polynomial time in the running time of the adversary and the *size of the password space*. Hence, because human-memorizable passwords are typically short and can be considered to be of logarithmic length, this provides a simulator whose complexity is polynomially related to the adversary's running time and the security parameter. Recall that the goal of the paper is indeed to derive *password-based* protocols.

Still, if one, nonetheless, considers passwords of super-logarithmic length, such as full-fledged cryptographic keys, the running time of our simulator becomes exponential. Whether this now provides the desired level of security or not depends on the definition of statistical (UC) security: If one considers statistical security with respect to the class unbounded algorithms, in the sense that for any potentially unbounded adversary there is a potentially unbounded simulator, then our simulator achieves this goal. If, on the other hand, one requires that the simulator's complexity is also polynomially related to the adversary's running time, then we no longer attain this security level. See the discussions in [6, 7] for more details on the two approaches. Here, we use the more liberal definition and state the runtime of the simulator in terms of the running time of the adversary and the security parameter in the theorems, allowing to impose either definition. We note that the idea of using strong, super-polynomial time simulators has been applied in the UC framework before, usually to overcome replace set-up assumptions [38, 40, 3, 32, 31]. We are not aware if our protocol can be modified to yield one, which is also secure according to the stronger notion.

### 1.3 Discussion

To support the argument of the difficulty to design secure key-exchange protocols based on one-time memory tokens note that, as mentioned before, such tokens provide a weaker version of oblivious transfer. The important difference, as pointed out in [21], is that a malicious receiver may decide to postpone the read-out of the token, whereas the sender in an oblivious transfer protocol can be assured about the point in time when the receiver obtained one of the two values. In analogy to such "forced read-outs" we note that such delayed measurements (in case of quantum

states) are the reason that the common security definition for the BB84 quantum key exchange protocol turned out to be inadequate when composed with the one-time pad encryption [41].

Additionally, referring to the case of authentic vs. unauthentic transmissions, most oblivious transfer protocols today are assumed to be run over an authenticated channel and merely provide computational security [19, 24, 14, 39, 42].[1] Furthermore, in order to be UC secure they also need some kind of set-up assumption like a common reference string [7]. It is currently unclear if one can implement such set-up assumptions if the tokens are not sent in an authenticated way. Hence, finding *very efficient* password-based key exchange protocols using these tokens is at least as hard as designing such protocol based on a (UC secure) oblivious transfer protocol. Moreover, to match the properties of our solution this should be true for unauthenticated channels, hold information-theoretically, and thwart adaptive corruptions (although reliable erasures are allowed)! We note that in a concurrent work, Canetti et al. [8] indeed present efficient UC-secure password-based key exchange protocols based on $\ell$ oblivious transfers. However, the security of their construction is based on computational assumptions and the CRS model. In light of the above discussion, our solution also relies on the weaker assumption of one-time memory tokens.

One-time memory tokens are advantageous due to their simple functionality. However, once the token has been queried, the token is useless. Ideally, one would like to have such tokens which implement a reusable version, but then usually information-theoretic security cannot be achieved anymore. If one settles for computational security, then the token can either be stateful, i.e., change its state after queries, or stateless. Clearly, the latter tokens are again preferred as they are, for example, resettably secure [12, 30, 23]. Our protocols can be easily implemented in a stateful version, as the values stored in the memory are purely random. Hence, one can easily use a pseudorandom function to let the token return (pseudo)random answers and update the key of the pseudorandom function afterwards. It remains open if one can implement a version using a stateless token, without involving additional assumptions like an oblivious transfer protocol. Another alternative, to reduce the number of tokens required per execution, is to use the (statistically-secure) sequential-OTM functionality derived from a single wrapper token [17]. While the solution requires additional interaction it is quite efficient; one would need to check, though, that the construction remains secure for unauthenticated transmissions.

Finally, we remark that there seems to be an obvious relation of one-time memory protocols to quantum cryptography. Like a one-time memory token a quantum state collapses when measured. There are, nonetheless, important differences. First note that one-time memory tokens allow "secure local" measurements, whereas a quantum adversary could entangle quantum states and any "local" measurement, e.g., based on the password, would immediately become available to the adversary. Indeed, the impossibility of unconditionally secure 1-out-of-2 oblivious transfer in the quantum setting is well established today [33, 34].

# 2 Security Model

## 2.1 Universal Composition Model

We analyze our password-based key exchange protocol in the Universal Composition Framework by Canetti [7]. In the UC framework one compares *real-world* executions of a protocol $\pi$ among (possibly corrupt) parties $P_1, \ldots, P_n$ in the presence of an adversary $\mathcal{A}$ to *ideal-world* executions in the presence of an ideal-model adversary Sim, usually called simulator. In the ideal-world, the parties have secure access to an ideal functionality $\mathcal{F}$ which captures the desired functional and security properties in an abstract way; in this case, the parties essentially only forward any input to the functionality and are thus called dummy parties. As usual, both adversaries may corrupt parties at the beginning only (static corruptions) or during protocol executions (adaptive corruptions). In the latter case, one often assumes that honest parties are able to erase unnecessary internal data reliably such that these data are not available to the adversary upon corruption. For technical reasons each execution is accompanied by a globally unique session identifier sid available to all parties.

In both worlds a special party, called the environment $\mathcal{Z}$, provides the input to the parties, reads their outputs, and communicates with the adversary. Roughly, the environment models an active distinguisher which determines

[1] It must be said though that, except for [19, 8], none of the papers is explicit about authenticated channels; we draw this conclusion by consultation of the proofs.

any setting in which the protocol may be executed. Ideally, the environment should not be able to tell apart real-world executions from ideal-model executions. That is, let $\text{REAL}_{\mathcal{A},\mathcal{Z},\pi}(\lambda)$ be the output of the environment when interacting with a real-world execution of protocol $\pi$ in the presence of $\mathcal{A}$ for security parameter $\lambda$, and $\text{IDEAL}_{\text{SIM},\mathcal{Z},\mathcal{F}}(\lambda)$ denote $\mathcal{Z}$'s output for the ideal world. Then, a protocol $\pi$ securely realizes $\mathcal{F}$ (information-theoretically) if for any efficient (resp. unbounded) adversary $\mathcal{A}$ there should be an efficient (resp. unbounded) simulator SIM such that for any efficient (resp. unbounded) environment $\mathcal{Z}$ we have

$$\text{REAL}_{\mathcal{A},\mathcal{Z},\pi}(\lambda) \approx \text{IDEAL}_{\text{SIM},\mathcal{Z},\mathcal{F}}(\lambda),$$

where $\approx$ denotes computationally indistinguishability (resp. statistical indistinguishability).[2]

