J. Cryptology (2006) 19: 169–209
DOI: 10.1007/s00145-005-0307-3

Journal of
CRYPTOLOGY

Strengthening Zero-Knowledge Protocols Using Signatures*

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Communicated by Ronald Cramer

Received 27 February 2003 and revised 14 March 2005
Online publication 19 December 2005

Abstract. Recently there has been an interest in zero-knowledge protocols with stronger properties, such as concurrency, simulation soundness, non-malleability, and universal composability. In this paper we show a novel technique to convert a large class of existing honest-verifier zero-knowledge protocols into ones with these stronger properties in the common reference string model. More precisely, our technique utilizes a signature scheme existentially unforgeable against adaptive chosen-message attacks, and transforms any Σ-protocol (which is honest-verifier zero-knowledge) into a simulation sound concurrent zero-knowledge protocol. We also introduce Ω-*protocols*, a variant of Σ-protocols for which our technique further achieves the properties of non-malleability and/or universal composability.

In addition to its conceptual simplicity, a main advantage of this new technique over previous ones is that it avoids the Cook–Levin theorem, which tends to be rather

* A preliminary version of this paper appeared in Eurocrypt 2003. The work of Philip MacKenzie was primarily performed at Bell Labs. This research by Ke Yang was done at Bell Labs and Carnegie Mellon University, and was also partially sponsored by DIMACS, and by National Science Foundation (NSF) Grants CCR-0122581 and CCR-0085982.

inefficient. Indeed, our technique allows for very efficient instantiation based on the security of some efficient signature schemes and standard number-theoretic assumptions. For instance, one instantiation of our technique yields a universally composable zero-knowledge protocol under the Strong RSA assumption, incurring an overhead of a small constant number of exponentiations, plus the generation of two signatures.

Key words. Zero knowledge, Signatures, Simulation soundness, Non-malleability.

1. Introduction

The concept of a zero-knowledge (ZK) proof, as defined by Goldwasser et al. [34], has become a fundamental tool in cryptography. Informally, if a prover proves a statement to a verifier in ZK, then the verifier gains no information except for being convinced of the veracity of that statement. In particular, whatever the verifier could do after the ZK proof, it could have done before the ZK proof, in some sense because it can "simulate" the proof itself. In early work, Goldreich et al. [33] showed that any NP statement could be proven in (computational) ZK. In another early work, Goldreich et al. [32] showed the usefulness of ZK proofs in multiparty protocols, in particular, in having the parties prove the correctness of their computations. There has been a great deal of work since then on all properties of ZK proofs. Here we focus on a few such properties, namely, concurrency, non-malleability, simulation soundness, and universal composability, with our main goal being to construct *efficient* protocols that achieve these properties.

The problem of concurrency was first discussed by Dwork et al. [23]. (The more constrained problem of *parallelism* was discussed previously by Goldreich and Krawczyk [31].) Informally, the problem arises when many verifiers are interacting with a prover. An adversary controlling all the verifiers may coordinate the timing of their messages so that a simulator would not be able to simulate the execution of the prover in polynomial time. Canetti et al. [12] showed that without additional assumptions, such as timing constraints or a common reference string, logarithmic rounds are necessary to achieve concurrent (black-box) ZK. Kilian and Petrank [40] showed that polylogarithmic rounds suffice, and later Prabhakaran et al. [51] showed that logarithmic rounds suffice. On the other hand, Damgård [18] showed that concurrent, constant-round ZK protocols can be achieved in the common reference string model. Furthermore, Barak [1] showed that by using a non-black-box simulator, constant-round, concurrent protocols can be constructed in the plain model.[1]

The problem of malleability was first pointed out by Dolev et al. [22]. Roughly speaking, the problem is that an adversary may be able to play a "man-in-the-middle" attack on a ZK protocol, playing the role of the verifier in a first protocol, and that of the prover in a second protocol, and such that using information from the first protocol he is able to prove something in the second protocol that he could not prove without that information. A ZK protocol that does not suffer from this problem is said to achieve *one-time non-malleability* (since the adversary only interacts with one prover). Dolev et al. give a construction of a one-time non-malleable ZK protocol that uses a polylog-

[1] His construction, however, only admits bounded concurrency, meaning that the number of sessions that the protocol can execute concurrently and still retain its ZK property is at most a *fixed* polynomial in the security parameter.

arithmic number of communication rounds. Katz [39] describes efficient protocols for one-time non-malleable proofs of plaintext knowledge for several encryption schemes. His protocols work in the common reference string model, and consist of three rounds and a constant number of exponentiations. However, since the witness extractor uses "rewinding," the resulting protocols were only proven secure in a concurrent setting with the introduction of timing constraints. Barak [2] gives a construction of constant-round, one-time non-malleable ZK protocols in the plain model. His construction uses a non-black-box proof of security and is not very efficient. Sahai [54] provides a definition for one-time non-malleability in the case of non-interactive ZK (NIZK) proofs. De Santis et al. [20] generalize this to *unbounded non-malleability* of NIZK proofs, where even any polynomial number of simulator-constructed proofs does not help an adversary to construct any new proof. (As they do, for the remainder of this paper we simply refer to this property as *non-malleability*, leaving off the "unbounded" modifier.) Their definition is very strong in that (in some sense) it requires a witness to be extractable from the adversary.[2]

Further, they introduce the notion of a *robust* NIZK argument, which, in addition to being non-malleable, requires the so-called "simulator" of the ZK property to use a common reference string with the same distribution (uniform) as the one used by the real prover. (Following [20], we call this the *same-string ZK* property.) Finally, they give two constructions of non-malleable (and robust) ZK proofs for any NP language. In fact, these proofs are non-interactive, and thus achieve concurrent (constant-round) ZK.

The notion of simulation soundness for NIZK proofs was introduced by Sahai [54] in the context of chosen-ciphertext security of the Naor–Yung [46] encryption scheme. Informally, an NIZK proof is one-time simulation sound if even after seeing a "simulated proof" (which could be of a false statement) generated by the simulator, the adversary cannot generate a proof for a false statement. Sahai notes that the Naor–Yung encryption scheme would be adaptive chosen-ciphertext secure if it used a one-time simulation-sound NIZK proof. De Santis et al. [20] further generalized this notion to *unbounded simulation soundness*. An NIZK proof is unbounded simulation sound if even after seeing any polynomial number of simulated proofs, the adversary cannot generate a proof of a false statement. The non-malleable NIZK protocols given in [20] are also unbounded simulation sound. (For the remainder of this paper we simply refer to this property as *simulation soundness*, leaving off the "unbounded" modifier.)

The notions of simulation soundness, non-malleability, and robustness extend naturally to the case of interactive proof systems; we do this in Section 2. Informally, we say an interactive ZK protocol is simulation sound if the adversary cannot generate a proof of a false statement, even after interacting with any number of (simulated) provers. (See [44] for an application of simulation-sound ZK protocols in a threshold password-authenticated key exchange protocol.) We say a ZK protocol is non-malleable if there exists an efficient witness extractor that successfully extracts a witness from an adversary if the adversary would cause the verifier to accept, even when the adversary is also

[2] Note that this is a stricter definition than that of [22], and the definition we propose in this paper is even stricter. However, all known protocols achieving non-malleability according to the definition of [22] also achieve non-malleability according to our definition. Also, our definition is useful in proving a relationship between non-malleability and universal composability, as we discuss in our results, below.

allowed to interact with any number of (simulated) provers. We note that this definition of non-malleability implies that the ZK protocol is a proof of knowledge, and also that it satisfies the notion of "witness-extended emulation" from [3]. Naturally, a non-malleable ZK protocol is also simulation sound. Finally, we call a ZK protocol that is non-malleable and same-string, a *robust* ZK protocol.

Universal composability is a notion proposed by Canetti [10] to describe protocols that behave like ideal functionalities, and can be composed in arbitrary ways. Universal composability can be defined in either the adaptive model or the static model, denoting whether the adversary is allowed to corrupt parties adaptively, or must decide which parties to corrupt before the protocol starts, respectively. Universal composability is a very strong notion. For example, a universally composable ZK (UCZK) protocol is both non-malleable (at least in an intuitive sense) and concurrent.

Canetti [10] proved that UCZK protocols do not exist in the "plain" model, where there is no assumption about the system set-up. On the other hand, UCZK is possible in the common reference string model, which is the model we focus on in this paper. As pointed out by Canetti et al. [13], the non-malleable NIZK protocols of [20] are also UCZK protocols in the static corruption model. Since they use non-interactive proof techniques and general NP reductions, these protocols are not very efficient. Canetti and Fischlin [11] give a construction of a UCZK protocol for any NP language secure in the adaptive model. Basically, they use a standard three-round ZK protocol for Hamiltonian Cycle, except that they use universally composable commitments as a building block. Damgård and Nielsen [19] use the same general ZK protocol construction as Canetti and Fischlin, but with a more efficient UC commitment scheme.[3] Specifically, for a security parameter k, their UC commitment scheme allows commitment to k bits using a constant number of exponentiations and $O(k)$ bits of communication. Their most efficient UC commitment schemes are based on the p-subgroup assumption [47] or the decisional composite residuosity assumption (DCRA) [48]. Note that even with the more efficient UC commitment scheme, this approach to constructing UCZK protocols tends to be fairly inefficient, since a general NP reduction to Hamiltonian Cycle or SAT is used.

Our results. We show a new technique that allows us to convert certain types of honest-verifier ZK protocols into ZK protocols with the stronger properties described above, i.e., concurrency, simulation soundness, non-malleability, robustness, and/or universal composability, in the common reference string model. More precisely, we can

1. transform any Σ-protocol [16] (which are special three-round, honest-verifier protocols where the verifier only sends random bits) into a simulation-sound ZK protocol; and
2. transform any Ω-protocol (which we introduce in this paper as a variant of Σ-protocols) into a non-malleable ZK protocol, and further into a UCZK protocol.

The main transformations (sufficient to achieve all results except for UCZK protocols secure in the adaptive model) use a signature scheme that is existentially unforgeable against adaptive chosen-message attacks [34], which exists if one-way functions ex-

[3] In a later version of their paper, Damgård and Nielsen use SAT instead of Hamiltonian Cycle [19].

ist [53], as well as a Σ-protocol to prove knowledge of a signature. Note that one-way functions can be used to construct commitments, and thus if one-way functions exist, Σ-protocols exist for any NP statement (say, through a Cook–Levin reduction, and a standard Σ-protocol for Hamiltonian Cycle). Hence the requirement of our main transformations is the existence of one-way functions. On the other hand, certain signature schemes, such as the Cramer–Shoup [17] scheme and the DSA scheme [41], admit very efficient Σ-protocols. Using these schemes (and at the price of specific number-theoretic assumptions), we are able to construct strengthened ZK protocols that are more efficient than all previously known constructions, since we can completely avoid the Cook–Levin theorem [15], [42].[4] To achieve a UCZK protocol that is secure in the adaptive model, we also require a *simulation-sound trapdoor commitment* scheme, a new type of commitment scheme that we introduce and which may be of independent interest. This may be based on trapdoor permutations, but we show a more efficient version based on DSA.

We now sketch the intuition behind our technique. We first select two signature schemes, the second of which being a one-time signature scheme [25].[5] The common reference string will contain a randomly generated verification key vk for the first signature scheme, and hence neither the prover nor the verifier will know the corresponding signing key. We then take an honest-verifier ZK protocol Π for an NP statement φ, and we modify it to Π^*, which consists of (1) a witness indistinguishable (WI) proof for the statement

> "Either φ is true, or I know the signature for the message vk' w.r.t. verification key vk,"

where vk' is a freshly generated verification key for the one-time signature scheme that is also sent to the verifier, and (2) a signature on the transcript of the WI proof using the secret key corresponding to vk'.[6] Informally, Π^* is the "OR" of Π and a proof of knowledge of a signature on vk'. It turns out that if both Π and the proof of knowledge of the signature are so-called Σ-protocols [16] (see Section 2.2), then Π^* can be constructed from Π very efficiently [16]. Furthermore, if the signature scheme admits a very efficient proof, then the total overhead is very small. In particular, we show that if the Cramer–Shoup signature scheme [17] or the DSA signature scheme [41] is used, then the total overhead is only a constant number of exponentiations plus the generation of two signatures.

After the transformation, the completeness of protocol Π is obviously preserved. Protocol Π^* is also ZK, since a simulator generating the verification key in the common reference string can simultaneously generate the corresponding signing key, and thus has no problem simulating Π^*, by the witness indistinguishability of Π^*. Furthermore, we show that Π^* is simulation sound: If an adversary $\mathcal{A}$ is able to cause the verifier to accept a false statement after interacting with a polynomial number of (simulated) prover instances, then we show how to construct a machine M, which, having access to the signing oracle and interacting with $\mathcal{A}$, manages to forge a signature.

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[4] We note that previous constructions using the Cook–Levin theorem were only meant to show feasability, not efficiency.

[5] The second signature scheme may be the same as the first, although for greater efficiency, a signature scheme that is specifically designed for one-time use may be employed.

[6] The technique of proving "Either φ is true, or I know something" has been used previously (e.g., [38]).

In order to achieve non-malleability (and also robustness and universal composability) in this paper we introduce Ω-*protocols*, a variant of Σ-protocols that may be of independent interest. In a nutshell, an Ω-protocol is similar to a Σ-protocol but it assumes the existence of a common reference string and allows for the extraction of a witness from a single execution of the protocol without rewinding. As one example, we present an efficient Ω-protocol for the discrete logarithm relation based on the strong RSA assumption [4] and DCRA [48]. As another example, we present a "partial-extracting" Ω-protocol for proving knowledge of the plaintext of an ElGamal ciphertext [24] based on the Decision Diffie–Hellman assumption [6].

We show that if the original protocol Π is an Ω-protocol, then the transformed protocol Π^* is non-malleable, basically by noting that if one could not extract a witness for Π, then one could extract (and thus forge) a signature. Furthermore, the distribution of reference strings output by the simulator in our construction is identical to the distribution of reference strings in the real protocol. Therefore our construction is also robust ZK.

We then show that a non-malleable ZK protocol can be easily augmented to obtain a UCZK protocol in the static model. Invoking this result, we show as a corollary that (an "augmented" version of) Π^* is also a UCZK protocol in the static model. Finally, we show that we can further modify Π^* to be a UCZK protocol in the adaptive model (with erasures), while still maintaining efficiency. To achieve this we follow the approach of Damgård [18] and Jarecki and Lysyanskaya [37] of using a trapdoor commitment to commit to the first message of a Σ-protocol, which is then opened when sending the last message. However, it turns out that a "plain" trapdoor commitment scheme does not provide the properties we need to deal with adaptive corruptions. We thus introduce a stronger type of trapdoor commitment scheme, which we call a *simulation-sound trapdoor commitment* (SSTC) scheme. Furthermore, we demonstrate an efficient construction of an SSTC scheme under the DSA assumption.

We remark that if the properties are taken in the order (1) simulation soundness, (2) non-malleability, and (3) universal composability, each subsequent property is in some sense stronger, and requires more involved (and somewhat less efficient) protocols to achieve. Since all have useful applications, and one would generally like to use the simplest and most efficient protocol to solve a problem, it is therefore of interest to study all these properties. It is also of theoretical interest to understand the definitions of all these properties and how they relate to each other.

Organization of the paper.. In Section 2 we present formulations of the various notions of interactive ZK protocols in the common reference string setting, together with some of the building blocks that we will be using in our protocols. In Section 3 we present the construction of simulation-sound ZK protocols. In Section 4 we introduce Ω-protocols and present the construction of non-malleable (and robust) ZK protocols. In Section 5 we first show that non-malleable ZK implies UCZK assuming static corruptions, and then we demonstrate how to achieve UCZK in the adaptive model with erasures using an SSTC scheme. Finally, in Section 6 we present some efficient instantiations of the constructions above. They include using the Cramer–Shoup signature scheme and/or the DSA signature scheme to construct simulation-sound ZK protocols and non-malleable ZK protocols; an SSTC scheme based on DSA; an efficient Ω-protocol for the discrete logarithm relation (implying efficient non-malleable ZK and UCZK protocols for discrete

logarithm); and a *generalized* Ω-protocol for proving knowledge of the plaintext of an ElGamal ciphertext (implying an efficient non-malleable ZK protocol for ElGamal plaintext knowledge).

2. Preliminaries and Definitions

All our results will be in the *common reference string* (CRS) model, which assumes that there is a string generated from some distribution and is available to all parties at the start of a protocol. Note that this is more liberal than the *public random string* model, where a uniform distribution over fixed-length bit strings is assumed.

For a distribution Δ, we say $a \in \Delta$ to denote any element that has non-zero probability in Δ, i.e., any element in the support of Δ. We say $a \xleftarrow{R} \Delta$ to denote a is randomly chosen according to distribution Δ. For a set S, we say $a \xleftarrow{R} S$ to denote that a is uniformly drawn from S.

Throughout this paper, adversaries are modeled as non-uniform (interactive) Turing machines. On the other hand, simulators and extractors are uniform Turing machines, but they might become non-uniform by running a non-uniform adversary as a subroutine.

2.1. *ZK Proofs and Proofs of Knowledge*

In this section we provide definitions related to ZK proofs and proofs of knowledge. They are based on definitions of NIZK proofs from [20], but modified to allow interaction.

For a relation R, let $L_R = \{x\colon (x, w) \in R\}$ be the *language* defined by the relation. For any NP language L, note that there is a natural *witness relation* R containing pairs (x, w) where w is the witness for the membership of x in L, and that $L_R = L$. We use $k = |x|$ as the security parameter.

For two interactive machines A and B, we define $\langle A, B\rangle_{[\sigma]}(x)$ as the local output of B after an interactive execution with A using CRS σ, and common input x. The transcript of a machine is simply the common input x appended to the messages on its input and output communication tapes.[7] Two transcripts *match* if the ordered input messages of one are equivalent to the ordered output messages of the other, and vice versa. We use the notation $tr \bowtie tr'$ to indicate tr matches tr' and $tr \not\bowtie tr'$ to indicate that tr does not match tr'.

For some definitions below, we need to define security when an adversary is allowed to interact with more than one instance of a machine. Therefore it will be convenient to define a common *wrapper* machine that handles this "multi-session" type of interaction.[8] For an interactive machine A, we define $\boxed{A}$ to be a protocol wrapper for A, that takes two types of inputs on its communication tape:

- (START, π, x, w): For this message $\boxed{A}$ starts a new interactive machine A with label π, common input x, private input w, a freshly generated random input r, and using the same CRS as $\boxed{A}$ (i.e., all machines started by $\boxed{A}$ use the same CRS).

[7] Note that a transcript fixes the common input x. Therefore our definition of simulation-sound ZK (and non-malleable ZK) below implies that proofs are *uniquely applicable*, and one would not need to require this as a separate property (as was done in [54]).

[8] This is similar to the "multi-session extension" concept in [14].

– (MSG, π, m): For this message $\boxed{A}$ sends the message m to the interactive machine with label π (if it exists), and returns the output message of that machine.

We say $\boxed{A}_1$ is the wrapper of A that ignores all the subsequent START messages after seeing the first one. Effectively, $\boxed{A}_1$ is a "single-session" version of A. We define the output of $\boxed{A}_1$ to be a tuple (x, tr, v), where x is the common input (from the first START message), tr is the transcript of A (i.e., the input and output messages of the single machine A started with the first START message), and v is the output of A. In particular, if A is a verifier in a ZK protocol, this output will be 1 for accept, and 0 for reject. If A is an extractor (such as the machine $\mathcal{E}_2$ in the definition of a non-malleable ZK proof of knowledge below), then this output will be a pair (b, w) where b is a bit designating accept/reject (as above), and w is the "extracted" witness if $b = 1$. We define the output of $\boxed{A}$ to be a tuple $(\vec{x}, \vec{tr}, \vec{v})$, where the machine started with the ith START message would produce the tuple $(x[i], tr[i], v[i])$, as described for the single machine started in $\boxed{A}_1$.

We say two interactive machines B and C are *coordinated* if they have a single control, but two distinct sets of input/output communication tapes. (Note that B and C may run concurrently.) For four interactive machines A, B, C, and D we define $(\langle A, B\rangle, \langle C, D\rangle)_{[\sigma]}$ as the local output of D after an interactive execution with C and after an interactive execution of A and B, all using CRS σ. Note that we will only be concerned with this if B and C are coordinated.

We note that all our ZK definitions use black-box, non-rewinding simulators, and our proofs of knowledge use non-rewinding extractors.

Definition 2.1 (Unbounded ZK Proof). $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2))$ is an *unbounded ZK proof* (resp., *argument*) *system* for an NP language L with witness relation R if $\mathcal{D}$ is an ensemble of polynomial-time samplable distributions, $\mathcal{P}$, $\mathcal{V}$, and $\mathcal{S}_2$ are probabilistic polynomial-time interactive machines, and $\mathcal{S}_1$ is a probabilistic polynomial-time machine, such that

Completeness. There exists a negligible function $\alpha(k)$ such that for all $k > 0$, for all $x \in L$ of length k, all w such that $R(x, w) = 1$, and all $\sigma \in \mathcal{D}_k$, the probability that $\langle \mathcal{P}(w), \mathcal{V}\rangle_{[\sigma]}(x) = 0$ is less than $\alpha(k)$.

Soundness. For all unbounded (resp., non-uniform probabilistic polynomial-time) adversaries $\mathcal{A}$, there exists a negligible function $\alpha(k)$ such that for all $k > 0$, for all $x \notin L$ of length k, the probability that $\langle \mathcal{A}, \mathcal{V}\rangle_{[\sigma]}(x) = 1$ where $\sigma \stackrel{R}{\leftarrow} \mathcal{D}_k$ is less than $\alpha(k)$.

Unbounded ZK. For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A}$, $|\Pr[\mathsf{Expt}_{\mathcal{A}}(k) = 1] - \Pr[\mathsf{Expt}^{\mathcal{S}}_{\mathcal{A}}(k) = 1]|$ is negligible, where the experiments $\mathsf{Expt}_{\mathcal{A}}(k)$ and $\mathsf{Expt}^{\mathcal{S}}_{\mathcal{A}}(k)$ are defined as follows:

$\mathsf{Expt}_{\mathcal{A}}(\kappa)$:	$\mathsf{Expt}^{\mathcal{S}}_{\mathcal{A}}(\kappa)$:
$\sigma \stackrel{R}{\leftarrow} \mathcal{D}_k$	$(\sigma, \tau) \leftarrow \mathcal{S}_1(1^k)$
Return $\langle \boxed{\mathcal{P}}, \mathcal{A}\rangle_{[\sigma]}$	Return $\langle \boxed{\mathcal{S}'(\tau)}, \mathcal{A}\rangle_{[\sigma]}$

where $\mathcal{S}'(\tau)$ runs as follows on common reference string σ, common input x, and private input w: if $R(x, w) = 1$, $\mathcal{S}'(\tau)$ runs $\mathcal{S}_2(\tau)$ on common reference string

σ and common input x; otherwise $\mathcal{S}'(\tau)$ runs $\mathcal{S}_{\text{null}}$, where $\mathcal{S}_{\text{null}}$ is an interactive machine that simply aborts.[9]

We point out that this definition only requires the simulator to simulate a valid proof, which is implemented by having $\mathcal{S}'$ have access to the witness w and only invoking $\mathcal{S}_2$ when w is valid.[10] However, $\mathcal{S}_2$ does not access the witness and will simulate a proof from the input x only.

We further note that by the definition of the wrapper machines, it is the adversary $\mathcal{A}$ that chooses the common input x in the definition of unbounded ZK.

Definition 2.2 (Same-String Unbounded ZK). $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2))$ is a *same-string unbounded ZK argument* system for an NP language L with witness relation R if Π is an unbounded ZK argument system for L with the additional property that the distribution of the reference string output by $\mathcal{S}_1(1^k)$ is exactly $\mathcal{D}_k$.

We only define same-string unbounded ZK arguments since, as shown in [20], any protocol that is same-string unbounded ZK must be an argument, and not a proof.

The following defines (unbounded) SSZK (SSZK). This has been useful in applications. In particular, as shown in [54], the one-time version of SSZK suffices for the security of a (non-interactive) ZK protocol in the construction of adaptive chosen-ciphertext secure cryptosystems using the Naor–Yung [46] paradigm. We directly define the unbounded version, needed in other applications such as threshold password-authenticated key exchange [44].

Definition 2.3 (Simulation-Sound ZK). $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2))$ is a *simulation-sound ZK proof* (resp., *argument*) *system* for an NP language L if Π is an unbounded ZK proof (resp., argument) system for L and, furthermore,

Simulation Soundness. For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, where $\mathcal{A}_1$ and $\mathcal{A}_2$ are coordinated, $\Pr[\mathsf{Expt}_{\mathcal{A}}(k) = 1]$ is negligible, where $\mathsf{Expt}_{\mathcal{A}}(k)$ is defined as follows:

$\mathsf{Expt}_{\mathcal{A}}(k)$:
$(\sigma, \tau) \leftarrow \mathcal{S}_1(1^k)$
$(x, tr, b) \leftarrow \left(\langle \boxed{\mathcal{S}''(\tau)}, \mathcal{A}_1 \rangle, \langle \mathcal{A}_2, \boxed{\mathcal{V}}_1 \rangle\right)_{[\sigma]}$
Let Q be the set of transcripts of machines in $\boxed{\mathcal{S}''(\tau)}$
Return 1 iff $b = 1$, $x \notin L$, and for all $tr' \in Q$, $tr \not\rhd\lhd tr'$

where $\mathcal{S}''(\tau)$ runs as follows on CRS σ, common input x and private input w: $\mathcal{S}''(\tau)$ runs $\mathcal{S}_2(\tau)$ on CRS σ and common input x.

[9] Without loss of generality, we assume that if the common input is not of size k (we implicitly assume that both $\mathcal{A}$ and $\mathcal{P}$ can determine k from σ) or the input to $\mathcal{P}$ is not a witness for the common input, $\mathcal{P}$ simply aborts.

[10] $\mathcal{A}$ must supply a witness, since $\mathcal{P}$ is restricted to polynomial time, and thus may not be able to generate a witness itself. This may seem odd compared with definitions of standard ZK that assume an unbounded prover, but it does seem to capture the correct notion of unbounded ZK, and in particular does not allow $\mathcal{A}$ to test membership in L. See [54] for more discussion.

In the above definition, we emphasize that $\mathcal{S}_2$ may be asked to simulate *false* proofs for $x \notin L_R$, since $\mathcal{S}''$ does not check whether $(x, w) \in R$. The idea is that even if the adversary is able to obtain acceptable proofs on false statements, it will not be able to produce any new acceptable proof on a false statement.

The following defines non-malleable zero-knowledge (NMZK) proofs (resp., arguments) of knowledge. If a protocol is NMZK according to our definition, then this implies the protocol is also NMZK in the explicit witness sense (as defined in [20]).[11] Moreover, we show that the protocol is also UCZK in the model of static corruptions. Also note that simulation soundness is implied by this definition.

Definition 2.4 (Non-Malleable ZK Proof/Argument of Knowledge). $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2), \mathcal{E} = (\mathcal{E}_1, \mathcal{E}_2))$ is a *non-malleable ZK proof* (resp., *argument*) *of knowledge system* for an NP language L with witness relation R if Π is an unbounded ZK proof (resp., argument) system for L and furthermore, $\mathcal{E}_1$ and $\mathcal{E}_2$ are probabilistic polynomial-time machines such that

- **Reference String Indistinguishability.** The distribution of the first output of $\mathcal{S}_1(1^k)$ is identical to the distribution of the first output of $\mathcal{E}_1(1^k)$.
- **Extractor Indistinguishability.** For any $\tau \in \{0, 1\}^*$, the distribution of the output of $\boxed{\mathcal{V}}_1$ is identical to the distribution of the restricted output of $\boxed{\mathcal{E}_2(\tau)}_1$, where the restricted output of $\boxed{\mathcal{E}_2(\tau)}_1$ does not include the extracted value.
- **Extraction.** For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, where $\mathcal{A}_1$ and $\mathcal{A}_2$ are coordinated machines, $|\Pr[\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k) = 1] - \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = 1]|$ is negligible, where the experiments $\mathsf{Expt}_{\mathcal{A}}(k)$ and $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$ are defined as follows:

$\mathsf{Expt}_{\mathcal{A}}(k)$:	$\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$:
$(\sigma, \tau) \leftarrow \mathcal{S}_1(1^k)$	$(\sigma, \tau_1, \tau_2) \leftarrow \mathcal{E}_1(1^k)$
$(x, tr, b) \leftarrow \left(\left\langle \boxed{\mathcal{S}''(\tau)}, \mathcal{A}_1 \right\rangle, \left\langle \mathcal{A}_2, \boxed{\mathcal{V}}_1 \right\rangle\right)_{[\sigma]}$	$(x, tr, (b, w)) \leftarrow \left(\left\langle \boxed{\mathcal{S}''(\tau_1)}, \mathcal{A}_1 \right\rangle, \left\langle \mathcal{A}_2, \boxed{\mathcal{E}_2(\tau_2)}_1 \right\rangle\right)_{[\sigma]}$
Let Q be the set of transcripts of machines in $\boxed{\mathcal{S}''(\tau)}$.	Let Q be the set of transcripts of machines in $\boxed{\mathcal{S}''(\tau_1)}$.
Return 1 iff $b = 1$ and for all $tr' \in Q$, $tr \not\bowtie tr'$	Return 1 iff $b = 1$, $(x, w) \in R$, and for all $tr' \in Q$, $tr \not\bowtie tr'$

  where $\mathcal{S}''(\tau)$ runs as follows on CRS σ, common input x, and private input w: $\mathcal{S}''(\tau)$ runs $\mathcal{S}_2(\tau)$ on CRS σ and common input x.

In the above definition, as in the definition of SSZK protocols, we emphasize that $\mathcal{S}_2$ may be asked to simulate *false* proofs for $x \notin L_R$, since $\mathcal{S}''$ does not check whether $(x, w) \in R$. The idea is that even if the adversary is able to obtain acceptable proofs

[11] Note that our notion of NMZK (along with that of [20]) implies a proof of knowledge, whereas the original notion of non-malleable zero-knowledge from [22] does not. However, to our knowledge there is no NMZK protocol that is not also a proof of knowledge. Since we achieve extraction in our protocols also, we feel that our definition is useful. Also, for brevity, we generally refer to this notion as an NMZK proof, rather than an NMZK proof of knowledge.

on false statements, it will not be able to produce any new acceptable proof for which a witness cannot be extracted.

To conclude with the ZK definitions, we generalize the notion of robust NIZK in [20] to the interactive setting.

Definition 2.5 (Robust ZK). Π is a *robust ZK argument of knowledge system* for an NP language L with witness relation R if Π is a non-malleable *and* same-string unbounded ZK argument of knowledge system for L.

2.2. Σ-*Protocols*

Here we provide an overview of the basic definitions and properties of Σ-protocols [16]. Assume $R = \{(x, w)\}$ is a relation such that for some given polynomial $p(\cdot)$ it holds that $|w| \leq p(|x|)$ for all $(x, w) \in R$. Furthermore, let R be testable in polynomial time. Recall that $L_R = \{x\colon (x, w) \in R\}$ is the *language* defined by the relation. For all $x \in L_R$, let $W_R(x) = \{w\colon (x, w) \in R\}$ be the *witness set* for x.

A Σ-protocol (A, B) is defined to be a three-move interactive protocol between a probabilistic polynomial-time prover A and a probabilistic polynomial-time verifier B, where the prover acts first. The verifier is only required to send random bits as a challenge to the prover. For some $(x, w) \in R$, the common input to both players is x while w is a private input to the prover. For such given x, let (a, c, z) denote the conversation between the prover and the verifier. To compute the first and final messages, the prover invokes efficient algorithms $a(\cdot)$ and $z(\cdot)$, respectively, using (x, w) and common random bits as input. Using an efficient predicate $\varphi(\cdot)$, the verifier decides whether the conversation is accepting with respect to x. The relation R and the algorithms $a(\cdot)$, $z(\cdot)$, and $\varphi(\cdot)$ are public. The length of the challenges is denoted t_B, and we assume that t_B only depends on the length of the common string x.

We need to broaden this definition slightly, to deal with cheating provers. We define $\hat{L}_R$ to be the input language, with the property that $L_R \subseteq \hat{L}_R$, and membership in $\hat{L}_R$ may be tested in polynomial time. We implicitly assume B only executes the protocol if the common input $x \in \hat{L}_R$.