To model hardware tokens we follow previous approaches and specify tokens through ideal functionalities as well, and work in the *hybrid world*. In this real-world surrogate all parties $P_1, \ldots, P_n$ and $\mathcal{A}$ now also have access to an ideal functionality $\mathcal{G}$, and we write $\text{HYBRID}_{\mathcal{A},\mathcal{Z},\pi^{\mathcal{G}}}(\lambda)$ for the environment's output. We say that $\pi$ *securely realizes $\mathcal{F}$ in the $\mathcal{G}$-hybrid world (information-theoretically)* if for any efficient (resp. unbounded) adversary $\mathcal{A}$ there should be an efficient (resp. unbounded) simulator SIM such that for any efficient (resp. unbounded) environment $\mathcal{Z}$ we have

$$\text{HYBRID}_{\mathcal{A},\mathcal{Z},\pi^{\mathcal{G}}}(\lambda) \approx \text{IDEAL}_{\text{SIM},\mathcal{Z},\mathcal{F}}(\lambda).$$

One may again assume here that the simulator is at most polynomial time in the adversary's running time; we do not impose this restriction.

### 2.2 Password-based Key Exchange

We next revisit the definition of an ideal functionality $\mathcal{F}_{\text{pwKE}}$ of Canetti et al. [9] in Figure 1. The ideal functionality captures the abstract properties of a key exchange protocol in the sense that two parties agree upon the same cryptographic key only if they hold the same passwords. Furthermore, by definition, the adversary can test only a single (explicit) password for correctness in each execution; for a wrong guess, the execution is interrupted. Note that the definition does not make any claims about the additional property of contributiveness, in the sense that both parties contribute to the key (as defined in [2]).

For more motivational discussion about the definition see [9], where it is shown that any protocol realizing this definition also provides security according to the common Bellare-Pointcheval-Rogaway notion for password-based authenticated key exchange [4] (except for technical issues regarding session identifiers). As pointed out in [1] any protocol meeting the definition immediately provides forward security as well, since the environment chooses passwords. The latter also makes the deployment of the joint-state UC framework [11] obsolete.

### 2.3 One-Time Memory Tokens

As explained above we model the hardware token as an ideal functionality and work in the corresponding hybrid model. We model sending of tokens in the protocol by having a party create the token and transmitting the unique token identifier tid (which corresponds to the session identifier of the token hybrid in the UC model) to the partner. We assume that each tid is of the form $\text{tid} = (\text{sid}, \text{tid}')$ for the session identifier sid for which the token is created for.[3] Hence, when considering authenticated transmissions of tokens, i.e., sending the tid over an authenticated channel, the adversary cannot change the identifier tid sent from an honest sender. Nonetheless, the adversary can still access the token via calling the ideal functionality for the tid. In previous definitions authentic token transfers have been ensured by letting the functionality inform the receiver about a newly created token by the sender.[4]

[2] Here, one may further assume that the simulator is at most polynomial time in the adversary's running time; as remarked in the introduction we stick to the more liberal definition and revisit this issue when stating the theorem.

[3] In a stronger model one could demand that tokens are not even session-specific. But then the benefits of using the UC model, analyzing a single protocol execution and concluding composition immediately, would be lost. Instead one would need to manually deal with multiple dependent executions, similar to the joint-state UC framework [11].

[4] Interestingly, most definitions (e.g., [12, 23]) neither allow calls of the adversary to the token created for the honest receiver and thus, formally, even exclude temporary access through the adversary when the token is in transmission.

**Functionality $\mathcal{F}_{\text{pwKE}}$**

The functionality $\mathcal{F}_{\text{pwKE}}$ is parameterized by the security parameter $\lambda$. The interaction with an adversary SIM and a set of parties is enabled via the following queries:

- On input $(\texttt{NewSession}, \text{sid}, P_i, P_j, \text{pw}, \text{role})$ from party $P_i$, send $(\texttt{NewSession}, \text{sid}, P_i, P_j, \text{role})$ to SIM. Furthermore, in case this is the first `NewSession` query, or the second query while there exists a record $(\text{sid}, P_j, P_i, \text{pw}')$, then record $(\text{sid}, P_i, P_j, \text{pw})$ and mark this record *fresh*.
- On input $(\texttt{TestPwd}, \text{sid}, P_i, \text{pw}')$ from the adversary SIM, check for a fresh record of the form $(\text{sid}, P_i, P_j, \text{pw})$. If found and
  - $\text{pw} = \text{pw}'$, mark the record as *compromised* and send "correct guess" back to SIM.
  - $\text{pw} \neq \text{pw}'$, mark the record as *interrupted* and send "wrong guess" back to SIM.
- On input $(\texttt{NewKey}, \text{sid}, P_i, sk)$ from SIM and $sk$ is an element in the session key space, check for a record of the form $(\text{sid}, P_i, P_j, \text{pw})$ and that this is the first `NewKey` query for $P_i$. If so, then
  - If this record is *compromised*, or either $P_i$ or $P_j$ is corrupted, then output $(\texttt{NewKey}, \text{sid}, sk)$ to party $P_i$.
  - If this record is *fresh*, and there is a record $(\text{sid}, P_j, P_i, \text{pw}')$ with $\text{pw}' = \text{pw}$, and a key $sk'$ was sent to $P_j$ where $(\text{sid}, P_j, P_i, \text{pw}')$ was *fresh* at the time, then output $(\texttt{NewKey}, \text{sid}, sk')$ to $P_i$.
  - Otherwise, pick a new random key $sk'$ in the key space and send $(\texttt{NewKey}, \text{sid}, sk')$ to $P_i$.

  In any case, this record is marked as *completed*.

Figure 1: Ideal functionality $\mathcal{F}_{\text{pwKE}}$ for password-based key exchange

To consider unauthenticated transmissions of tokens, we let the adversary intercept tokens and read them, even when the designated receiver is honest. This requires the adversary also to create tokens (within the session sid) and to replace the depleted token. We note that we do not need to make any assumption about who currently is in possession of the token (i.e., that the adversary cannot read the token if it has already been delivered). The reason is that we can assume that, if the adversary forgoes reading the token when sent and instead delivers it to the receiver who then reads the token and ultimately makes it inaccessible for the adversary. Figure 2 illustrates the ideal functionalities of One-Time-Memories in both versions.

In the sequel, we often merely say that a party creates a token. It is understood that the party calls the functionality in the hybrid model with the corresponding input command `create`. The party sends the token to another party if it later transmits the token identifier to this party. The receiver is assumed to check that the identifier matches the session identifier. We then say that the receiver queries or accesses the token if it sends a `choice` command.