All Σ-protocols presented here will satisfy the following security properties:

- *Weak special soundness*: Let (a, c, z) and (a, c', z') be two conversations that are accepting for some given $x \in \hat{L}_R$. If $c \neq c'$, then $x \in L_R$. The pair of accepting conversations (a, c, z) and (a, c', z') with $c \neq c'$ is called a *collision*.
- *Special honest-verifier zero knowledge* (SHVZK): A Σ-protocol is honest-verifier zero knowledge (HVZK) if there is a (probabilistic polynomial-time) simulator M that on input $x \in L_R$ generates accepting conversations with a distribution that is indistinguishable[12] from when A and B execute the protocol on common input x (and A is given a witness w for x), and B indeed honestly chooses its challenges uniformly at random. Furthermore, the Σ-protocol is SHVZK if this simulator can additionally take a string c as input, and output an accepting conversation for x where c is the challenge, where the distribution of accepting conversations is

[12] Often this is required to be perfectly indistinguishable, but we generalize the definition slightly to require only computational indistinguishability.

indistinguishable from when A and B execute the protocol on common input x, conditioned on c being the challenge. (Note in particular that this indistinguishability requirement must hold for *every* challenge c.) In fact, we assume the simulator is able to output an accepting conversation for every challenge c for not only $x \in L_R$, but also any $x \in \hat{L}_R$. Naturally, there is no requirement for indistinguishability in this case, since A would simply not execute the protocol.

Formally, we can specify the indistinguishability requirement as follows. For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$ there is a negligible function $\alpha(k)$ such that for all $k > 0$ and all $c \in \{0,1\}^k$, $|\Pr[\mathsf{Expt}_{\mathcal{A}}(k,c) = 1] - \Pr[\mathsf{Expt}^M_{\mathcal{A}}(k,c) = 1]| \leq \alpha(k)$, where the experiments $\mathsf{Expt}_{\mathcal{A}}(k,c)$ and $\mathsf{Expt}^M_{\mathcal{A}}(k,c)$ are defined as follows:

$\mathsf{Expt}_{\mathcal{A}}(k,c)$:	$\mathsf{Expt}^M_{\mathcal{A}}(k,c)$:
$(x,w,s) \leftarrow \mathcal{A}_1(c)$	$(x,w,s) \leftarrow \mathcal{A}_1(c)$
If $(x,w) \notin R$ return 0	If $(x,w) \notin R$ return 0
$r \stackrel{R}{\leftarrow} \{0,1\}^*$	Return $\mathcal{A}_2(s, M(x,c))$
$a \leftarrow a(x,w,r)$	
Return $\mathcal{A}_2(s,(a,c,z(x,w,r,c)))$	

Some of the Σ-protocols also satisfy the following property:

- *Special soundness*: Let (a,c,z) and (a,c',z') be two conversations that are accepting for some given x, with $c \neq c'$. Then given x and those two conversations, a witness w such that $(x,w) \in R$ can be computed efficiently.

In our results to follow, we need a particular simple instance of the main theorem from [16]. Specifically, we use a slight generalization of a corollary in [16] which enables a prover, given two relations (R_1, R_2), values $(x_1, x_2) \in \hat{L}_{R_1} \times \hat{L}_{R_2}$, and corresponding three-move Σ-protocols $((A_1, B_1), (A_2, B_2))$, to present a three-move Σ-protocol $(A_{\text{or}}, B_{\text{or}})$ for proving the existence of a w such that either $(x_1, w) \in R_1$ or $(x_2, w) \in R_2$. We call this the "OR" protocol for $((A_1, B_1), (A_2, B_2))$. Technically, this "OR" protocol is for the relation $R = \{((x_1,x_2), w)\colon ((x_1,w) \in R_1 \wedge x_2 \in \hat{L}_{R_2}) \vee ((x_2,w) \in R_2 \wedge x_1 \in \hat{L}_{R_1})\}$ and input language $\hat{L}_R = \{(x_1,x_2)\colon x_1 \in \hat{L}_{R_1} \wedge x_2 \in \hat{L}_{R_2}\}$.

We describe the protocol assuming the challenges from (A_1, B_1) and (A_2, B_2) are of the same length. This can easily be generalized, as long as the challenge length in the combined protocol is at least as long as the challenges from either protocol. The protocol consists of (A_1, B_1) and (A_2, B_2) running in parallel, but with the verifier's challenge c split into $c = c_1 \oplus c_2$, with c_1 as the challenge for (A_1, B_1), and c_2 as the challenge for (A_2, B_2).

The protocol for A_{or} is as follows: Without loss of generality, say A_{or} knows w such that $(x_1, w) \in R_1$. Let M_2 be the simulator for S_2. Then A_{or} runs $M_2(x_2)$ to generate (m, e, z). It sends the first message of (A_1, B_1), along with m as the first message of (A_2, B_2). On challenge c, it chooses $c_2 = e$, and $c_1 = c \oplus c_2$. It is able to provide the final response in (A_1, B_1) because it knows w, and the final response in (A_2, B_2) is simply z. The final message of A_{or} includes c_1 along with the final responses for (A_1, B_1) and (A_2, B_2).

We note that if (A_2, B_2) satisfies special soundness, then $(A_{\text{or}}, B_{\text{or}})$ satisfies the following property:

- *Half-weak special soundness*: Let (a, c, z) and (a, c', z') be two conversations that are accepting for some given (x_1, x_2), with $c \neq c'$. Then either there exists a w_1 such that $(x_1, w_1) \in R_1$ or given x and those two conversations, a witness w_2 such that $(x_2, w_2) \in R_2$ can be computed efficiently.

For two Σ-protocols, (A_1, B_1) and (A_2, B_2), let $(A_1, B_1) \vee (A_2, B_2)$ denote the "OR" protocol for $((A_1, B_1), (A_2, B_2))$.

A simple but important result from [16] states that if a Σ-protocol is HVZK (with perfectly indistinguishable simulations), the protocol is *witness indistinguishable* (WI) [26]. Although the corresponding result with computationally indistinguishable simulations does not apply, one can show a specific result for an "OR" protocol:

Lemma 2.6. *Say Σ-protocol $(A, B) = (A_1, B_1) \vee (A_2, B_2)$, where (A_1, B_1) and (A_2, B_2) are both HVZK (with computationally indistinguishable simulations). Let R, R_1, and R_2 be the associated relations as discussed above. Then (A, B) is WI over R', where $R' = \{((x_1, x_2), w)\colon ((x_1, x_2), w) \in R \wedge x_1 \in L_{R_1} \wedge x_2 \in L_{R_2}\}$.*

Proof. The intuitive reason is that the challenge is split randomly between the two Σ-protocols (A_1, B_1) and (A_2, B_2), and thus in some sense one can reduce the security to the honest verifier case. Formally, we proceed as follows. Say there is a string $(x_1, x_2) \in L_{R'}$ with two witnesses w_1 and w_2, and there exists a verifier B' and string y such that the output of $B'((x_1, x_2), y)$ interacting with $A((x_1, x_2), w_1)$ is distinguishable from the output of $B'((x_1, x_2), y)$ interacting with $A((x_1, x_2), w_2)$. Call these experiments E_1 and E_2, respectively. Without loss of generality, we may assume $(x_1, w_1) \in R_1$ and $(x_2, w_2) \in R_2$.[13] Now define $R^* = \{((x_1, x_2), (w_1, w_2)) : (x_1, w_1) \in R_1 \wedge (x_2, w_2) \in R_2\}$, and consider the Σ-protocol (A^*, B^*) over R^* that simply runs (A_1, B_1) and (A_2, B_2), but where A^* chooses a random challenge c_1 and sets $c_2 \leftarrow c \oplus c_1$, where c is the challenge of B^*. Consider $B'((x_1, x_2), y)$ interacting with $A^*((x_1, x_2), (w_1, w_2))$ and call this experiment E^*. Obviously the output of B' in experiment E^* must be distinguishable from the output of B' in either experiment E_1 or E_2. Say it is E_2. (The other case is similar.) Then we will show that (A_1, B_1) is not HVZK. (Formally, the definition of HVZK is like the definition of SHVZK, except that c is chosen randomly after $\mathcal{A}_1$ is executed, instead of being an input to the two experiments. Also, the input to $\mathcal{A}_1$ is simply 1^k. It is easy to see that the SHVZK property implies the HVZK property.)

Construct an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$ as follows. Let $((x_1, x_2), (w_1, w_2))$ be such that for some y the output of $B'((x_1, x_2), y)$ in experiment E^* is distinguishable from the output of $B'((x_1, x_2), y)$ in experiment E_2. Let $\mathcal{A}_1(1^k)$ output $(x_1, w_1, (x_1, x_2, w_1, w_2, y))$. Let $\mathcal{A}_2((x_1, x_2, w_1, w_2, y), (a_1, c_1, z_1))$ generate $r_2 \stackrel{R}{\leftarrow} \{0, 1\}^k$ and $a_2 \leftarrow a_2(x_2, w_2, r_2)$, and then invoke $B'((x_1, x_2), y)$ with first message (a_1, a_2). When B' returns challenge

[13] For example, if $(x_1, w_1), (x_1, w_2) \in R_1$, then there exists a w_3 such that $(x_2, w_3) \in R_2$ (by the definition of R') and the output of $B'((x_1, x_2), y)$ interacting with $A((x_1, x_2), w_3)$ would necessarily be distinguishable from the output of the interaction with either $A((x_1, x_2), w_1)$ or $A((x_1, x_2), w_2)$. Then one could use either w_1 and w_3, or w_2 and w_3, as the two witnesses.

c, compute $c_2 \leftarrow c \oplus c_1$ and $z_2 \leftarrow z_2(x_2, w_2, r_2, c_2)$, and send back $(z_1, z_2), c_1$. Then output whatever B' outputs. The distribution of $\mathsf{Expt}_{\mathcal{A}}(k)$ is exactly the distribution of the output of E^*, and the distribution of $\mathsf{Expt}^{M}_{\mathcal{A}}(k)$ is exactly the distribution of the output of E_2, and thus they are distinguishable. □

2.3. *Signature Schemes*

A *signature scheme* SIG is a triple (sig_gen, sig_sign, sig_verify) of algorithms, the first two being probabilistic, and all running in polynomial time (with a negligible probability of failing). sig_gen takes as input 1^k and outputs a public key pair (sk, vk), i.e., $(sk, vk) \leftarrow \mathsf{sig\_gen}(1^k)$. sig_sign takes a message m and a secret key sk as input and outputs a signature σ for m, i.e., $\sigma \leftarrow \mathsf{sig\_sign}(sk, m)$. sig_verify takes a message m, a public key vk, and a candidate signature σ' for m as input and returns the bit $b = 1$ if σ' is a valid signature for m for the corresponding private key, and otherwise returns the bit $b = 0$. That is, $b \leftarrow \mathsf{sig\_verify}(vk, m, \sigma')$. Naturally, if $\sigma \leftarrow \mathsf{sig\_sign}(sk, m)$, then $\mathsf{sig\_verify}(vk, m, \sigma) = 1$.

Security for signature schemes. We specify existential unforgeability against adaptive chosen-message attacks [35] for a signature scheme $\mathsf{SIG} = (\mathsf{sig\_gen}, \mathsf{sig\_sign}, \mathsf{sig\_verify})$. A forger is given vk, where $(sk, vk) \leftarrow \mathsf{sig\_gen}(1^k)$, and tries to forge signatures with respect to vk. It is allowed to query a signature oracle OSign_{vk} (that produces signatures that can be verified with vk) on messages of its choice. It succeeds if after this it outputs a message/signature pair (m, σ) that is valid (i.e., where $\mathsf{sig\_verify}(vk, m, \sigma) = 1$), but m was not one of the messages signed by the signature oracle. We say a forger (t, q, ε)-breaks a scheme if the forger runs in time $t(k)$ makes $q(k)$ queries to the signature oracle, and succeeds with probability at least $\varepsilon(k)$. A signature scheme SIG is existentially unforgeable against adaptive chosen-message attacks if for all t and q polynomial in k, if a forger (t, q, ε)-breaks SIG, then ε is negligible in k.

In a *strong one-time* signature scheme [54], security is formulated as above except that (1) the forger may only query the signature oracle once, and (2) the forger succeeds if it outputs a valid message/signature pair (m, σ) such that either m was not signed by the signature oracle, or σ was not output by the signature oracle when queried with m. We call the resulting security property strong existential unforgeability against a one-message attack. We note that strong one-time signature schemes can be made very efficient since they do not need public-key cryptographic operations [25].

3. SSZK

We are now ready to present the first result achieved with our technique: An SSZK protocol for a relation $R = \{(x, w)\}$. We assume that we have the following building blocks:

1. Σ^R: a Σ-protocol for the binary relation R.
2. $\mathsf{SIG}_{\mathsf{adap}} = (\mathsf{sig\_gen}_{\mathsf{adap}}, \mathsf{sig\_sign}_{\mathsf{adap}}, \mathsf{sig\_verify}_{\mathsf{adap}})$: a signature scheme (existentially unforgeable against chosen-message attacks).

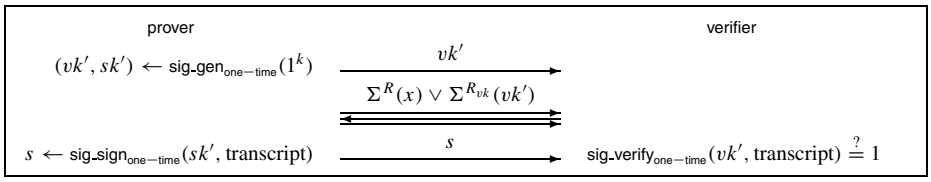

Fig. 1. $\mathsf{SS}^R_{[vk]}(x)$: An SSZK protocol for relationship R with CRS vk (drawn from the distribution $\mathsf{sig\_gen}_{\mathsf{adap}}(1^k)$), and common input x. The prover also knows the witness w such that $R(x, w) = 1$.

3. $R_{vk} = \{(m, s) \mid \mathsf{sig\_verify}_{\mathsf{adap}}(vk, m, s) = 1\}$: a binary relation of message–signature pairs.
4. $\Sigma^{R_{vk}}$: a Σ-protocol with the special soundness property for the binary relation R_{vk}.
5. $\mathsf{SIG}_{\mathsf{one-time}} = (\mathsf{sig\_gen}_{\mathsf{one-time}}, \mathsf{sig\_sign}_{\mathsf{one-time}}, \mathsf{sig\_verify}_{\mathsf{one-time}})$: a strong one-time signature scheme.

The protocol $\mathsf{SS}^R_{[vk]}(x)$ is shown in Fig. 1. It assumes the prover and verifier share a common input x to a Σ-protocol Σ^R, and the prover knows w such that $(x, w) \in R$. The CRS σ is the verification key vk of a signature scheme that is existentially unforgeable against adaptive chosen-message attacks. The prover generates a pair (vk', sk') for a strong one-time signature scheme, and sends vk' to the verifier. After this, vk' is the common input to a Σ-protocol $\Sigma^{R_{vk}}$ satisfying special soundness. Then the prover uses the OR construction for Σ-protocols to prove that either $x \in L_R$ or it knows a signature for vk' under verification key vk. (Note that since $\Sigma^{R_{vk}}$ satisfies special soundness, intuitively it is a proof of knowledge.) Finally, the prover signs the transcript with sk', and sends the resulting signature to the verifier.

Now we must describe $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2)$ for $\mathsf{SS}^R_{[vk]}(x)$. $\mathcal{S}_1(1^k)$ first generates signature keys $(vk, sk) \leftarrow \mathsf{sig\_gen}_{\mathsf{adap}}(1^k)$ and outputs $(\sigma, \tau) = (vk, sk)$. $\mathcal{S}_2(sk)$ first checks that common input $x \in \hat{L}_R$. If not, it aborts. Otherwise it runs the protocol as normal, except generating $s' \leftarrow \mathsf{sig\_sign}_{\mathsf{adap}}(sk, vk')$, and using knowledge of s' to complete the Σ-protocol $\Sigma^R(x) \vee \Sigma^{R_{vk}}(vk')$.

Theorem 3.1. *The protocol* $\mathsf{SS}^R_{[vk]}(x)$ *is an SSZK argument.*

Proof. *Completeness*: Straightforward.

Unbounded ZK: By inspection, $\mathcal{S}_1(1^k)$ produces exactly the same distribution as the real protocol. Then by the fact that $\mathcal{S}'(\tau)$ runs $\mathcal{S}_2(\tau)$ only when $(x, w) \in R$, and by the fact that $\Sigma^R(x) \vee \Sigma^{R_{vk}}(vk')$ is an "OR" type of Σ-protocol, and thus witness indistinguishable,[14] unbounded ZK follows by a straightforward hybrid argument.