We note that the token identifier tid is immediately available to the adversary, once the token has been created by an honest sender. She could thus access the token before it is even delivered, i.e., before tid is sent to the intended receiver. However, since we can assume that the honest sender never accesses its own token (because it already knows the encapsulated values) and that tid may be sent by the protocol, this does not violate generality. Analogously, the protocol can easily enforce that an intended honest receiver $P_j$ does not access the token before actually receiving tid.

# 3 Key Exchange in the Authenticated Setting

In this section, we propose an information-theoretically secure password-based authenticated key-exchange protocol using one-time memory tokens. As explained in the introduction we first assume that all transmissions are authenticated, including the token handover. In other words, in each execution the adversary may eavesdrop on the communication of honest parties, or impersonate either side if the corresponding party has already been corrupted.

**Functionality $\mathcal{F}_{\text{OTM}}^{auth}$**
(authenticated transmission)

The functionality $\mathcal{F}_{\text{OTM}}^{auth}$ is parameterized by the security parameter $\lambda$. The interaction with an adversary and a set of parties is enabled via the following queries:

- On input $(\texttt{create}, tid, P_i, P_j, (s_0, s_1))$ from party $P_i$, store $(s_0, s_1)$. Go to state **ready** and send the publicly delayed output $(\textbf{ready}, tid, P_i, P_j)$ to party $P_j$.
- On input $(\texttt{choice}, tid, P_i, P_j, x)$ from party $P_j$ or the adversary and $x \in \{0,1\}$, check if current state is **ready**. If so, send $(\textbf{out}, tid, P_i, P_j, s_x)$ to $P_j$ resp. the adversary and go to state **dead**.

**Functionality $\mathcal{F}_{\text{OTM}}^{unauth}$**
(unauthenticated transmission)

The functionality $\mathcal{F}_{\text{OTM}}^{unauth}$ is parameterized by the security parameter $\lambda$. The interaction with an adversary and a set of parties is enabled via the following queries:

- On input $(\texttt{create}, tid, P_i, P_j, (s_0, s_1))$ from party $P_i$ or the adversary store $(s_0, s_1)$. Go to state **ready** and send the publicly delayed output $(\textbf{ready}, tid, P_i, P_j)$ to party $P_j$.
- On input $(\texttt{choice}, tid, P_i, P_j, x)$ from party $P_j$ or the adversary and $x \in \{0,1\}$, check if current state is **ready**. If so, send $(\textbf{out}, tid, P_i, P_j, s_x)$ to $P_j$ resp. the adversary and go to state **dead**.

Figure 2: Ideal functionalities $\mathcal{F}_{\text{OTM}}^{auth}$ and $\mathcal{F}_{\text{OTM}}^{unauth}$ for One-Time-Memory (OTM)

Since the adversary cannot modify tokens nor the communication, we thus explicitly exclude man-in-the-middle attacks, i.e., the adversary can only send a chosen token in an execution if it impersonates the actual sender in that execution.

### 3.1 The Protocol $\mathsf{KE}_{\text{auth}}$

The idea of our protocol $\mathsf{KE}_{\text{auth}}$ is as follows. For each bit $\text{pw}_i$ of the password pw the sender $A$ hands over an OTM token to the receiver $B$ with two tuples $(s_{b,i}^A, s_{b,i}^B, s_{b,i}^K)$ for $b \in \{0,1\}$ of random $s$-bit strings. The receiver fetches the tuple corresponding to bit $\text{pw}_i$ and the other tuple is then erased by the definition of the token's functionality. Next, both parties authenticate themselves using the password by having $A$ compute $\bigoplus s_{\text{pw}_i,i}^A$ resp. having $B$ send $\bigoplus s_{\text{pw}_i,i}^B$. The computation of $A$ actually requires $B$ to first send pairs $(t_{0,i}, t_{1,i})$ of random $s$-bit strings and have $A$ compute $\bigoplus(s_{\text{pw}_i,i}^A \oplus t_{\text{pw}_i,i})$ instead. Otherwise a malicious $A$ could choose the $s_{b,i}^A$'s such that multiple passwords map to the same string, e.g., all $s_{b,i}^A = 0$, could then easily pass the authentication step and eventually find out valuable information about pw through well chosen $s_{b,i}^B$'s from $B$'s authentication step.

Figure 3 illustrates our protocol $\mathsf{KE}_{\text{auth}}$ in more detail. For technical reasons, we let a party send a random answer in case of an inconsistent incoming message in the third step or fourth step, and set the session key to a random string of length $s$. This is necessary because, in case of premature aborts, the UC simulator simulating an execution between honest parties would need to know if the passwords of the two parties match. This information, however, is not available through the ideal functionality $\mathcal{F}_{\text{pwKE}}$.

In fact, the possibility that two honest parties may have distinct passwords, e.g., by mistyping them, also requires the reliable erasure on the sender's side for adaptive corruptions. In case either of the parties gets corrupted, we learn the password of this party, but can only test for equality with the partner's password. That is, we cannot determine the actual password of the partner. However, then the other parties' sent data must still be independent of its password in the simulation, and this is ensured by noting that the sender's internal data lacks the random token inputs for false password bits which the receiver (with a distinct password) has used. For the receiver, this holds by construction, because reading the token at a different position has somewhat erased the other data anyway.

### 3.2 Security

**Theorem 3.1** *Protocol $\mathsf{KE}_{auth}$ securely realizes $\mathcal{F}_{pwKE}$ in the $\mathcal{F}_{OTM}^{auth}$-hybrid world information-theoretically with adaptive corruptions, assuming reliable erasures and authenticated channels. This holds as long as $s - 2\ell = \Omega(\lambda)$*