Simulation soundness: For an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, recall the experiment $\mathsf{Expt}_{\mathcal{A}}(k)$ in the definition of SSZK. Let $p = \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = 1]$. Our goal is to show that p is negligible.

[14] This implication (from Lemma 2.6) requires $x \in L_R$ and $vk' \in L_{R_{vk}}$, the former always being true when proving unbounded ZK, and the latter always being true by the definition of signatures.

Say a *forgery* occurs if $\mathcal{V}$ accepts, and the one-time verification key vk' in that session was used by $\mathcal{S}_2(\tau)$, but on a different transcript, or resulting in a different signature. Let $\mathsf{Expt}'_{\mathcal{A}}(k)$ be $\mathsf{Expt}_{\mathcal{A}}(k)$ except that if a *forgery* occurs, the experiment halts and fails. Let $p' = \Pr[\mathsf{Expt}'_{\mathcal{A}}(k) = 1]$.

First, by the strong existential unforgeability property of $\mathsf{SIG}_{\mathsf{one-time}}$, we show that the difference between p and p' is negligible. We do this by constructing a non-uniform probabilistic polynomial-time attacker $\mathcal{B}_1$ that can break $\mathsf{SIG}_{\mathsf{one-time}}$ with probability $\varepsilon_1 = (1/c)(p - p')$, where c is the number of sessions $\mathcal{A}_2$ starts with the simulator in $\mathsf{Expt}_{\mathcal{A}}(k)$. The input to $\mathcal{B}_1$ is a verification key vk' and a one-time signature oracle $\mathsf{OSign}_{vk'}$. $\mathcal{B}_1$ chooses $d \stackrel{R}{\leftarrow} \{1, \ldots, c\}$, and then runs the experiment $\mathsf{Expt}_{\mathcal{A}}(k)$, running the simulator and verifier as normal, except for inserting vk' into the dth instance of $\mathcal{S}_2(\tau)$ and using $\mathsf{OSign}_{vk'}$ to perform the signature operation for vk' in that instance. If a forgery occurs with verification key vk', $\mathcal{B}_1$ halts and outputs the forgery, i.e., the transcript and signature provided by $\mathcal{A}_2$ for its session with $\mathcal{V}$. The view of $\mathcal{A}$ in this slightly modified experiment is the same as the view of $\mathcal{A}$ in $\mathsf{Expt}_{\mathcal{A}}(k)$ until a forgery occurs. Thus, since a forgery occurs with probability $p - p'$, and since if a forgery occurs, $\mathcal{B}_1$ will break the $\mathsf{SIG}_{\mathsf{one-time}}$ on vk' with probability $(1/c)$, $\mathcal{B}_1$ breaks $\mathsf{SIG}_{\mathsf{one-time}}$ with probability $\varepsilon_1 = (1/c)(p - p')$.

Now by the existential unforgeability property of $\mathsf{SIG}_{\mathsf{adap}}$, we show that p' is negligible. We do this by constructing a non-uniform probabilistic polynomial-time attacker $\mathcal{B}_0$ that can break $\mathsf{SIG}_{\mathsf{adap}}$ with at most $2c$ signature oracle queries (again, where c is the number of sessions $\mathcal{A}_2$ starts with the simulator in $\mathsf{Expt}_{\mathcal{A}}(k)$), and with probability at least $\varepsilon_0 = (p')^2 - 2^{-k}$.[15] The input to $\mathcal{B}_0$ is a verification key vk and a signature oracle OSign_{vk}. $\mathcal{B}_0$ runs experiment $\mathsf{Expt}'_{\mathcal{A}}(k)$, running the simulator and verifier as normal, except for inserting vk into the CRS and using OSign_{vk} to perform all signature operations with respect to vk. Also, before $\mathcal{V}$ sends a challenge to $\mathcal{A}_2$, $\mathcal{B}_0$ forks the experiment and continues independently in each sub-experiment (thus giving independent random challenges to $\mathcal{A}_2$). $\mathcal{B}_0$ then examines the output (x, tr_1, b_1) and (x, tr_2, b_2) in each sub-experiment. Now consider the case when $b_1 = b_2 = 1$ and $x \notin L_R$ (call this a *successful* sub-experiment), and also the challenges in each sub-experiment are distinct. Then since $\Sigma^R(x) \vee \Sigma^{R_{vk}}(vk')$ satisfies half-weak special soundness, $\mathcal{B}_0$ can generate a signature s on vk' with respect to key vk using the two transcripts tr_1 and tr_2. (Here vk' is the one-time verification key sent in the first message of both tr_1 and tr_2. By the definition of $\mathsf{Expt}'_{\mathcal{A}}(k)$, vk' could not have been used in any instance of $\mathcal{S}_2$ in either sub-experiment.) Thus $\mathcal{B}_0$ generates a signature (on a new message vk') with respect to vk, and breaks $\mathsf{SIG}_{\mathsf{adap}}$. By inspection, $\mathcal{B}_0$ makes at most $2c$ calls to the signature oracle.

Now we determine the success probability of $\mathcal{B}_0$. First note that for each sub-experiment, the view of $\mathcal{A}$ is perfectly indistinguishable from the view of $\mathcal{A}$ in $\mathsf{Expt}'_{\mathcal{A}}(k)$, and thus the probability of success in each sub-experiment is p'. Second, note that the probability of a random collision on k-bit challenges is 2^{-k}. Then we can determine the

[15] The following argument is a simple version of the Forking Lemma [50], although it does not follow directly, since we are using a signature oracle, and the adversary's output is not actually a signature from that scheme, but a Σ-protocol of knowledge of the signature. Consequently, rather than trying to force our results into the notation of [50] and prove why the Forking Lemma should hold in our situation, we simply prove our result directly.

success probability of $\mathcal{B}_0$ using Lemma A.1, as follows. A is a random variable denoting possible runs of experiments up to the challenge from $\mathcal{V}$. B_a is a random variable denoting the remainder of a run of an experiment after initial part a in the support of A. For any a in the support of A, and for any b_1 and b_2 in the support of B_a, the predicate $\mathsf{Coll}_a(b_1, b_2)$ is defined to be true if the challenges from $\mathcal{V}$ are equal in b_1 and b_2. Thus a pair (a, b) indicates a full run of the experiment, the predicate $\varphi(a, b)$ indicates success in the experiment, and the predicate $\varphi(a, b_1, b_2)$ indicates success in each sub-experiment corresponding to runs (a, b_1) and (a, b_2), with the challenges from $\mathcal{V}$ in b_1 and b_2 being distinct. Therefore $\varphi(a, b_1, b_2)$ indicates that $\mathcal{B}_0$ succeeds, and hence, by Lemma A.1, we see that $\mathcal{B}_0$ succeeds with probability at least $\varepsilon_0 = (p')^2 - 2^{-k}$. □

4. NMZK

Our general NMZK construction is similar to the SSZK construction above, but with a Σ-protocol replaced by an Ω-*protocol* (defined here) to achieve the NMZK extraction properties.

4.1. Ω-*Protocols*

An Ω-protocol $(A, B)_{[\sigma]}$ for a relation $R = \{(x, w)\}$ and CRS σ, is a Σ-protocol for relation R with the following additional properties:

1. For a given distribution ensemble $\mathcal{D}$, a common reference string σ is drawn from $\mathcal{D}_k$ and each function $a(\cdot)$, $z(\cdot)$, and $\varphi(\cdot)$ takes σ as an additional input. (Naturally, the simulator M in the definition of Σ-protocols may also take σ as an additional input.)
2. There exists a polynomial-time extractor $\mathcal{E} = (\mathcal{E}_1, \mathcal{E}_2)$ such that the reference string output by $\mathcal{E}_1(1^k)$ is statistically indistinguishable from $\mathcal{D}_k$. Furthermore, given $(\sigma, \tau) \leftarrow \mathcal{E}_1(1^k)$, if there exist two accepting conversations (a, c, z) and (a, c', z') with $c \neq c'$ for some given $x \in \hat{L}_R$, then $\mathcal{E}_2(x, \tau, (a, c, z))$ outputs w such that $(x, w) \in R$.[16]

Informally, one way to construct Ω-protocols is as follows. Our common reference string will consist of a random public key pk for a semantically secure encryption scheme. Then for a given $(x, w) \in R$, we construct an encryption e of w under key pk, and then construct a Σ-protocol to prove that there is a w such that $(x, w) \in R$ and that e is an encryption of w.[17]

As with Σ-protocols, we use the $\vee$ notation to denote an "OR" protocol, even if one or both of these protocols are Ω-protocols.

[16] Notice that this extraction property is similar to that of weak special soundness of Σ-protocols, where there exists an accepting conversation even for an invalid proof, but two accepting conversations guarantee that the proof is valid. Here, the extractor can always extract something from any conversation, but it might not be the witness if there is only one accepting conversation. However, having two accepting conversations sharing the same a guarantees that the extracted information is indeed a witness.

[17] We remark that the idea of using encryption and a ZK proof of membership together to allow extraction has been considered previously, e.g., [21].

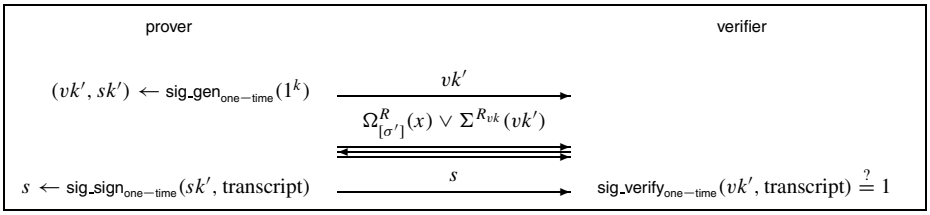

Fig. 2. $\mathsf{NM}^R_{[vk,\sigma']}(x)$: An NMZK protocol for relationship R with common input x and common reference string (vk, σ'), where σ' is drawn from the distribution associated with $\Omega^R_{[\sigma']}(x)$.

4.2. *NMZK Protocol*

Let $\Omega^R_{[\sigma']}(x)$ be an Ω-protocol for a relation R with common reference string σ' and common input x. Let $\mathsf{NM}^R_{[vk,\sigma']}(x)$ be the $\mathsf{SS}^R_{[vk]}(x)$ protocol with $\Sigma^R(x)$ replaced by $\Omega^R_{[\sigma']}(x)$. (For every σ', the resultant protocol is also a Σ-protocol.) Let $\mathcal{E}_\Omega = (\mathcal{E}_{\Omega,1}, \mathcal{E}_{\Omega,2})$ be the extractor for $\Omega^R_{[\sigma']}(x)$. The protocol $\mathsf{NM}^R_{[vk,\sigma']}(x)$ is shown in Fig. 2.

We now describe $\mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2)$ for $\mathsf{NM}^R_{[vk,\sigma']}$. $\mathcal{S}_1(1^k)$ generates signature keys $(vk, sk) \leftarrow \mathsf{sig\_gen}_{\mathsf{adap}}(1^k)$ and then sets $\sigma' \stackrel{R}{\leftarrow} \mathcal{D}_k$, where $\mathcal{D}$ is the distribution ensemble for $\Omega^R_{[\sigma']}$. Next, $\mathcal{S}_1(1^k)$ outputs $((vk, \sigma'), sk)$. $\mathcal{S}_2(sk)$ first checks that common input $x \in \hat{L}_R$. If not, it aborts. Otherwise it runs the protocol as normal, except generating $s' \leftarrow \mathsf{sig\_sign}_{\mathsf{adap}}(sk, vk')$, and using knowledge of s' to complete the protocol $\Omega^R_{[\sigma']}(x) \vee \Sigma^{R_{vk}}(vk')$.

Finally, we must describe $\mathcal{E} = (\mathcal{E}_1, \mathcal{E}_2)$ for $\mathsf{NM}^R_{[vk,\sigma']}(x)$. $\mathcal{E}_1(1^k)$ generates signature keys $(vk, sk) \leftarrow \mathsf{sig\_gen}_{\mathsf{adap}}(1^k)$, generates $(\sigma', \tau') \leftarrow \mathcal{E}_{\Omega,1}(1^k)$, and then outputs $((vk, \sigma'), sk, \tau')$. $\mathcal{E}_2(\tau')$ simply runs as $\mathcal{V}$ until $\mathcal{V}$ outputs a bit b. If $b = 1$, $\mathcal{E}_2(\tau')$ takes the conversation (a, c, z) produced by $\Omega^R_{[\sigma']}(x)$, and generates $w \leftarrow \mathcal{E}_{\Omega,2}(x, \tau', (a, c, z))$. If $b = 0$, $\mathcal{E}_2(\tau')$ sets $w \leftarrow \perp$. Then $\mathcal{E}_2(\tau')$ outputs (b, w).

Theorem 4.1. *The protocol* $\mathsf{NM}^R_{[vk,\sigma']}(x)$ *is an NMZK argument of knowledge for the relation* R.

Proof. *Completeness*: Straightforward.

Reference string indistinguishability: Straightforward.

Extractor indistinguishability: It follows from the extractor indistinguishability of $\Omega^R_{[\sigma']}(x)$.

Unbounded ZK: By inspection, $\mathcal{S}_1(1^k)$ produces exactly the same distribution as the real protocol. Then by the fact that $\mathcal{S}'(\tau)$ runs $\mathcal{S}_2(\tau)$ only when $(x, w) \in R$, and by the fact that for every σ', $\Omega^R_{[\sigma']}(x) \vee \Sigma^{R_{vk}}(vk')$ is an "OR" type of Σ-protocol, and thus witness indistinguishable, unbounded ZK follows by a straightforward hybrid argument.

Extraction: For an adversary $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, recall the experiments $\mathsf{Expt}_{\mathcal{A}}(k)$ and $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$ in the definition of NMZK. Let $p_1 = \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = 1]$ and $p_2 = \Pr[\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k) = 1]$. Our goal is to show that $|p_2 - p_1|$ is negligible.

Say a *forgery* occurs if $\mathcal{V}$ or $\mathcal{E}_2$ accepts, and the one-time verification key vk' in that session was used by $\mathcal{S}_2(\tau)$, but on a different transcript, or resulting in a different

signature. Let $\mathsf{Expt}'_{\mathcal{A}}(k)$ be $\mathsf{Expt}_{\mathcal{A}}(k)$ except that if a *forgery* occurs, the experiment halts and fails. Let $p'_1 = \Pr[\mathsf{Expt}'_{\mathcal{A}}(k) = 1]$. Similar to the proof of Theorem 3.1, we can show that $p'_1 = p_1 - c_1\varepsilon_1$, where c_1 is the number of sessions $\mathcal{A}_2$ starts with the simulator in $\mathsf{Expt}_{\mathcal{A}}(k)$, and ε_1 is negligible.

Now let $\mathsf{Expt}''_{\mathcal{A}}(k)$ be $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$ except that if a *forgery* occurs, the experiment halts and fails. As above, we can show that $p'_2 = p_2 - c_2\varepsilon_2$, where c_2 is the number of sessions $\mathcal{A}_2$ starts with the simulator in $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$, and ε_2 is negligible.

Let p'' be the probability in $\mathsf{Expt}''_{\mathcal{A}}(k)$ that $\mathcal{E}_2(\tau)$ outputs $(1, w)$ for a session with common input x, and $(x, w) \notin R$. Using the extraction property of $\Omega^R_{[\sigma']}(x)$, as in the proof of Theorem 3.1 one can show that there is a non-uniform probabilistic polynomial-time breaker $\mathcal{B}_0$ that makes at most $2c$ oracle queries and breaks $\mathsf{SIG}_{\mathsf{adap}}$ with probability at least $\varepsilon_0 = (p'')^2 - 2^k$. Thus by the existential unforgeability of $\mathsf{SIG}_{\mathsf{adap}}$, p'' is negligible.

By extractor indistinguishability again, the probability of producing output $b = 1$ with a unique transcript in $\mathsf{Expt}'_{\mathcal{A}}(k)$ and $\mathsf{Expt}''_{\mathcal{A}}(k)$ is the same, so $p'_2 = p'_1 - p''$.

Then $p_1 = p'_1 + c_1\varepsilon_1 = p'_2 + p'' + c_1\varepsilon_1 = p_2 - c_2\varepsilon_2 + c_1\varepsilon_1 + p''$, so $|p_2 - p_1| \leq c_1\varepsilon_1 + c_2\varepsilon_2 + p''$, which is negligible. □

We observe that the construction for protocol $\mathsf{NM}^R_{[vk,\sigma']}(x)$ is in fact same-string unbounded ZK, and thus we have the following:

Corollary 4.2. *The protocol* $\mathsf{NM}^R_{[vk,\sigma']}(x)$ *is a robust ZK argument of knowledge for the relation* R.

5. UCZK

First we review the framework of universal composability [10]. Then we prove that any NMZK protocol with certain simple properties can be augmented to be UCZK in the model of static corruptions. This result implies as a corollary that a slight generalization of our protocol from the previous section can be augmented to be UCZK in this model. Then we give a new construction that is UCZK in the model of adaptive corruptions.

5.1. *The Universal Composability Framework*

This framework was suggested by Canetti for defining the security and composition of protocols [10]. To define security in this framework, one first specifies an *ideal functionality*, describing the desired behavior of the protocol using a trusted party. Then one proves that a particular protocol operating in the real world securely realizes this ideal functionality, as defined below. We briefly summarize this framework:[18]

- **Communication model:** We assume an *asynchronous* network, without guaranteed delivery of messages. Further, we assume that the messages are authenticated, since authentication can be added in standard ways (i.e., the $\mathcal{F}_{\text{AUTH}}$ model in [10]).
- **Entities:** The basic entities involved are n parties $P_1, \ldots, P_n$, an adversary $\mathcal{A}$, and an environment $\mathcal{Z}$. All the entities are modeled as probabilistic interactive Turing machines.

[18] The material in this section is taken from [10], [13], and [14]; refer to these references for further detail.

- **Session IDs and sub-session IDs:** Each message also carries a *session ID* (*sid*), and if the message is for a multi-session functionality (see below), an additional *sub-session ID* (*ssid*). These IDs are used to ensure the uniqueness of the sessions. It is required that no two instances of protocols have the same ID, and this is enforced by protocols at a higher level. In other words, only when the uniqueness of *sid*/*ssid* is established is the security of the protocols guaranteed. See [10] and [14] for more discussions on IDs.
- **Corruptions:** We specify either *static* or *adaptive* corruptions, as in [10]. In the static case, the adversary corrupts parties only at the onset of the computation; in the adaptive case, the adversary chooses which parties to corrupt as the computation evolves. Once the adversary corrupts a party, it learns all its internal information, including the private input, the communication history, and the random bits used, *except* the information explicitly erased by the party before the corruption. Once they are corrupted, the behavior of the parties is arbitrary, or malicious.
- **Real-life execution:** At a high level, the execution of a protocol π, run by the parties in the presence of $\mathcal{A}$ and an environment machine $\mathcal{Z}$, with input z, is modeled as a sequence of *activations* of the entities, with $\mathcal{Z}$ activated first. When $\mathcal{Z}$ is activated, it may write messages on the other entities input tapes (and thus activate it next), and read messages from the other entities output tapes. When $\mathcal{A}$ is activated, it may read messages from a party's outgoing communication tapes, and write a message to a party's incoming communication tapes, thus activating the party. It may also corrupt parties, as discussed above. When a party is activated, it runs the protocol π. (See [10] for more detail on the exact description of all the activations.) Finally, the environment outputs one bit and halts.

  For security parameter $k \in \mathbb{N}$ and input $z \in \{0,1\}^*$, let $\text{REAL}_{\pi,\mathcal{A},\mathcal{Z}}$ denote the distribution ensemble of random variables describing $\mathcal{Z}$'s output when interacting with adversary $\mathcal{A}$ and parties running protocol π, with input z, security parameter k, and uniformly chosen random tapes for all the entities.
- **Ideal process:** The security of the protocol is defined by comparing the real execution of the protocol (as described above) to an ideal process in which an additional entity, the ideal functionality $\mathcal{F}$, is introduced; essentially, $\mathcal{F}$ is an incorruptible trusted party that is programmed to produce the desired functionality of the given task. Additionally, the parties are replaced by dummy parties, who do not communicate with each other, but instead have access to $\mathcal{F}$. In this idealized execution, again the environment is activated first, generating the inputs. Whenever a dummy party is activated, it forwards its input to $\mathcal{F}$. Let $\mathcal{S}$ denote the adversary in this idealized execution. $\mathcal{S}$ can see the destinations of the messages between the parties and $\mathcal{F}$, but not the contents. (Again, see [10] for the exact description of the activations.) As in the real-life execution, at some point the environment outputs one bit and halts.

  Let $\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}$ denote the distribution ensemble of random variables describing $\mathcal{Z}$'s output after interacting with adversary $\mathcal{S}$ in the ideal process for $\mathcal{F}$, with input z, security parameter k, and uniformly chosen random tapes for all the participating entities ($\mathcal{Z}$, $\mathcal{S}$, and $\mathcal{F}$).
- **Security:** In this framework, a protocol π *securely realizes* an ideal functionality $\mathcal{F}$ if for any real-life adversary $\mathcal{A}$ there exists an ideal-process adversary $\mathcal{S}$ such that

no environment $\mathcal{Z}$, on any input, can tell with non-negligible probability whether it is interacting with $\mathcal{A}$ and parties running π in the real-life execution, or with $\mathcal{S}$ in the ideal process for $\mathcal{F}$. More precisely, two corresponding binary distribution ensembles are *indistinguishable*, denoted $\text{REAL}_{\pi,\mathcal{A},\mathcal{Z}} \stackrel{c}{\approx} \text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}$ in [10] (meaning that for any $d \in \mathbb{N}$ there exists $k_0 \in \mathbb{N}$ such that for all $k > k_0$ and for all inputs z, $|\Pr[\,\text{REAL}_{\pi,\mathcal{A},\mathcal{Z}}(k, z)\,] - \Pr[\,\text{IDEAL}_{\mathcal{F},\mathcal{S},\mathcal{Z}}(k, z)\,]| < k^{-d}$).

- **The hybrid model:** Protocols typically would invoke other sub-protocols. The hybrid model is like a real-life execution, except that some invocations of the sub-protocols are replaced by the invocation of an instance of an ideal functionality $\mathcal{F}$; this is called the "$\mathcal{F}$-hybrid model." Specifically, the model is identical to the real-life model, with the addition that besides sending messages to each other, the parties may exchange messages with an unbounded number of copies of $\mathcal{F}$, where each copy is identified via a unique *session identifier* (*sid*). The communication between the parties and each one of these copies mimics the ideal execution.

  Let $\text{HYB}^{\mathcal{F}}_{\pi,\mathcal{A},\mathcal{Z}}$ denote the distribution ensemble of random variables describing the output of $\mathcal{Z}$, after interacting with $\mathcal{A}$ and parties running protocol π in the $\mathcal{F}$-hybrid model. Now let ρ be a protocol that securely realizes $\mathcal{F}$. The composed protocol π^ρ is constructed by replacing the first message to $\mathcal{F}$ in π by an invocation of a new copy of ρ, with fresh random input, the same *sid*, and with the contents of that message as input; each subsequent message to that copy of $\mathcal{F}$ is replaced with an activation of the corresponding copy of ρ, with the contents of that message as new input to ρ.
- **The composition theorem:** The composition theorem [10] basically says that if ρ securely realizes $\mathcal{F}$ in the $\mathcal{G}$-hybrid model, for some functionality $\mathcal{G}$, then an execution of the composed protocol π^ρ, running in the $\mathcal{G}$-hybrid model, "emulates" an execution of protocol π in the $\mathcal{F}$-hybrid model. That is, no environment machine $\mathcal{Z}$ can distinguish whether it is interacting with $\mathcal{A}$ and π^ρ in the $\mathcal{G}$-hybrid model, or it is interacting with $\mathcal{S}$ and π in the $\mathcal{F}$-hybrid model. In other words, $\text{HYB}^{\mathcal{G}}_{\pi^\rho,\mathcal{A},\mathcal{Z}} \stackrel{c}{\approx} \text{HYB}^{\mathcal{F}}_{\pi,\mathcal{S},\mathcal{Z}}$.

The ZK functionality. We now recall the ideal ZK functionality [10]. As a convention, all the messages from the parties to the ideal functionality take the form (action, sid, ...), where action is in lower case, and all messages from the ideal functionality take the form (OBJECT, sid, ...), where OBJECT is in upper case. The functionality is given in Fig. 3. In the functionality, parameterized by a relation R, the prover sends to the functionality the input x together with a witness w. If $R(x, w)$ holds, then the functionality forwards x to the verifier.[19] As pointed out in [10], this is actually a proof of knowledge in that the verifier is assured that the prover actually knows w.

One shortcoming of the above formulation is that we will be designing and analyzing protocols in the common reference string model, and so they will be operating in the $\mathcal{F}^{\mathcal{D}}_{\text{CRS}}$-hybrid model, where $\mathcal{F}^{\mathcal{D}}_{\text{CRS}}$ is the functionality that, for a given security parameter k, chooses a string from distribution $\mathcal{D}_k$ and hands it to all parties. However, directly realizing $\mathcal{F}^{R}_{\text{ZK}}$ in the $\mathcal{F}^{\mathcal{D}}_{\text{CRS}}$-hybrid model and using the universal composition theorem

[19] As in [13], we assume there is a symbol $\perp$ such that for any relation R and any string x, $(x, \perp) \notin R$.

Functionality $\mathcal{F}_{\text{ZK}}^{R}$

$\mathcal{F}_{\text{ZK}}^{R}$ proceeds as follows, running with security parameter k, a prover P_i, a verifier P_j, and an adversary $\mathcal{S}$:

- Upon receiving (zk-prover, sid, P_i, P_j, x, w) from P_i: If $R(x, w)$ then send (ZK-PROOF, sid, P_i, P_j, x) to P_j and $\mathcal{S}$ and halt. Otherwise, ignore.

Fig. 3. The ZK functionality (for relation R).

would result in a composed protocol where a new instance of the reference string is needed for each proof, which (1) is extremely inefficient, and (2) does not reflect the notion of the CRS model, where an unbounded number of protocol instances would use the same copy of the string. Canetti and Rabin [14] suggested the following notion to cope with this problem:

- **Universal composition with joint state:** Let $\mathcal{F}$ and $\mathcal{G}$ be ideal functionalities, and let $\hat{\mathcal{F}}$ denote the "multi-session extension of $\mathcal{F}$," in that $\hat{\mathcal{F}}$ will run multiple copies of $\mathcal{F}$, where each copy is identified by a special *sub-session identifier* (*ssid*). Now let π be a protocol in the $\mathcal{F}$-hybrid model, and let $\hat{\rho}$ be a protocol that securely realizes $\hat{\mathcal{F}}$ in the $\mathcal{G}$-hybrid model. Then construct the composed protocol $\pi^{[\hat{\rho}]}$ by replacing all the copies of $\mathcal{F}$ in π by a single copy of $\hat{\rho}$. The universal composition with joint state theorem states that $\pi^{[\hat{\rho}]}$, running in the $\mathcal{G}$-hybrid model, correctly emulates π in the $\mathcal{F}$-hybrid model.

The definition of $\hat{\mathcal{F}}_{\text{ZK}}^{R}$, the multi-session extension of $\mathcal{F}_{\text{ZK}}^{R}$, is shown in Fig. 4. Note the two types of indices: the *sid*, which, as before, differentiates messages to $\hat{\mathcal{F}}_{\text{ZK}}^{R}$ from messages sent to other functionalities, and *ssid*, the sub-session ID, which is unique per input message (or proof).

5.2. *NMZK Implies UCZK*

Let Π be an NMZK protocol between a prover and verifier. We say Π is *augmentable* if the prover sends the first message, and this message contains the common input x, along with auxiliary data aux that may contain any arbitrary public values. The reason for aux is discussed below.

First note that the addition of an auxiliary data field is not a trivial modification: since the auxiliary field becomes an integral part of the proof transcript, the non-malleability property should apply to this field as well. In other words, the adversary should not be able to produce a new proof even in the case where the *only* difference between the new proof and the proof the adversary sees is the auxiliary field. We do not know if,

Functionality $\hat{\mathcal{F}}_{\text{ZK}}^{R}$

$\hat{\mathcal{F}}_{\text{ZK}}^{R}$ proceeds as follows, running with security parameter k, parties $P_1, \ldots, P_n$, and an adversary $\mathcal{S}$:

- Upon receiving (zk-prover, sid, $ssid$, P_i, P_j, x, w) from P_i: If $R(x, w)$ then send (ZK-PROOF, sid, $ssid$, P_i, P_j, x) to P_j and $\mathcal{S}$ and halt. Otherwise, ignore.

Fig. 4. The multi-session zero-knowledge functionality (for relation R).

in general, an NMZK protocol is augmentable. However, it is not hard to show that the construction of NMZK protocols in Section 4 is augmentable: one simply includes the auxiliary field in the strong one-time signature.

We will show how to augment Π with additional information in each message to allow it to be used between two parties in the universal composability framework. This augmented protocol is denoted $\hat{\Pi}$, and is constructed as follows.

For an instance of $\hat{\Pi}$ run between parties P_i and P_j, set aux to $(ssid, P_i, P_j)$, where $ssid$ is defined in the previous section, P_i is the identity of the prover, and P_j is the identity of the verifier.[20] Then the ℓth prover message is formatted as $(\mathsf{prv}_\ell, sid, ssid, P_i, \mathsf{prv\text{-}data}_\ell)$, where prv_ℓ is the label for the ℓth prover message, and $\mathsf{prv\text{-}data}_\ell$ is the data field containing the ℓth message sent by the prover in Π. Analogously, the ℓth verifier message is formatted as $(\mathsf{ver}_\ell, sid, ssid, P_j, \mathsf{ver\text{-}data}_\ell)$, where ver_ℓ is the label for the ℓth verifier message, and $\mathsf{ver\text{-}data}_\ell$ is the data field containing the ℓth message sent by the verifier in Π. Finally, before accepting, the verifier checks that aux corresponds to the values $(ssid, P_i, P_j)$ outside the prover data field, and that aux was not used previously.

Here we show that the augmented NMZK protocol $\hat{\Pi}$ is a UCZK protocol, assuming static corruptions.[21]

Theorem 5.1. *Let $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S}_\Pi = (\mathcal{S}_{\Pi,1}, \mathcal{S}_{\Pi,2}), \mathcal{E}_\Pi = (\mathcal{E}_{\Pi,1}, \mathcal{E}_{\Pi,2}))$ be an augmentable NMZK protocol for a relation R. Then the augmented protocol $\hat{\Pi}$ securely realizes functionality $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$ in the $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$-hybrid model, assuming static corruptions.*

Proof. Let $\mathcal{A}$ be an adversary that operates against protocol $\hat{\Pi}$ in the $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$-hybrid model. We construct an ideal process adversary (i.e., a simulator) $\mathcal{S}$ such that no environment $\mathcal{Z}$ can tell whether it is interacting with $\mathcal{A}$ and $\hat{\Pi}$ in the $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$-hybrid model, or with $\mathcal{S}$ in the ideal process for $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$.

For simplicity, we assume only one copy of $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$ is accessed by $\mathcal{Z}$. Obviously we could duplicate the actions of $\mathcal{S}$ for each copy of $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$ (differentiated by the sid value).

Simulator $\mathcal{S}$ generates $(\sigma, \tau_1, \tau_2) \leftarrow \mathcal{E}_{\Pi,1}(1^k)$, uses σ as the common reference string for $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$, and stores τ_1 and τ_2.

Simulator $\mathcal{S}$ runs a simulated copy of $\mathcal{A}$. Messages received from $\mathcal{Z}$ are forwarded to the simulated $\mathcal{A}$, and messages sent by the simulated $\mathcal{A}$ to its environment are forwarded to $\mathcal{Z}$.

If $\mathcal{S}$ receives a message $(\mathsf{ZK\text{-}PROOF}, sid, ssid, P_i, P_j, x)$ from $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$, i.e., P_i is uncorrupted and wishes to perform a ZK proof for common input x, then $\mathcal{S}$ simulates P_i in $\hat{\Pi}$. In particular, $\mathcal{S}$ sets the prover data field in the messages of P_i using protocol $\mathcal{S}_{\Pi,2}(\tau_1)$. If P_j is also uncorrupted, then $\mathcal{S}$ simulates P_j in $\hat{\Pi}$, setting the verifier data field in the messages of P_j using the actual verifier protocol. In this case, when the simulated P_j receives

[20] This auxiliary data aux is necessary since NMZK allows copying proofs exactly, but the UCZK functionality does not, and thus we need some way to make every proof distinct.