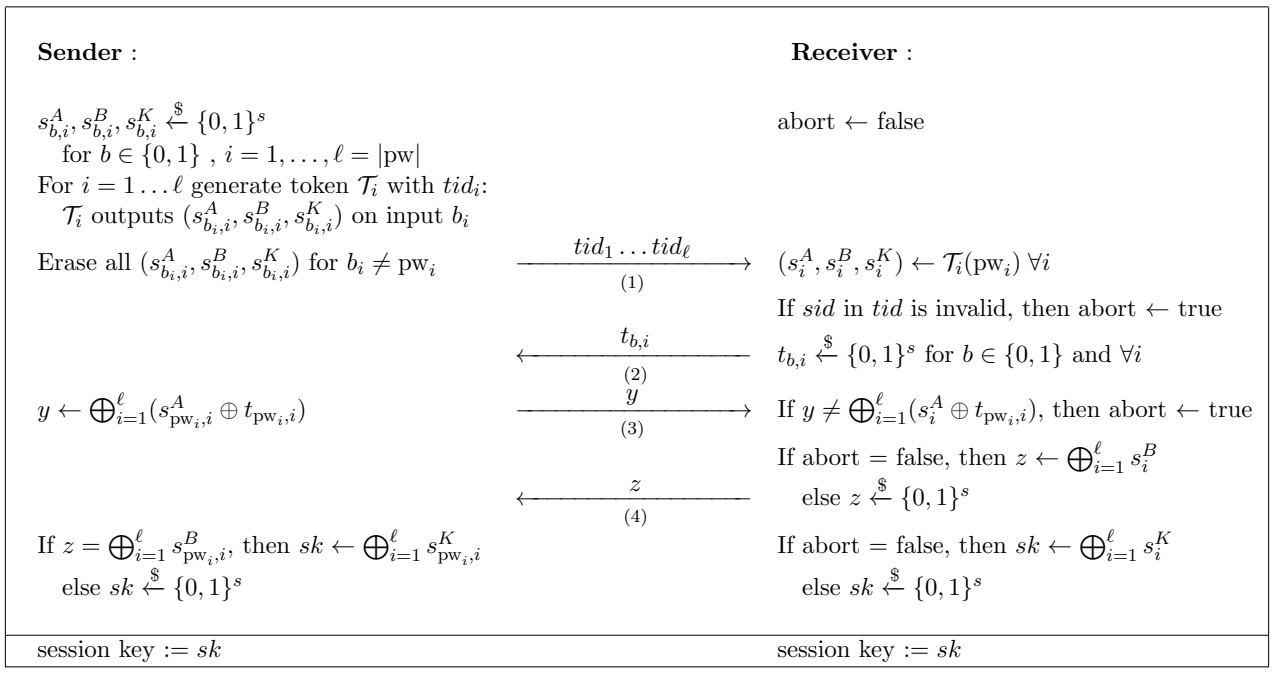



Figure 3: Password-based authenticated key-exchange protocol $\mathsf{KE}_{\mathrm{auth}}$ using OTM tokens

*for security parameter $\lambda$. The simulator's running time is polynomial in the adversary's running time, the security parameter, and $2^\ell$.*

*Proof.* Fix an arbitrary (unbounded) adversary $\mathcal{A}$ for the real-world protocol $\mathsf{KE}_{\mathrm{auth}}$. We describe an (unbounded) ideal-model simulator Sim running essentially a black-box simulation of $\mathcal{A}$ as a sub routine by feigning an execution in the real-world, using the data it gets in the ideal world. The simulator relays any communication between the adversarial sub routine $\mathcal{A}$ and the environment $\mathcal{Z}$, and since we consider adaptive corruptions, the black-box simulation therefore boils down to provide replies on behalf of (simulated) honest parties and to be able to provide a valid-looking internal state in case of a corruption. In what follows, when we say that an honest party or the simulator "rejects", we mean that the honest party continues running the protocol but sends random values instead of honest responses, ultimately setting the session key to an $s$-bit random value.

We note that in the simulated hybrid world the simulator has full control over the token functionality and simulates it internally. In particular, it can read any data which the adversary submits to this functionality and can prepare replies as required. We especially make use of the fact that the simulator does not need to explicitly query the token on behalf of an honest receiver (but the simulator then still needs to ignore further adversarial read requests because ideal token would do so as well).

**Description of the Simulator.** Next, we describe the simulator in detail. At any point in the simulation we consider the point in time in which a corrupt receiver in an execution or the adversary makes the query to the tokens for the last remaining bit position. We then say that the tokens are *depleted*, otherwise the tokens are called *unexhausted*. In case of depletion the simulator possibly has to adapt the (last) token's answer according to previous steps. Note that simulated honest receivers in the hybrid model can be assumed not to query the tokens at all.

**Tokens are depleted.** If at any point during the simulation the tokens are depleted by a query about bit position $i$, then the simulator adapts the token's final answer as follows. We note that the simulator already holds candidate values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$ for this token, either because a corrupt sender has put them into the token (and which Sim can simply read off in the hybrid model), or because the simulator has chosen them on behalf of the honest sender (and possibly adapted them before, when the sender or receiver got corrupted).

The simulator first determines the password candidate as $\overline{\text{pw}} := \text{pw}'_1 || \ldots || \text{pw}'_\ell$ for the adversary's queries to the token, including the final one $\text{pw}'_i$, and then —if the sender is still honest— tests it by calling $\mathcal{F}_{\text{pwKE}}$ with input $(\texttt{TestPwd}, \text{sid}, P_i, \overline{\text{pw}})$ where $P_i$ is the sender's identity.

- If the test returns "correct guess", then SIM stores this password $\text{pw} = \overline{\text{pw}}$ and $sk = \bigoplus_{i=1}^{\ell} s^K_{\overline{\text{pw}}_i,i}$ for the values in the token for further use. In particular, if the simulated sender later receives a correct value $z$ then set the session key of the sender to $sk$ via a `NewKey` query about $(\text{sid}, P_i, sk)$.

  If the simulator, at that point, has already sent $y$ in this simulated execution on behalf of an honest sender, then SIM resets the entry $s^A_{\text{pw}_i,i}$ to

$$s^A_{\text{pw}_i,i} = y \oplus \bigoplus_{j=1, j\neq i}^{\ell} (s^A_{\text{pw}_j,j} \oplus t_{\text{pw}_j,j})$$

  such that $y$ matches the sender's answer it would give for the password.

- Suppose that the test returns "wrong guess". Then the simulator not change anything at this point. If not done already, it will later send an independent random value $y$, and reject any answer $z$ of the receiver.

In either case return the (possibly updated) values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$.

**Sender Corruption.** Suppose that the adversary corrupts the sender in a simulated execution. Then the simulator corrupts the sender in the ideal world and learns the input password pw. For any execution with unexhausted tokens in which SIM has already sent $y$, adapt the value $s^A_{\text{pw}_i,i}$ for unqueried position $i$ as in the case of token depletion. For an honest receiver test the password under the receiver's identity and possibly adapt the value $s^B_{\text{pw}_i,i}$ accordingly, if $z$ has already been sent.

If $y$ has not been sent yet, or the tokens are already depleted, the values remain unchanged. Hand over the password and the internal data (excluding the token inputs for the inverse bit values of $\text{pw}_i$, which have been erased before).

**Receiver Corruption.** As in the case of sender corruption SIM then learns the actual input password. Test this password also for the sender (if it is still honest). If $y$ has already been sent in an execution with a sender who is still honest, and the password test with the sender returned "correct guess", then adapt the values $s^A_{b,i}$ in $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$ allegedly in the token accordingly, else the values remain unchanged. If, in addition, the simulated receiver has sent $z$ already, then also adapt the values $s^B_{b,i}$. Besides the password, present the adversary the $t_{b,i}$'s (if chosen already) and the entries $(s^A_{\text{pw}_i,i}, s^B_{\text{pw}_i,i}, s^K_{\text{pw}_i,i})$.