[21] The problem with handling adaptive corruptions is when a prover's first message is simulated, and then that prover gets corrupted. Given our simulation, it seems difficult to provide the adversary with internal prover data that would correctly correspond to its first message.

the final message from the simulated P_i, $\mathcal{S}$ forwards (ZK-PROOF, *sid*, *ssid*, P_i, P_j, x) to the actual uncorrupted P_j.

If $\mathcal{A}$, controlling a corrupted party P_i, starts an interaction as a prover with an uncorrupted party P_j using *ssid*, then $\mathcal{S}$ learns common input x (since it is included in the first message) and simulates P_j in $\hat{\Pi}$. In particular, it sets the verifier data field in the messages of P_j using protocol $\mathcal{E}_{\Pi,2}(\tau_2)$. At the end of the interaction $\mathcal{E}_{\Pi,2}(\tau_2)$ will output (b, w). If $b = 1$, $\mathcal{S}$ sends (zk-prover, *sid*, *ssid*, P_i, P_j, x, w) to $\hat{\mathcal{F}}_{\text{ZK}}^{R}$; otherwise, it sends nothing. Then it forwards any response from $\hat{\mathcal{F}}_{\text{ZK}}^{R}$ to P_j.

Now we show that $\text{HYB}_{\hat{\Pi},\mathcal{A},\mathcal{Z}}^{\mathcal{F}_{\text{CRS}}^{\mathcal{D}}} \overset{c}{\approx} \text{IDEAL}_{\hat{\mathcal{F}}_{\text{ZK}}^{R},\mathcal{S},\mathcal{Z}}$.

First we define a new experiment $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$. The new experiment runs simulated copies of $\mathcal{Z}$ and $\mathcal{A}$. Messages received from $\mathcal{Z}$ are forwarded to the simulated $\mathcal{A}$, and messages sent by the simulated $\mathcal{A}$ to its environment are forwarded to $\mathcal{Z}$. The simulator for Π, $\mathcal{S}_{\Pi,1}(1^k)$ is run to produce (σ, τ), and queries to $\mathcal{F}_{\text{CRS}}^{\mathcal{D}}$ are answered with σ. If an uncorrupted party P_i receives input (zk-prover, *sid*, *ssid*, P_i, P_j, x, w) from $\mathcal{Z}$ with $(x, w) \in R$ it sets the prover data field of its messages by running protocol $\mathcal{S}_{\Pi,2}(\tau)$ with reference string σ, and common input x. An uncorrupted party P_j responds to a prover as in the actual verifier protocol in $\hat{\Pi}$. The output of each experiment is the output of $\mathcal{Z}$.

Let $\text{MIX}_{\mathcal{A},\mathcal{Z}}$ denote the distribution ensemble of random variables describing the outputs of $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$.

By the unbounded ZK property, we have $\text{HYB}_{\hat{\Pi},\mathcal{A},\mathcal{Z}}^{\mathcal{F}_{\text{CRS}}^{\mathcal{D}}} \overset{c}{\approx} \text{MIX}_{\mathcal{A},\mathcal{Z}}$. To see this, note that we could construct an adversary $\mathcal{A}'$ that takes a reference string and runs the protocol $\hat{\Pi}$, except that $\mathcal{A}'$ calls a protocol wrapper with label $\mathsf{aux} = (\textit{ssid}, P_i, P_j)$ when simulating uncorrupted parties acting as provers. If the wrapper contains an actual prover, then the distribution of outputs of $\mathcal{A}'$ will be the same as $\text{HYB}_{\hat{\Pi},\mathcal{A},\mathcal{Z}}^{\mathcal{F}_{\text{CRS}}^{\mathcal{D}}}$, and if the wrapper contains a simulator, then the distribution will be the same as $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$.

Now we must show that $\text{MIX}_{\mathcal{A},\mathcal{Z}} \overset{c}{\approx} \text{IDEAL}_{\hat{\mathcal{F}}_{\text{ZK}}^{R},\mathcal{S},\mathcal{Z}}$. This will follow from the *unbounded extraction* property (which is simply a generalized extraction property for NMZK implied by the standard extraction property, see Definition 5.2 and Lemma 5.3). Say that the two distributions can be distinguished with probability $\gamma(k)$. Since both $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ and $\mathcal{S}$ run the same simulation for the prover, and the output messages of the extractor run by $\mathcal{S}$ are perfectly indistinguishable from the output messages of the verifier, the only difference comes from when the extractor outputs an incorrect witness for a session started by $\mathcal{A}$, and thus $\mathcal{Z}$ receives an output message (indicating a correct proof) in $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ but not when interacting with $\mathcal{S}$. (Note that the transcripts of corrupted prover/uncorrupted verifier sessions will never be the same as transcripts of uncorrupted prover/corrupted verifier sessions because of the auxiliary data aux.) Let $\vec{b}$ be the vector corresponding to simulated verifier sessions, with $b = 1$ corresponding to whether $\mathcal{Z}$ receives an output message. Then the statistical difference between the distribution of vectors $\vec{b}$ resulting from $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ and vectors $\vec{b}$ resulting from $\mathcal{S}$ is at least $\gamma(k)$.

Now we construct an NMZK adversary $\mathcal{A}'$ that takes a reference string and runs $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ except that it uses the given reference string instead of generating a new one, and that it calls a "simulator" protocol wrapper when simulating uncorrupted parties acting as provers with corrupted verifiers, and a "verifier" protocol wrapper when simu-

lating uncorrupted parties acting as verifiers with corrupted provers. Then in $\mathsf{Expt}_{\mathcal{A}'}(k)$, the vector $\vec{b}$ will have the same distribution as the one resulting from $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$. On the other hand, in $\mathsf{Expt}^{\mathcal{E}_\Pi}_{\mathcal{A}'}(k)$, the vector $\vec{b}$ will have the same distribution as the one resulting from $\mathcal{S}$, up until $\mathcal{Z}$ receives an output message in $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ that would not have appeared in $\mathcal{S}$. It should be clear that the distributions of $\vec{b}$ in the two experiments are statistically distinguishable with the same probability as the distributions of $\vec{b}$ resulting from $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ and $\mathcal{S}$, i.e., $\gamma(k)$. By the unbounded extraction property, $\gamma(k)$ is negligible. □

We now define *unbounded extraction* NMZK, and show that it is implied by our standard NMZK definition. This notion was used in the previous proof.

Definition 5.2 (Unbounded-Extraction NMZK Proof/Argument of Knowledge). $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2), \mathcal{E} = (\mathcal{E}_1, \mathcal{E}_2))$ is an *unbounded-extraction NMZK proof* (resp., *argument*) *of knowledge system* for an NP language L with witness relation R if Π is an NMZK proof (resp. argument) system for L and, furthermore,

Unbounded Extraction. For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A} = (\mathcal{A}_1, \mathcal{A}_2)$, where $\mathcal{A}_1$ and $\mathcal{A}_2$ are coordinated machines, $\sum_{\vec{b}\in\{0,1\}^*} |\Pr[\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k) = \vec{b}] - \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = \vec{b}]|$ is negligible, where the experiments $\mathsf{Expt}_{\mathcal{A}}(k)$ and $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$ are defined as follows:

$\mathsf{Expt}_{\mathcal{A}}(k)$:	$\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$:
$(\sigma, \tau_1) \leftarrow \mathcal{S}_1(1^k)$	$(\sigma, \tau_1, \tau_2) \leftarrow \mathcal{E}_1(1^k)$
$(\vec{x}, \vec{tr}, \vec{b}) \leftarrow \big(\langle \boxed{\mathcal{S}''(\tau_1)}, \mathcal{A}_1\rangle, \langle \mathcal{A}_2, \boxed{\mathcal{V}}\rangle\big)_{[\sigma]}$	$(\vec{x}, \vec{tr}, (\vec{b}, \vec{w})) \leftarrow \big(\langle \boxed{\mathcal{S}''(\tau_1)}, \mathcal{A}_1\rangle, \langle \mathcal{A}_2, \boxed{\mathcal{E}_2(\tau_2)}\rangle\big)_{[\sigma]}$
Let Q be the set of transcripts of machines in $\boxed{\mathcal{S}''(\tau_1)}$.	Let Q be the set of transcripts of machines in $\boxed{\mathcal{S}''(\tau_1)}$.
For all i,	For all i,
if $\exists tr' \in Q, tr[i] \bowtie tr'$	if $(x[i], w[i]) \notin R$ or $\exists tr' \in Q, tr[i] \bowtie tr'$
then $b[i] \leftarrow 0$	then $b[i] \leftarrow 0$
Return $\vec{b}$	Return $\vec{b}$

where $\mathcal{S}''(\tau)$ runs as follows on common reference string σ, common input x and private input w: $\mathcal{S}''(\tau)$ runs $\mathcal{S}_2(\tau)$ on common reference string σ and common input x.

Lemma 5.3. *Let $\Pi = (\mathcal{D}, \mathcal{P}, \mathcal{V}, \mathcal{S} = (\mathcal{S}_1, \mathcal{S}_2), \mathcal{E} = (\mathcal{E}_1, \mathcal{E}_2))$ be an NMZK protocol for a relation R. Then Π is an unbounded extraction NMZK protocol for R.*

Proof. First notice that since $\mathcal{V}$ and $\mathcal{E}_2(\tau_2)$ have exactly the same behavior, there will be an exact correspondence of vectors returned in the two experiments, except that in some cases some bits that were 1 in $\mathsf{Expt}_{\mathcal{A}}(k)$ would be 0 in $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$. Let $\beta_{\mathcal{A}}(k) = \sum_{\vec{b}\in\{0,1\}^*} |\Pr[\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k) = \vec{b}] - \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = \vec{b}]|$. Now we perform a hybrid argument. Let $\mathsf{Expt}^{\mathcal{E},j}_{\mathcal{A}}(k)$ be the same as $\mathsf{Expt}^{\mathcal{E}}_{\mathcal{A}}(k)$ except that for the first condition $((x[i], w[i]) \notin$

R), "For all i" is replaced with "For all $i \leq j$." Let ℓ denote the maximum number of sessions of $\mathcal{E}_2(\tau_2)$ started by $\mathcal{A}_2$, and notice that ℓ is polynomial in k. Then $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},0}(k)$ is the same as $\mathsf{Expt}_{\mathcal{A}}(k)$ and $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},\ell}(k)$ is the same as $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E}}(k)$. By a telescoping argument,

$$\sum_{j=1}^{\ell} \sum_{\vec{b}\in\{0,1\}^*} |\Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j}(k) = \vec{b}] - \Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j-1}(k) = \vec{b}]| \geq \beta_{\mathcal{A}}(k).$$

Now let $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j,1}(k)$ be the same as $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E}}(k)$ except that for the first condition $((x[i], w[i]) \notin R)$, "For all i" is replaced with "For $i = j$." Because $\mathcal{V}$ and $\mathcal{E}_2(\tau_2)$ have exactly the same behavior, and thus the output for any machine not equal to j in $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j-1}(k)$ and $\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j}(k)$ would be the same, it is easy to verify that

$$\begin{aligned}
&\sum_{j=1}^{\ell} \sum_{v\in\{0,1\}^*} |\Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j,1}(k) = v] - \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = v]| \\
&\qquad = \sum_{j=1}^{\ell} \sum_{\vec{b}\in\{0,1\}^*} |\Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j}(k) = \vec{b}] - \Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j-1}(k) = \vec{b}]| \\
&\qquad \geq \beta_{\mathcal{A}}(k).
\end{aligned}$$

Now consider a new adversary $\mathcal{A}' = (\mathcal{A}_1, \mathcal{A}_2')$ that chooses $j \in \{1, \ldots, \ell\}$ randomly, where $\mathcal{A}_2'$ runs $\mathcal{A}_2$ but simulates $\mathcal{V}$ in all but the jth session. In the jth session it calls the one-time wrapper given to it. From the definition of NMZK, $|\Pr[\mathsf{Expt}_{\mathcal{A}'}^{\mathcal{E}}(k) = 1] - \Pr[\mathsf{Expt}_{\mathcal{A}'}(k) = 1]| \leq \alpha(k)$, and by the analysis above,[22]

$$\begin{aligned}
&|\Pr[\mathsf{Expt}_{\mathcal{A}'}^{\mathcal{E}}(k) = 1] - \Pr[\mathsf{Expt}_{\mathcal{A}'}(k) = 1]| \\
&\qquad = \tfrac{1}{2}|\Pr[\mathsf{Expt}_{\mathcal{A}'}^{\mathcal{E}}(k) = 1] - \Pr[\mathsf{Expt}_{\mathcal{A}'}(k) = 1]| \\
&\qquad\quad + |\Pr[\mathsf{Expt}_{\mathcal{A}'}^{\mathcal{E}}(k) = 0] - \Pr[\mathsf{Expt}_{\mathcal{A}'}(k) = 0]| \\
&\qquad = \frac{1}{2\ell} \sum_{j=1}^{\ell} \sum_{b\in\{0,1\}} |\Pr[\mathsf{Expt}_{\mathcal{A}'}^{\mathcal{E}}(k) = b | \mathcal{A}' \text{ chooses } j] \\
&\qquad\qquad\qquad - \Pr[\mathsf{Expt}_{\mathcal{A}'}(k) = b | \mathcal{A}' \text{ chooses } j]| \\
&\qquad = \frac{1}{2\ell} \sum_{j=1}^{\ell} \sum_{v\in\{0,1\}^*} |\Pr[\mathsf{Expt}_{\mathcal{A}}^{\mathcal{E},j,1}(k) = v] - \Pr[\mathsf{Expt}_{\mathcal{A}}(k) = v]| \\
&\qquad \geq \frac{\beta_{\mathcal{A}}(k)}{\ell},
\end{aligned}$$

so $\beta_{\mathcal{A}}(k) \leq 2\ell \cdot \alpha(k)$. The theorem follows. □

We say a protocol $\hat{\Pi}$ is a *UCZK protocol* for R if it securely realizes functionality $\hat{\mathcal{F}}_{\mathrm{ZK}}^{R}$ in the $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$-hybrid model, for some $\mathcal{D}$.

[22] Note that we use the fact that $|\sum_k V_k| = \sum_k |V_k|$ when all V_k have the same sign.

Corollary 5.4. *Let* Π *be protocol* $\mathsf{NM}^R_{[vk,\sigma']}(x)$ *from Fig.* 2 *with the addition of the common input* x *and* $\mathsf{aux} = (\text{ssid}, P_i, P_j)$ *in the first message. Then the augmented protocol* $\hat{\Pi}$ *is a UCZK protocol for* R*, assuming static corruptions.*

5.3. *UCZK*: *Adaptive Corruptions*

To deal with adaptive corruption, we apply a technique proposed by Damgård [18] and Jarecki and Lysyanskaya [37] in which a *trapdoor commitment* is used to commit to the first message of a Σ-protocol, and then this commitment is opened when sending the third message. Informally, a trapdoor commitment is a commitment scheme with the additional property that there is a secret trapdoor such that knowing the trapdoor allows a committer to decommit to an arbitrary value. More precisely, $\mathsf{TC} = (\mathsf{TCgen}, \mathsf{TCcom}, \mathsf{TCver}, \mathsf{TCkeyver}, \mathsf{TCfake})$ is a trapdoor commitment scheme if it satisfies the properties of completeness, binding, perfect secrecy, and trapdoorness. The first three properties are the same as in any unconditionally hiding commitment scheme. The trapdoor property says (informally) that $\mathsf{TCgen}(1^k)$ outputs a secret key (the trapdoor) along with the public key, and that using this secret key and a commitment/decommitment pair (c, d) associated with a value v (i.e., $(c, d) \leftarrow \mathsf{TCcom}(pk, v)$), the function TCfake can for any value v' output a decommitment d' that is a valid decommitment of c resulting in v' (i.e., $\mathsf{TCver}(pk, c, v', d') = 1$).

However, this technique alone does not seem to yield a UCZK protocol for adaptive corruption. There are two problems remaining. First, it does not yield a non-rewinding witness extractor, which is needed for UCZK. Second, in the setting of UCZK, an ideal adversary $\mathcal{S}$ might use the trapdoor to "cheat," i.e., to decommit to arbitrary values, while at the same time it still needs the binding property for the real-life adversary $\mathcal{A}$. A "plain" trapdoor commitment scheme does not provide such a guarantee.

We solve these two problems by (1) using an Ω-protocol in the place of the Σ-protocol; recall that Ω-protocols allow for non-rewinding extractors, and (2) introducing a stronger type of trapdoor commitment scheme, which we call a *simulation-sound trapdoor commitment* (SSTC) scheme.[23] Roughly speaking, an SSTC scheme is a trapdoor commitment scheme with an extra input *id* to the commitment protocol, which guarantees that a commitment made by the adversary using input *id* is binding, even if the adversary has seen any commitment using input *id* opened (using a simulator that knows a trapdoor) once to any arbitrary value, and, moreover, any commitment using $id' \neq id$ opened (again using the simulator) an unbounded number of times to any arbitrary values. Such a trapdoor commitment scheme enables an ideal adversary to "cheat" while maintaining the binding property for the real-life adversary. We shall see that when we apply these two solutions, the protocol becomes universally composable with respect to adaptive corruption.[24]

[23] Universally composable commitments [11], [13] would also suffice, and can be constructed using trapdoor permutations. However, this construction is not as efficient as the SSTC scheme in this paper.

[24] As a technical note, we comment that on the face, this construction does not use the technique of adding a proof of knowledge of signature, as in previous constructions. However, such a technique will be used in the construction of the SSTC schemes.

Here we formally define an SSTC scheme, building on the formalization for trapdoor commitment schemes by Reyzin [52].

Definition 5.5 (Simulation-Sound Trapdoor Commitment (SSTC) Scheme). $\mathsf{TC} = (\mathsf{TCgen}, \mathsf{TCcom}, \mathsf{TCver}, \mathsf{TCkeyver}, \mathsf{TCfake})$ is an *SSTC scheme* if TCgen, TCcom, TCver, $\mathsf{TCkeyver}$, and TCfake are probabilistic polynomial-time algorithms such that

Completeness. For all id, all $k > 0$, and for all values v,

$$\Pr[(pk, sk) \stackrel{R}{\leftarrow} \mathsf{TCgen}(1^k); (c, d) \stackrel{R}{\leftarrow} \mathsf{TCcom}(pk, v, id): \mathsf{TCkeyver}(pk, 1^k) = \mathsf{TCver}(pk, c, v, id, d) = 1] = 1.$$

Simulation-Sound Binding. For all non-uniform probabilistic polynomial-time adversaries $\mathcal{A}$, there is a negligible function $\alpha(k)$ such that for all $k > 0$,

$$\Pr[(pk, sk) \stackrel{R}{\leftarrow} \mathsf{TCgen}(1^k); (c, id, v_1, v_2, d_1, d_2) \stackrel{R}{\leftarrow} \langle \mathcal{S}(sk), \mathcal{A} \rangle(pk): (\mathsf{TCver}(pk, c, v_1, id, d_1) = \mathsf{TCver}(pk, c, v_2, id, d_2) = 1) \wedge (v_1 \neq v_2) \wedge id \notin Q] \leq \alpha(k),$$

where $\mathcal{S}(sk)$ operates as follows, with Q initially set to $\emptyset$:

- On input (commit, v, id):
  compute $(c, d) \leftarrow \mathsf{TCcom}(pk, v, id)$, store (c, v, id, d), and return c.
- On input $(\mathsf{decommit}, c, v')$:
  if for some v, id, d a tuple (c, v, id, d) is stored, compute $d' \leftarrow \mathsf{TCfake}(pk, sk, c, v, id, d, v')$. If some previous $(\mathsf{decommit}, c, *)$ has been input, add id to Q. Return d'.

Hiding. For all $k > 0$ and all pk such that $\mathsf{TCkeyver}(pk, 1^k) = 1$, for all id, and for all v_1, v_2 of equal length, the following probability distributions are identical:

$$\{(c_1, d_1) \stackrel{R}{\leftarrow} \mathsf{TCcom}(pk, v_1, id): c_1\} \quad \text{and} \quad \{(c_2, d_2) \stackrel{R}{\leftarrow} \mathsf{TCcom}(pk, v_2, id): c_2\}.$$

Trapdoor Property. For all $k > 0$ and all (pk, sk) generated with non-zero probability by $\mathsf{TCgen}(1^k)$, for all id, and for all v_1, v_2 of equal length, the following probability distributions are identical:

$$\{(c, d_1) \stackrel{R}{\leftarrow} \mathsf{TCcom}(pk, v_1, id); d_2' \stackrel{R}{\leftarrow} \mathsf{TCfake}(pk, sk, c, v_1, id, d_1, v_2) : (c, d_2')\}$$

and

$$\{(c, d_2) \stackrel{R}{\leftarrow} \mathsf{TCcom}(pk, v_2, id) : (c, d_2)\}.$$

(In particular, faked commitments are correct.)

We remark that an SSTC scheme is similar to an ID-based trapdoor commitment scheme as defined in [5], but where multiple adaptively chosen IDs may be used. Also, building on this paper, a simpler definition of an SSTC scheme (that also can be used to construct UCZK proofs secure against adaptive corruptions) is given in [45].

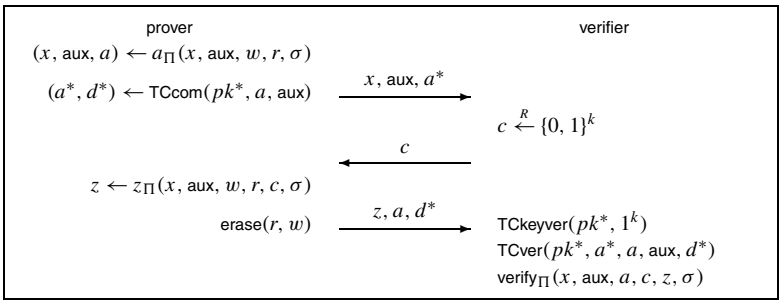

Fig. 5. $\mathsf{UC}^R_{[pk^*,\sigma]}(x; \mathsf{aux})$: A UCZK protocol for R with common reference string (pk^*, σ) drawn from $\mathcal{D}_{pk}(\mathsf{TC}) \times \mathcal{D}_\sigma(\Omega^R)$, common input x, and auxiliary input aux where $\Pi = \Omega^R_\sigma(x; \mathsf{aux})$.

Now, let Π be an augmentable NMZK protocol with common input x, auxiliary input aux, prover random bits r, and common reference string σ. As for Σ-protocols, we use the notation $a_\Pi(\cdot)$, $z_\Pi(\cdot)$, and $\mathsf{verify}_\Pi(\cdot)$ to denote the algorithms for computing the two messages of the prover, and for verifying the proof, respectively. Using this notation, the protocol $\mathsf{UC}^R_{[pk^*,\sigma]}(x; \mathsf{aux})$ is shown in Fig. 5.

Theorem 5.6. *Let Π be an Ω-protocol $\Omega^R_{[pk^*,\sigma]}(x; \mathsf{aux})$, where $\mathsf{aux} = (\mathrm{ssid}, P_i, P_j)$. Then the augmented protocol $\hat{\Pi}$ securely realizes functionality $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$ in the $\mathcal{F}^{\mathcal{D}}_{\mathrm{CRS}}$-hybrid model where erasing is allowed, assuming adaptive corruptions.*

Proof. Let $\mathcal{A}$ be an adversary that operates against protocol $\hat{\Pi}$ in the $\mathcal{F}^{\mathcal{D}}_{\mathrm{CRS}}$-hybrid model. We construct an ideal process adversary $\mathcal{S}$ such that no environment $\mathcal{Z}$ can tell whether it is interacting with $\mathcal{A}$ and $\hat{\Pi}$ in the $\mathcal{F}^{\mathcal{D}}_{\mathrm{CRS}}$-hybrid model, or with $\mathcal{S}$ in the ideal process for $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$.

For simplicity, we assume only one copy of $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$ is accessed by $\mathcal{Z}$. Obviously we could duplicate the actions of $\mathcal{S}$ for each copy of $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$ (differentiated by the *sid* value).

Formally, let Π be an Ω-protocol with simulator $\mathcal{S}_\Pi$ and extractor $\mathcal{E}_\Pi = (\mathcal{E}_{\Pi,1}, \mathcal{E}_{\Pi,2})$.

At the beginning of the ideal process, the ideal adversary $\mathcal{S}$ generates $(\sigma, \tau) \leftarrow \mathcal{E}_{\Pi,1}(1^k)$, generates $(pk^*, sk^*) \stackrel{R}{\leftarrow} \mathsf{TCgen}(1^k)$, uses (pk^*, σ) as the common reference string for $\mathcal{F}^{\mathcal{D}}_{\mathrm{CRS}}$, and stores sk^* and τ.

During the ideal process, $\mathcal{S}$ runs a simulated copy of $\mathcal{A}$. Messages received from $\mathcal{Z}$ are forwarded to the simulated $\mathcal{A}$, and messages sent by the simulated $\mathcal{A}$ to its environment are forwarded to $\mathcal{Z}$.

If $\mathcal{S}$ receives a message (ZK-PROOF, *sid*, *ssid*, P_i, P_j, x) from $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$, i.e., P_i is uncorrupted and has given a witness w to $\hat{\mathcal{F}}^R_{\mathrm{ZK}}$ such that $(x, w) \in R$, then $\mathcal{S}$ simulates P_i in $\hat{\Pi}$. In particular, $\mathcal{S}$ sets the prover data field in the first message of P_i by generating a commitment (as in the actual prover protocol) to an arbitrary string $\hat{a}$ with appropriate length (say, $\hat{a} = 0^l$, where l is the size of field "a" in the output of $a_\Pi(\cdot)$). More precisely, $\mathcal{S}$ invokes $(\hat{a}^*, \hat{d}^*) \leftarrow \mathsf{TCcom}(pk^*, \hat{a}, \mathsf{aux})$ and sends $(x, \mathsf{aux}, \hat{a}^*)$ to P_j as the first message. After receiving the challenge (as the second message) c from P_j, $\mathcal{S}$ invokes the simulator $\mathcal{S}_\Pi$ and obtains $(a, c, z) = M_\Pi(x, \sigma, c)$. Then $\mathcal{S}$ fakes a decommitment for

a by invoking $d^* = \mathsf{TCfake}(pk^*, sk^*, \hat{a}^*, \hat{a}, \mathsf{aux}, \hat{d}^*, a)$, and sends (z, a, d^*) to P_j as the final message. If P_i is corrupted before receiving a challenge, then the witness w is revealed. In this case, $\mathcal{S}$ invokes the actual first-message function a_Π to produce the first message a, instead of using the simulator $\mathcal{S}_\Pi$. Again, $\mathcal{S}$ fakes a decommitment in this case.

If P_j is also uncorrupted, then $\mathcal{S}$ simulates P_j in $\hat{\Pi}$, setting the verifier data field in the message of P_j (in particular, the random challenge) using the actual verifier protocol. In this case, when the simulated P_j receives the final message from the simulated P_i, $\mathcal{S}$ forwards ($\mathsf{ZK\text{-}PROOF}$, sid, $ssid$, P_i, P_j, x) to the actual uncorrupted P_j.

If $\mathcal{A}$, controlling a corrupted party P_i, starts an interaction as a prover with an uncorrupted party P_j using $ssid$, then $\mathcal{S}$ learns common input x (since it is included in the first message) and simulates P_j (as the verifier) in $\hat{\Pi}$. More precisely, it fills the verifier data field with a random challenge c, receives as the final message (z, a, d^*) from $\mathcal{A}$, and verifies the messages. At the end of the interaction, if all the verifications pass, the extractor $\mathcal{E}_{\Pi,2}(x, \tau, (a, c, z))$ will be invoked and output a witness w. If $R(x, w) = 1$, $\mathcal{S}$ sends ($\mathsf{zk\text{-}prover}$, sid, $ssid$, P_i, P_j, x, w) to $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$; otherwise, it sends nothing. Then it forwards any response from $\hat{\mathcal{F}}_{\mathrm{ZK}}^R$ to P_j.

Now we show that

$$\mathrm{HYB}_{\hat{\Pi},\mathcal{A},\mathcal{Z}}^{\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}} \overset{c}{\approx} \mathrm{IDEAL}_{\hat{\mathcal{F}}_{\mathrm{ZK}}^R,\mathcal{S},\mathcal{Z}},$$

which implies our theorem.

First we define a new experiment $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$. Intuitively, this new experiment is a "mixture" of the hybrid model and the ideal process, in that an uncorrupted party acting as a prover is handled as in the ideal process (i.e., $\mathcal{S}$ will use the trapdoor to simulate a proof), but an uncorrupted party acting as a verifier is handled as in the hybrid model (i.e., no extraction takes place). More precisely, the new experiment runs simulated copies of $\mathcal{Z}$ and $\mathcal{A}$. Messages received from $\mathcal{Z}$ are forwarded to the simulated $\mathcal{A}$, and messages sent by the simulated $\mathcal{A}$ to its environment are forwarded to $\mathcal{Z}$. $\mathcal{E}_{\Pi,1}(1^k)$ is run to produce (σ, τ), then $(pk^*, sk^*) \overset{R}{\leftarrow} \mathsf{TCgen}(1^k)$ are generated. Just as in the case of $\mathrm{IDEAL}_{\hat{\mathcal{F}}_{\mathrm{ZK}}^R,\mathcal{S},\mathcal{Z}}$, (pk^*, σ) is used as the common reference string for $\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}$, and sk^* and τ are stored. If an uncorrupted party P_i receives input ($\mathsf{zk\text{-}prover}$, sid, $ssid$, P_i, P_j, x, w) from $\mathcal{Z}$ with $(x, w) \in R$, it sets the prover data field of its messages in the same way as $\mathcal{S}$ above. Corruptions are handled in the same way as $\mathcal{S}$ above. An uncorrupted party P_j responds to a prover as in the actual verifier protocol in $\hat{\Pi}$. The output of each experiment (hybrid model, ideal process, and $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$) is the output of $\mathcal{Z}$.

Let $\mathrm{MIX}_{\mathcal{A},\mathcal{Z}}$ denote the distribution ensemble of random variables describing the outputs of $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$. First, we can show that $\mathrm{HYB}_{\hat{\Pi},\mathcal{A},\mathcal{Z}}^{\mathcal{F}_{\mathrm{CRS}}^{\mathcal{D}}} \overset{c}{\approx} \mathrm{MIX}_{\mathcal{A},\mathcal{Z}}$. In fact, it comes from the fact that the SSTC scheme is perfectly hiding and a straightforward hybrid reduction to the simulator $\mathcal{S}_\Pi$ of the Ω-protocol Π.[25]

Now we must show that $\mathrm{MIX}_{\mathcal{A},\mathcal{Z}} \overset{c}{\approx} \mathrm{IDEAL}_{\hat{\mathcal{F}}_{\mathrm{ZK}}^R,\mathcal{S},\mathcal{Z}}$, which will finish the proof to our theorem. This will follow similar to the proof of Theorem 5.1, but also using the simulation-sound binding property of the trapdoor commitment scheme.

[25] Note that if a corruption occurs between the first and second messages to the wrapper machine for the simulation, it will be just as if the simulation never received the second message.

Let $p = \Pr[\text{IDEAL}_{\hat{\mathcal{F}}^R_{\text{ZK}},\mathcal{S},\mathcal{Z}}(k)]$ and $p' = \Pr[\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)]$. Similar to the proof of Theorem 5.1, the only difference between $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ and $\mathcal{S}$ comes from when the extractor in $\mathcal{S}$ outputs an incorrect witness for a session started by $\mathcal{A}$, and thus $\mathcal{Z}$ receives an output message (indicating a correct proof) in $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ but not when interacting with $\mathcal{S}$. (Note that the transcripts of corrupted prover/uncorrupted verifier sessions will never be the same as transcripts of uncorrupted prover/corrupted verifier sessions because of the auxiliary data aux.) Let $\vec{b}$ be the vector corresponding to simulated verifier or extractor sessions, with $b = 1$ corresponding to whether $\mathcal{Z}$ receives an output message. Let ρ be the statistical difference between the distribution of vectors $\vec{b}$ resulting from $\mathsf{Mix}_{\mathcal{A},\mathcal{Z}}(k)$ and vectors $\vec{b}$ resulting from $\mathcal{S}$. (Note that $\rho \geq |p - p''|$.) Let u be an upper bound on the number of verifier sessions. Then the average probability of a difference in a given bit position is at least ρ/u.

To complete the proof, we simply need to show that ρ is negligible. Let C be the number of times $\mathcal{Z}$ sends $\mathsf{zk\text{-}prover}$ messages to the parties. Now we construct an adversary $\mathcal{B}$ that breaks the SSTC scheme TC with probability $\frac{1}{2}((\rho/u)^2 - 2^{-k})$ and with at most $2C$ calls to the commitment-revealing oracle. Therefore, it will follow that ρ is negligible.