**Simulating Steps of Honest Parties.** We show next how to simulate the other steps on behalf of honest parties.

Simulating Step (1): For an honest sender the simulator picks values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$ at random and puts them into a token.

Simulating Step (2): If the receiver is honest, then the simulator picks the values $t_{b,i}$ at random but does not query the token. It sends the $t_{b,i}$'s on behalf of the receiver in the simulation.

Simulating Step (3): If the tokens are already depleted and the test query returned "correct guess", then the simulator computes $y$ genuinely according to the (possibly adapted) values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$ and the $t_{b,i}$'s. If the test returned "wrong guess" or the tokens are unexhausted at this point, then send a random $y$ instead.

Simulating Step (4): In this step the receiver is supposed to verify the sender's value $y$ and to provide its own authentication value $z$.

Assume that the sender is controlled by the adversary when sending $y$. Then SIM already knows all values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$ stored in the token (either because it put them there on behalf of an honest sender before being corrupted, or since the adversary put them into the token right away). The simulator now tries all possible

passwords pw, in polynomial time $2^\ell = \text{poly}(\lambda)$, and for each password computes $y_{\text{pw}}$ as the honest sender would (given the $s_{b,i}$'s and $t_{b,i}$'s). If none value matches $y$, then SIM lets the receiver abort; if more than two values match, then SIM aborts the simulation with an error message. Else, the simulator makes a `TestPwd` query about the (only) candidate password pw for sid and the receiver's identity $P_j$. If this query returns "correct guess", then compute the final message $z$ and the key $sk$ according to this password and make a call $(\texttt{NewKey}, \text{sid}, P_j, sk)$. For the reply "wrong guess" let the receiver abort the execution, i.e., use random data from this point on.

For an honest sender the receiver accepts and sends a random answer $z$. The simulator calls the functionality about $(\texttt{NewKey}, \text{sid}, P_j, sk_0)$ for an arbitrary key $sk_0$ in the key space.

Simulating the Sender's Final Step: When the honest sender $P_i$ obtains the final message $z$ from an honest receiver, it accepts, and the simulator calls the functionality about $(\texttt{NewKey}, \text{sid}, P_i, sk_0)$ for an arbitrary key $sk_0$ in the key space. If the receiver is corrupt at this point and the tokens are unexhausted, then SIM lets the sender reject. If the tokens of the corrupt receiver are depleted at this point, then the simulator has already tested a password. If this test returned "correct guess", then the simulated sender uses the guessed password to compute the final answer and creates the output via $(\texttt{NewKey}, \text{sid}, P_i, sk)$ for the previously derived key. Else, for "wrong guess", the simulator lets the sender reject.

**Analysis.** We show next that the above simulation through SIM is indistinguishable from a real-world execution, i.e., any (computationally unbounded) environment is unable to differentiate between the protocol execution built up by SIM or a real execution influenced by the adversary $\mathcal{A}$.

We first observe that unexhausted tokens leave the simulator enough freedom to adapt values perfectly, i.e., the distributions of computing the correct value $y$ from random $s^A_{b,i}$ and $t_{b,i}$, and from picking $y$ at random and later setting the "undetermined" value $s^A_{b,j}$ accordingly, are identical. The same is true for the receiver's values. This holds, in particular, in case of corruption of a party where testing the password on the side of the honest partner either yields a correct guess (in which case the values are adapted correctly), or a wrong guess in which case the partner has a distinct password and derives independent random strings. The latter is perfectly indistinguishable if the receiver gets corrupted, because the receiver reads out at least one token at a different position such that the other random string originally stored in the token is not available to the receiver (and eventually to the adversary) anymore. It also holds in case of a sender corruption because then the sender has already erased reliably the other data.

It remains to argue that the simulator's verifications of the values $y$ and $z$ are sound. For two honest parties, since we assume immutable transmissions, and because the ideal functionality takes care of deciding equality of the passwords of the two parties, calling `NewKey` as described provides an almost perfect simulation.

For a malicious partner in the simulation, the difference to the real-world protocol is $y$ could match two possible passwords such that the simulator aborts, but the genuine receiver would accept (in the case that $y$ does not match any of the $y_{\text{pw}}$ of possible passwords the receiver in the real-world protocol would also reject). We next bound the probability of two passwords $\text{pw} \neq \text{pw}'$ yielding the same $y_{\text{pw}} = y_{\text{pw}'}$. Note that the values $s^A_{b,i}$ are chosen and put into the token before the $t_{b,i}$'s are sent (either by the simulator or the adversary). Furthermore, the receiver is still honest, such that the $t_{b,i}$'s are random, mutually independent, and independent of the $s^A_{b,i}$. We conclude that the probability (over the $t_{b,i}$'s) that

$$y_{\text{pw}} = \bigoplus_{i=1}^{\ell} (s_{\text{pw}_i,i} \oplus t_{\text{pw}_i,i}) = \bigoplus_{i=1}^{\ell} (s_{\text{pw}'_i,i} \oplus t_{\text{pw}'_i,i}) = y_{\text{pw}'}$$

for fixed $\text{pw} \neq \text{pw}'$, is at most $2^{-s}$. Hence, with probability at most $2^{2\ell-s-1}$ there exist some $\text{pw} \neq \text{pw}'$ for which a collision occurs. By the choice of parameters this is exponentially small.

Finally, we note that in the case the tokens are unexhausted, or are depleted but the simulator obtained the reply "wrong guess" for the tested password, the probability that $z$ sent by the adversary would pass the verification step in the real-world protocol, is at most $2^{\ell-s}$ (implying that the simulator's strategy of rejecting is valid except for an exponentially small error). This can be seen as follows. In case the tokens are unexhausted the adversary

has not obtained the independent random value $s^B_{b,j}$ for some $b, j$, such that, for a fixed password, the value $z$ can be correct with probability at most $2^{-s}$ only. This is also true in case the tokens are depleted but the password was wrong, because then one of the correct values $s^B_{b,i}$ is irrevocably lost for the adversary. Summing over all $2^\ell$ passwords yields the claim. □

We note again that the simulator above is polynomial-time in the running time of the adversary and of the number of possible passwords. For passwords of logarithmic length this number is polynomial in the security parameter.

# 4 Key Exchange in the Unauthenticated Setting

We have shown so far that information-theoretically secure authenticated key exchange is possible using tamper-proof hardware tokens. However, the security relies in the assumptions that the adversary does not tamper with the tokens as man-in-the-middle. Here we show that the previous protocol succumbs to such attacks and then present a method to secure it in the unauthenticated setting.

## 4.1 Attacking our Protocol in the Unauthenticated Setting

We have shown that $\mathsf{KE}_{\text{auth}}$ securely realizes $\mathcal{F}_{\text{pwKE}}$ in the $\mathcal{F}^{auth}_{\text{OTM}}$-hybrid setting information-theoretically. In this section, we provide a successful attack scenario against $\mathsf{KE}_{\text{auth}}$ when the adversary is allowed to act as man-in-the-middle accessing the transmitted tokens and creating new hardware tokens for the same session. Thereby, we conclude that $\mathsf{KE}_{\text{auth}}$ do not realize $\mathcal{F}_{\text{pwKE}}$ in the (unauthenticated) $\mathcal{F}^{unauth}_{\text{OTM}}$-hybrid world. Roughly, the problem is that the simulation in the proof of Theorem 3.1 implicitly assumes that, whenever an honest party sends $y$ or $z$ to the honest partner, the simulated honest receiver should accept. This may not be a viable strategy if, as below, the adversary may modify transmissions.

We first describe our attack for simplicity not in the UC framework, but as a protocol attack to recover the password. It is easy to turn this into a successful attack in the UC setting (as described below). The attack works as follows. Let $A$ and $B$ be an honest sender and receiver, respectively. When $A$ sends the prepared tokens to $B$, the adversary $\mathcal{A}$ intercepts these tokens and queries the $i$-th token for some $i \in \{1, \ldots, |\text{pw}|\}$ with input $b = 0$. Adversary $\mathcal{A}$ obtains the values $(\text{sid}, s^A_{0,i}, s^B_{0,i}, s^K_{0,i})$. This token is now exhausted and needs to be replaced by a new token such that $B$ does not notice the modification. This is done by creating a new token $\mathcal{T}_i$ by $\mathcal{A}$ which outputs $(\text{sid}, s^A_{0,i}, s^B_{0,i}, s^K_{0,i})$ for input bit 0, and some independent random values for the other value. All the tokens are now handed over to $B$. Henceforth, $\mathcal{A}$ eavesdrop on the acceptance or rejection reaction of the receiver when the value $y$ is sent. This information is usually conveyed to the adversary for free in the BPR model [4] resp. eventually available through the environment in the UC setting, even though in our protocol both parties first continue the execution. If $y$ is accepted, then $\mathcal{A}$ knows that the $i$-th bit of the password is 0 (with overwhelming probability). Otherwise, the $i$-th password bit is 1. Hence, the adversary learns one bit of the password and by repeating this attack the adversary can recover the entire password after $|\text{pw}|$ executions between $A$ and $B$.

Note that this attack is even possible without any modifications on the tokens. For this $\mathcal{A}$ simply does not touch the tokens but instead modifies the value $t_i = (t_{0,i}, t_{1,i})$. More precisely, $\mathcal{A}$ waits until the receiver sends $t_{b,i}$ values and resets $t_{0,i}$ for some $i \in \{1, \ldots, |\text{pw}|\}$ to an independent random value before delivering all $t_{b,i}$ to $A$. Now, if $y$ sent by $A$ is accepted by $B$, then $\mathcal{A}$ knows that the $i$-bit of the password is 1 with overwhelming probability. Otherwise, rejection implies that the $i$-bit is 0.

We note that (either one of) the above attacks can be easily turned into a successfully distinguishing UC environment. Namely, $\mathcal{Z}$ has the two honest parties use the same password whose first bit is 0 or 1, with probability $\frac{1}{2}$ each. Now it asks the adversary to do the modification as above, and finally checks if both parties agree upon the same key. For an actual protocol execution this will only happen iff the first bit of the password was really 0 and thus with probability $\frac{1}{2}$. In the ideal model, however, both parties would always derive the same key because they hold the same passwords.

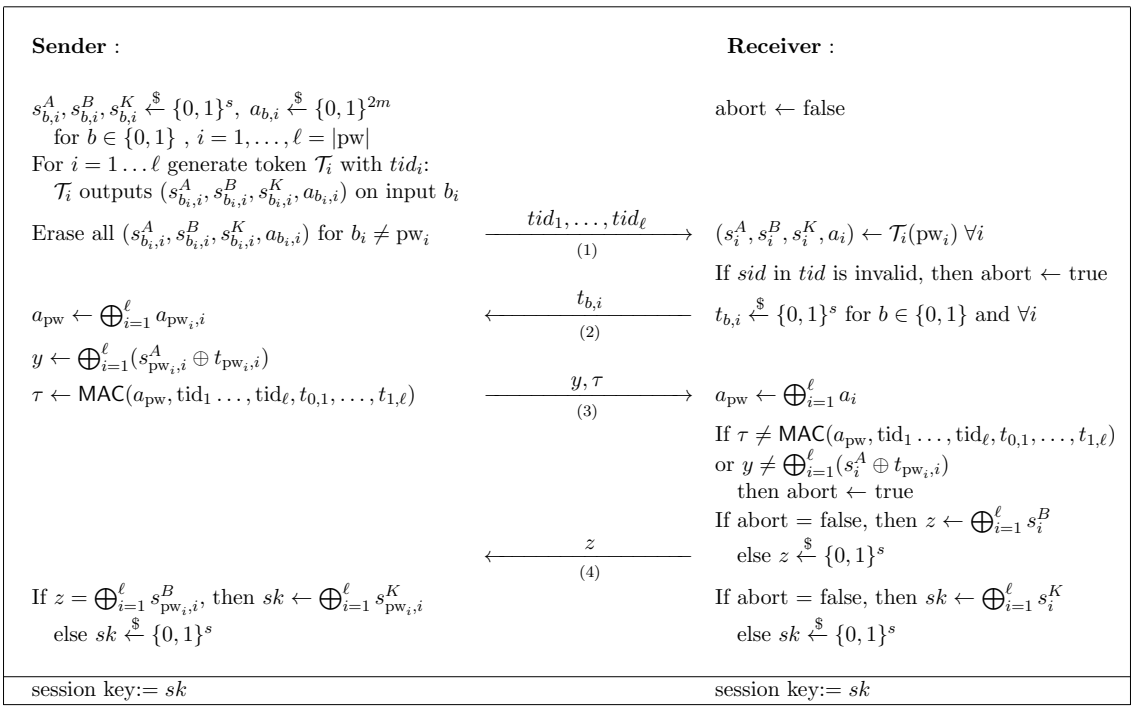



Figure 4: Password-based authenticated key-exchange protocol $\mathsf{KE}_{\mathrm{unauth}}$ using OTM tokens

## 4.2 The Protocol $\mathsf{KE}_{\mathrm{unauth}}$

The idea to make our protocol immune against the above modification attacks is to thwart the adversary from learning individual password bits through one's party reaction. For this we add an authentication step which provides integrity of the tokens and of the $t_{b,i}$ values via the password. It is only necessary to authenticate the transmission from the sender to the receiver, as in case of an invalid authentication code the (honest) receiver will reply with random and thus most likely invalid $z$ anyway.

To authenticate the tokens and the $t_{b,i}$ values we assume an unconditionally secure authentication code like the one of Carter and Wegman [44] which takes an $m$-bit message $M$ (we assume that $m$ is at least the security parameter, else one may pad messages first) and a key $a = (\alpha, \beta) \in \{0,1\}^{2m}$ and (deterministically) computes $\tau = \mathsf{MAC}(a, M) = \alpha M + \beta$ over $\mathrm{GF}(2^m)$. To verify a MAC one recomputes the MAC and compares it to the given $\tau$. Note that, besides providing a (one-time) unforgeability security level of $2^{-m}$ against unbounded adversaries, this message authentication code has another useful property: given a random $m$-bit string $\tau$ and an $m$-bit message $M$ one can easily solve an equation to find a matching key $a = (\alpha, \beta) \in \{0,1\}^{2m}$ with $\tau = \mathsf{MAC}(a, M)$.

We now describe the full protocol $\mathsf{KE}_{\mathrm{unauth}}$. We assume the total length of messages $\mathrm{tid}_1, \ldots, \mathrm{tid}_\ell$ and of the $t_{b,i}$ values is at most $m$. Note that $m = \Omega(s) = \Omega(\lambda)$. As before, for each bit $\mathrm{pw}_i$ of the password pw the sender $A$ hands over an OTM token to the receiver $B$ with two tuples of random $s$-bit strings $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i})$, but this time each tuple is augmented by random $2m$-bit strings $a_{b,i}$ for the MAC. The receiver again reads the tuple corresponding to bit $\mathrm{pw}_i$ such that the other tuple is irrevocably lost. It sends again strings $(t_{0,i}, t_{1,i})$. Next, both parties as before authenticate themselves using the password by having $A$ compute $\bigoplus(s^A_{\mathrm{pw}_i,i} \oplus t_{\mathrm{pw}_i,i})$ resp. having $B$ send $\bigoplus s^B_{\mathrm{pw}_i,i}$. Only this time $A$ also computes $a_{\mathrm{pw}} = \bigoplus a_{\mathrm{pw}_i,i}$ and sends $\tau = \mathsf{MAC}(a_{\mathrm{pw}}, \mathrm{tid}_1, \ldots, \mathrm{tid}_\ell, t_{0,1}, \ldots, t_{1,\ell})$ in addition to $y$. The receiver answers with $z$ as before (or a random value $z$ in case the value $y$ or the value $\tau$ does not verify).

Figure 4 illustrates our protocol $\mathsf{KE}_{\mathrm{unauth}}$ in more detail. Before giving the proof we first discuss why the attack described in the previous section does not work anymore against this protocol. There, the adversary randomly changed a single token tid or value $t_{b,i}$ in order to test for single bit of the password: if the guess was right both parties would continue the execution, else they would reject. Here, any such change would make the MAC over all tid and $t_{i,b}$ invalid and thus make both parties reject any modification. This only option for the adversary is

to compute a new valid MAC. But since the key for the MAC is distributed over all tokens, the adversary would either need to forge a MAC without the key —which is improbable— or need to read *all* tokens to potentially get the key. In the latter case the adversary commits to a password guess which can be used by the simulator for a test and for a corresponding reaction.

### 4.3 Security

**Theorem 4.1** *Protocol $\mathsf{KE}_{unauth}$ securely realizes $\mathcal{F}_{pwKE}$ in the $\mathcal{F}_{OTM}^{unauth}$-hybrid world information-theoretically with adaptive corruptions, assuming reliable erasure. This holds as long as $s - 2\ell = \Omega(\lambda)$. The simulator's running time is polynomial in the adversary's running time, the security parameter, and $2^\ell$.*

*Proof.* We again assume an adaptive (unbounded) adversary $\mathcal{A}$ and construct a corresponding (unbounded) ideal-model simulator Sim via black-box simulation. We note that we can assume that, once the adversary has taken control over a party, that the adversary does not change transmissions from or to that party (over the channel); the adversary could easily make these modifications internally. This, in particular, means that the adversary only modifies transmissions or injects messages in executions between (currently) honest parties.

**Description of the Simulator.** We use the same notion of depleted and unexhausted tokens as in the proof of Theorem 3.1. It is also easy to see that, analogously to there, the simulator can easily adapt the values in case of a corruption. Here, we use the fact that the authentication code allows to adapt the key accordingly.

**Simulation of Honest Parties.** It remains to discuss how to simulate honest parties. Since this requires additional steps, compared to the proof of Theorem 3.1, we discuss this in more detail:

Simulating Step (1): For an honest sender the simulator picks values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i}, a_{b,i})$ at random and puts them into a token.

Simulating Step (2): If the receiver is honest, then the simulator picks the values $t_{b,i}$ at random but does not query the token. It sends the $t_{b,i}$'s on behalf of the receiver in the simulation.

Simulating Step (3): If the tokens are already depleted and the test query returned "correct guess", then the simulator computes $y$ and $\tau$ genuinely according to the (possibly adapted) values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i}, a_{b,i})$ and the $t_{b,i}$'s. If the test returned "wrong guess" or the tokens are unexhausted at this point, then send random $y$ and $\tau$ instead.

Simulating Step (4): In this step the receiver is supposed to verify the sender's values $y, \tau$ and to provide its own authentication value $z$. There are four cases:

1. Assume that the sender is controlled by the adversary when sending $y, \tau$ (in which case the adversary also delivers these values to the honest receiver by assumption). Then Sim already knows all values $(s^A_{b,i}, s^B_{b,i}, s^K_{b,i}, a_{b,i})$ stored in the token (either because it put them there on behalf of an honest sender before being corrupted, or since the adversary put them into the token right away). The simulator now tries all possible passwords pw and for each password computes $y_{\text{pw}}$ as the honest sender would (given the $s_{b,i}$'s and $t_{b,i}$'s). If none value matches $y$, then Sim lets the receiver abort; if more than two values match, then Sim aborts the simulation with an error message. Else, the simulator makes a `TestPwd` query about the (only) candidate password pw for sid and the receiver's identity $P_j$. If this query returns "correct guess", then compute $a_{\text{pw}}$ according to the protocol and verify $\tau$ with this key. If this test also succeeds then compute the final message $z$ and the key $sk$ according to this password and make a call $(\texttt{NewKey}, \text{sid}, P_j, sk)$. If the verification of $\tau$ fails or the reply is "wrong guess" in the password test, let the receiver send random $z$ and abort the execution.

2. For an honest sender, and if the adversary has not tampered the tokens, i.e., delivered the same tid's as the sender has sent in this session, we have the receiver send a random answer $z$. If, in addition, the adversary has delivered the same $t_{b,i}$ values sent by the receiver to the sender in this session, and the same reply $y, \tau$ from the sender to the receiver, then the simulator calls the functionality about $(\texttt{NewKey}, \text{sid}, P_j, sk_0)$ for an arbitrary key $sk_0$ in the key space; else, if the adversary has tampered with the $t_{b,i}$ values or the sender's reply, then the simulator does nothing beyond sending a random value, i.e., lets the receiver abort.
3. If the sender is still honest, but the adversary has sent only new tokens $\text{tid}'_1, \ldots, \text{tid}'_\ell \notin \{\text{tid}_1, \ldots, \text{tid}_\ell\}$, then the simulator determines a password candidate from $y$ as in the case of a corrupt sender and proceeds according to this case (as described above).
4. If the sender is honest, but the adversary has forwarded one of the tokens created by the honest sender but also at least one fresh token (in which case the sender's tokens must be unexhausted), then we let the simulated receiver answer with a random $z$ and internally abort.

Simulating the Sender's Final Step: When the honest sender $P_i$ obtains the final message $z$ from a receiver, it does the following:

1. If the receiver is corrupt at this point and the token is unexhausted, then Sim lets the sender simply abort.
2. If the tokens are depleted at this point, independently of the question whether the receiver is honest or malicious, then the simulator has already tested a password. If this test returned "correct guess", then the simulated sender uses the guessed password to verify the value $z$. If this succeeds, then it creates the output via $(\texttt{NewKey}, \text{sid}, P_i, sk)$ for the previously derived key. Else, if the verification tests fail or the password test has returned "wrong guess", the simulator lets the sender simply abort.
3. If the receiver is still honest and the adversary has simply forwarded all tokens of the sender (in the right order) to the receiver and forwarded also the $t_{b,i}$ values without modifications, and the receiver has not aborted, then the simulator lets the sender accept and calls the functionality about $(\texttt{NewKey}, \text{sid}, P_i, sk_0)$ for an arbitrary key $sk_0$ in the key space, if and only if the adversary faithfully delivers the receiver's final message $z$.
4. In any other case the simulator lets the sender abort.

**Analysis.** We note that the analysis in case of a malicious sender or a malicious receiver is as before. Moreover, if the simulator extracts a password from the adversary and tests it, correctness follows as before. It thus suffices to look at the case of the adversary mounting a man-in-the-middle attack on an execution between honest parties.

First note that, if the adversary simply relays all the data, in particular, without depleting the tokens, then the simulation (basically sending and accepting random values) is perfectly indistinguishable from an actual protocol execution in the real world. This is still true for executions in which the adversary does not tamper with the tokens nor the $t_{b,i}$-values but possibly modifies any of the final two messages. By construction, the corresponding simulated party then aborts, which is what would also happen in the real world (as the messages $y, \tau, z$ are deterministically determined from the first two messages). It thus remains to analyze the case that the adversary somehow modifies any of the two first messages.

Assume that the adversary changes one of the tokens (or the order) but still passes one of sender's tokens to the receiver, or faithfully relays the tokens but alters the $t_{b,i}$ values. Then the adversary is completely oblivious about the secret(s) on the sender's and on the receiver's side. (Note that if both parties use distinct passwords, then they also hold distinct secrets). It follows that the adversary cannot make the receiver accept the subsequent message $y, \tau$, except with exponentially small probability, because it would need to forge a MAC for the distinct series of token identifiers or $t_{b,i}$ values (in case of identical secrets on both sides the adversary would need to create another MAC for a new message, in case of distinct passwords on the sender's and on the receiver's side, the adversary

would need to forge a MAC from scratch). Hence, our simulation in which the receiver then aborts and sends a random answer, is statistically close to an actual protocol execution.

Note that letting the receiver abort and send random values implies that the honest sender then, too, most likely reacts correctly. That is, the only case in which the sender in a protocol execution should *not* reject is when either the adversary by chance supplies correct values to the sender without having corrupted the receiver (which then happens with negligible probability only by the unforgeability of the MAC), or if the adversary corrupts the receiver after the receiver has sent its final random values, and then replaces the random values by correct values with the help of the now known password. In this case, however, the simulator tests the password candidate and rejects if the guess is correct, but the verification fails (as the actual sender would), or if the guess is incorrect and verification would, then also fail with overwhelming probability in the real world.

Hence, assume now that the adversary creates new tokens $\mathrm{tid}'_1, \ldots, \mathrm{tid}'_\ell$, all distinct from the tokens created by the sender. Then we can think of this as essentially two executions where the adversary in one plays the malicious receiver, and in the other one the malicious sender (but where, depending on the adversary's actions, the data may not be independent). In the latter execution, the adversary-receiver execution, the incoming message $y, \tau$ to the receiver allows to determine a password candidate and then proceed as in the real-world case, i.e., the simulation is statistically close. Analogously, in the sender-adversary execution we then act like the real sender would, except that we let the sender immediately reject if the tokens are unexhausted (whereas in the real-world there is a negligible probability that the data $z$ still verifies).

In summary, for any possible adversarial strategy the simulation is statistically close to real-world executions. This proves security. □

# 5 Conclusion

We view the contribution of this paper as three-fold. First, we raise the issue that, previously, token-based protocols neglected man-in-the-middle adversaries which can intercept and substitute tokens while in transmission. Secondly, we show that one can still design secure protocols based on simple tokens, even when allowing such strong adversaries, and what techniques may be useful. Thirdly, complementing previous feasibility results, we exemplified for password-based key exchange that such protocols can be very efficient and still provide high security standards.

One question our solutions leave open is, if there are key exchange protocols with comparable security, but efficient simulators even for "passwords" of super-logarithmic length. It is tempting to believe that splitting the longer passwords into smaller chunks and then executing our protocol for passwords of logarithmic size would be a viable strategy. The main problem with this approach is that our simulator would need to know where the password chunks of the two parties coincide, and where they differ; in our proof the simulator acts differently in both cases. However, the simulator's password testing for the composed protocol through the ideal functionality only reveals if the long passwords match or do not, but does not give out the required information about the individual chunks. It seems as if using error correction or other types of tokens is not helpful in this regard. The other main question imposed by our paper is if previous results about token-based cryptography are affected by our stronger token model with unauthenticated transmissions.

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