We describe the adversary $\mathcal{B}$. Let $\mathcal{B}$ take a public key pk of TC along with a TC simulator. First $\mathcal{B}$ chooses a random $\ell \in \{1, \ldots, u\}$, and then it runs as $\mathcal{S}$, except for (1) changing the common reference string from $(*, \sigma)$ to (pk, σ), and (2) using the TC simulator to fake commitments. Also, before sending a challenge (as the second message) in session ℓ, $\mathcal{B}$ forks the experiment and continues independently in each sub-experiment (thus giving random independent challenges to $\mathcal{A}$). Then $\mathcal{B}$ examines the output (x, tr_1, b_1) and (x, tr_2, b_2) in each sub-experiment. If $b_1 = b_2 = 1$ and $x \notin L_R$ (call this a *successful* sub-experiment), and also the challenges in each sub-experiment are distinct, then we know that $\mathcal{A}$ has decommitted differently in two sub-experiments. This is because of the property of the Ω-protocol; if $\mathcal{A}$ had decommitted in the same way, then there exist two accepting conversations with the same first-message, and then a witness should be extracted, indicating that $x \in L_R$. However, now $\mathcal{B}$ has obtained two different decommitments, successfully breaking TC. By Lemma A.1, a successful sub-experiment occurs with probability at least $(\rho/u)^2 - 2^{-k}$, and thus either $\mathcal{B}$ will break the SSTC scheme TC with probability $\frac{1}{2}((\rho/u)^2 - 2^{-k})$, as claimed above. □

6. Efficient Instantiations

Here we briefly describe some efficient instantiations of our constructions. First, we discuss two efficient signature schemes (namely, the Cramer–Shoup signature scheme and the DSA signature scheme) and two associated efficient Σ-protocols that can be plugged into our constructions of SSZK, NMZK, and UCZK protocols. Second, we construct an efficient SSTC scheme based on DSA that can be used in our construction of a UCZK protocol. Third, we give an example of an efficient Ω-protocol for the discrete logarithm relation, thus implying efficient NMZK and UCZK protocols for discrete logarithms. Finally, we describe a generalized definition of Ω-protocols, which can replace Ω-protocols in an appropriately generalized definition of NMZK protocols.[26]

[26] We note that this generalization is not applicable to UCZK protocols.

Then we present a very efficient[27] generalized Ω-protocol for proving knowledge of the plaintext of an ElGamal ciphertext, thus implying an efficient NMZK protocol for ElGamal plaintext knowledge.

6.1. *Signature Schemes*

First we note that for our constructions we can use a more general version of the Σ-protocol for proving knowledge of signatures, as follows. Consider the binary relation $R_{vk} = \{(m, s)\}$ for a signature scheme SIG. We say a polynomial-time computable function f is a *partial knowledge function* of SIG, if there exists a probabilistic polynomial-time machine M such that for every m and vk, $\{s_1\colon\ s_1 \leftarrow M(m, vk)\}$ and $\{s_1\colon\ s \leftarrow \mathsf{sig\_sign}(vk, m);\ s_1 \leftarrow f(m, vk, s)\}$ have the same distribution. Intuitively, a partial knowledge function carries part of the information about the signature, yet can be efficiently sampled without even knowing one.[28] For a signature scheme SIG with partial knowledge function f, the relation $R'_{vk} = \{((m, s_1), s)\colon\ (m, s) \in R_{vk} \wedge s_1 = f(m, vk, s)\}$ can replace R_{vk} in the constructions for $\mathsf{SS}^R_{[vk]}$, $\mathsf{NM}^R_{[vk,\sigma']}(x)$, and $\mathsf{UC}^R_{[pk^*,vk,\sigma']}(x)$, with $\mathcal{P}$ sending a randomly sampled s_1 (partial knowledge) before running the Σ-protocol $\Sigma^R(x) \vee \Sigma^{R'_{vk}}(vk', s_1)$. We say R'_{vk} is a *partial signature relation* for SIG.

Here we show that the Cramer–Shoup signature scheme [17] and the DSA signature scheme [41] both admit efficient Σ-protocols for proving knowledge of signatures using this more general definition, and thus can be plugged into our constructions.

The Cramer–Shoup signature scheme. Cramer and Shoup [17] presented an efficient signature scheme that is existentially unforgeable against adaptive chosen-message attacks under the Strong RSA Assumption, formally defined in Appendix 6.4. In addition to the main security parameter k, they use a secondary security parameter k' for public key modulus size:[29] The value k' is dependent on k and is set so that known attacks on public key systems with modulus size k' are at least as hard as known attacks on hash functions and other brute-force attacks on systems with main security parameter k. Here we describe their scheme, which we denote $\mathsf{SIG}_{\mathsf{CS}} = (\mathsf{sig\_gen}_{\mathsf{CS}}, \mathsf{sig\_sign}_{\mathsf{CS}}, \mathsf{sig\_verify}_{\mathsf{CS}})$:[30]

- $\mathsf{sig\_gen}_{\mathsf{CS}}(1^k)$:
  $p, q \stackrel{R}{\leftarrow} \textsc{SafePrime}(1^{k'/2});\ N \leftarrow pq;\ x, h \stackrel{R}{\leftarrow} \mathsf{QR}_N;\ e' \stackrel{R}{\leftarrow} \textsc{Prime}(1^{k+1});$
  $H \stackrel{R}{\leftarrow} \textsc{Hash}(1^k);\ sk \leftarrow \langle p, q\rangle;\ vk \leftarrow \langle N, h, x, e', H\rangle;$
  return (sk, vk).
- $\mathsf{sig\_sign}_{\mathsf{CS}}(sk, m)$:
  $y' \stackrel{R}{\leftarrow} \mathsf{QR}_N;\ x' \leftarrow (y')^{e'} \cdot h^{-H(m)} \bmod N;\ e \stackrel{R}{\leftarrow} \textsc{Prime}(1^{k+1})\backslash\{e'\};$

[27] In particular, this protocol is more efficient than the best (strict) Ω-protocol that we have found.

[28] We comment that there alway exist *trivial* partial knowledge functions, e.g., the constant function. Naturally we are only interested in ones that allows efficient ZK proofs.

[29] For today's technology, reasonable values may be $k = 256$ and $k' = 1024$.

[30] Some technical notations: a prime number p is a *safe prime* if $(p - 1)/2$ is also a prime number. $\textsc{SafePrime}(1^n)$ is the set of all n-bit safe prime numbers; $\textsc{Prime}(1^n)$ is the set of all n-bit prime numbers; QR_N is the set of all quadratic residues in $\mathbb{Z}^*_N$, and $\textsc{Hash}(1^n)$ is a set of efficient hash functions that maps strings of arbitrary length to an n-bit string.

$y \leftarrow (xh^{-H(x')})^{e^{-1} \bmod \varphi(N)} \bmod N$;
return $\langle e, y, y' \rangle$;

- $\mathsf{sig\_verify}_{\mathsf{CS}}(vk, m, \langle e, y, y' \rangle)$:
  if e is not an odd $k+1$ bit number, or $e = e'$, return 0;
  $x' \leftarrow (y')^{e'} \cdot h^{-H(m)} \bmod N$;
  if $x \equiv y^e h^{H(x')} \bmod N$ return 1, else return 0.

As a technical note, instead of an expected polynomial-time algorithm for prime generation, we assume a probabilistic *strict* polynomial-time algorithm that has a negligible probability of failing. This has no effect on the following security result.

Theorem 6.1 [17]. *The Cramer–Shoup signature scheme is secure against adaptive chosen-message attack, under the Strong RSA assumption and the assumption that H is collision-resistant.*

Note that from a public key vk, a message m, and a signature $\langle e, y, y' \rangle$ on m, one can extract the pair (e, y'). Also note that for a randomly generated signature, this pair (e, y') is random, i.e., e is a random k-bit prime not equal to e', y' is a random element of QR_N,[31] and they are independent. Therefore, function $f(m, vk, \langle e, y, y' \rangle) = (e, y')$ is a partial knowledge function for Cramer–Shoup. Furthermore, given vk, m, and (e, y'), one can compute $x' \leftarrow (y')^{e'} \cdot h^{-H(m)} \bmod N$, and then y is simply a root of a known element, i.e., y is the eth root of $x \cdot h^{H(x')} \bmod N$. Guillou and Quisquater [36] presented a Σ-protocol for proving knowledge of roots that has the special soundness property. Their protocol can be directly adopted here for proving the partial signature relation R'_{vk}.

DSA. The Digital Signature Algorithm [41] was proposed by NIST in April 1991, and in May 1994 was adopted as a standard digital signature scheme in the U.S. [28]. It is a variant of the ElGamal signature scheme [24], and is defined as follows, with two security parameters k and k' as in the Cramer–Shoup signature scheme:[32]

- $\mathsf{sig\_gen}_{\mathsf{DSA}}(1^k)$:
  $q \leftarrow \textsc{Prime}(1^k)$; $p \xleftarrow{R} \textsc{Prime}(1^{k'})$, where $q | (p-1)$; $g \xleftarrow{R} \mathbb{Z}_p^*$, where $\mathsf{order}(g) = q$;
  $x \xleftarrow{R} \mathbb{Z}_q$; $y \leftarrow g^x \bmod p$; $sk \leftarrow \langle g, p, q, x \rangle$; $vk \leftarrow \langle g, p, q, y \rangle$;
  return (sk, vk).
- $\mathsf{sig\_sign}_{\mathsf{DSA}}(sk, m)$:
  $v \xleftarrow{R} \mathbb{Z}_q$; $r \leftarrow g^v \bmod p$; $s \leftarrow v^{-1}(H(m) + xr) \bmod q$;
  return $\langle r \bmod q, s \rangle$.
- $\mathsf{sig\_verify}_{\mathsf{DSA}}(vk, m, \langle r', s \rangle)$:
  If $0 < r' < q$, $0 < s < q$, and $r' \equiv ((g^{H(m)} y^{r'})^{s^{-1} \bmod q} \bmod p) \bmod q$, return 1, else return 0.

[31] We assume that e' is not a factor of $\varphi(N)$, which is false with only negligible probability.

[32] In the DSA standard, k, k', and H are fixed in the following way: $k = 160$, k' is set to a multiple of 64 between 512 and 1024, inclusive, and hash function H is defined as SHA-1 [27]. However, we use these parameters as if they could be varied according to the security level desired.

The security of DSA intuitively rests on the hardness of computing discrete logarithms, but there is no known security reduction that proves this. However, it is often simply assumed that DSA is existentially unforgeable against adaptive chosen-message attack.

Note that from a public key vk, a message m, and a signature $\langle r', s\rangle$, one can efficiently compute a value $r \leftarrow g^{H(m)s^{-1}} y^{r's^{-1}} \bmod p$. Also note that for a randomly generated signature, the value r is a random element generated by g. Therefore, $f(m, vk, \langle r', s\rangle) = r$ is a partial knowledge function for DSA. Furthermore, given vk, m, and r, s is simply a discrete log base r of the known element $g^{H(m)} y^r \bmod p$. Schnorr [55] presents a Σ-protocol for proving knowledge of a discrete log, which satisfies the special soundness property. This protocol can be used to prove the partial signature relation R'_{vk}.

6.2. *SSTC Scheme*

Here we present an efficient SSTC scheme TC based on DSA. First, though, we describe a slightly simpler scheme TC' for *weak* SSTCs, when *id* is always the empty string (and thus, in essence, no double reveal queries to the trapdoor commitment simulator are allowed). We can implement this simpler scheme over elements from a group $(G, +)$ by using a technique similar to that in [19] that involves two trapdoor commitment schemes TC_0 and TC_1 that commit to elements in G. The trapdoor in TC' is the trapdoor of one of TC_0 or TC_1 along with a bit indicating which. To commit to a message m, generate random $m_0 \in G$, set $m_1 \leftarrow m - m_0$, and commit to m_0 and m_1 using TC_0 and TC_1, respectively, i.e., generating commitment (c_0, c_1). To open a commitment (c_0, c_1), open each commitment, say to (m_0, m_1). Then $m = m_0 + m_1$ is the decommitted value. To open a commitment of (c_0, c_1), say of (m_0, m_1), to an arbitrary value m' using trapdoor (b, sk_b), i.e., trapdoor sk_b of TC_b, open commitment c_{1-b} normally, and use sk_b to open commitment c_b to $m' - m_{1-b}$. (A proof that this satisfies the weak simulation-sound binding property follows closely from [19].) This scheme does not satisfy the full notion of simulation-sound binding, since after revealing a commitment in two different ways (even one with an arbitrary *id*), the adversary can determine which trapdoor is used, and this would cause the proof from [19] to fail.[33]

Our scheme TC that satisfies simulation-sound binding uses the same technique as above of being built over two commitment schemes TC_0 and TC_1, but each of those will be built over DSA as follows. Given a DSA public key (g, p, q, y), a commitment to a message m using *id* is generated as follows. First compute $\alpha \stackrel{R}{\leftarrow} \mathbb{Z}_q$, $g' \leftarrow g^\alpha \bmod p$, and $h = g^{H(id)} y^{g'} \bmod p$. (Note that s is the discrete log of h over g' if and only if $(g' \bmod q, s)$ is a DSA signature for *id*.) Then use a Pedersen commitment [49] over bases (g', h) to commit to m, i.e., choose $\beta \stackrel{R}{\leftarrow} \mathbb{Z}_q$ and compute commitment (g', c) where $c \leftarrow (g')^m h^\beta$. To open this commitment, output (m, β).

To show the simulation-sound binding property, we show that if an adversary can break this property, we can break DSA as follows. (We assume that DSA is existentially unforgeable against a adaptive chosen-message attack.) Given a DSA key vk_0 and

[33] One could use this scheme with weak simulation soundness in our UCZK adaptive protocol, but it would require the common reference string to contain one trapdoor commitment public key for each party.

signature oracle, we generate another DSA key pair (vk_1, sk_1), choose a bit b, and say (vk_b, vk_{1-b}) is the public key for our commitment scheme.

Now say we know which id the adversary is going to use in its commitment with double opening. To commit to a value v using id, we compute an actual signature for id using sk_1, and then commit to some value using that signature. Then we use the knowledge of the signature to decommit to an arbitrary value m. (Note that we can do this, since knowledge of the DSA signature implies knowledge of the discrete log used in the Pedersen commitment, which further allows equivocation of the commitment.) To commit to a value v using $id' \neq id$, we choose a bit b' to decide which scheme to compute a signature (and thus which scheme will be used to equivocate on commitments, as discussed above). If $b' = 0$, we compute a signature using the DSA signature oracle on id', and if $b' = 1$ we compute a signature using sk_1.

Now the adversary's view is independent of b, and thus if the adversary gives a double opening with id, then with probability at least $\frac{1}{2}$, there will be different openings (m_0, β_0) and (m'_0, β'_0) of (g'_0, c_0), so $(g'_0 \bmod q, (\beta'_0 - \beta_0)/(m_0 - m'_0) \bmod q)$ is a signature (with respect to the public key vk_0) on id, breaking DSA. Note that if we do not know which id will be used by the adversary, we would have to guess this, reducing the probability of breaking DSA by a polynomial factor (at most proportional to the running time of the adversary).

6.3. *An Efficient Ω-Protocol*

We describe an efficient Ω-protocol for proving knowledge of a discrete logarithm. This protocol is based on the Decisional Composite Residuosity assumption and the Strong RSA assumption, formally defined in Appendix 6.4.

Let (g, p, q) be public parameters, where q and p are primes with $q|(p-1)$, and $g \in \mathbb{Z}_p^*$ with $\mathsf{order}(g) = q$. Let R be the discrete logarithm relation: $R = \{(y, x)\colon\ y \equiv g^x \bmod p\}$. Our Ω-protocol for R is constructed as follows: The common reference string consists of two parts: (1) a Paillier public key $pk = \langle N, h\rangle$ where N is an RSA modulus and $h \in \mathbb{Z}_{N^2}^*$ with $N|\mathsf{order}(h)$, and (2) a triple $\langle \tilde{N}, h_1, h_2\rangle$ where $\tilde{N} = (2\tilde{P}+1)(2\tilde{Q}+1)$ is an RSA modulus where $\tilde{P}$ and $\tilde{Q}$ are prime, and h_1 and h_2 are generators in $\mathbb{Z}^*_{\tilde{P}\tilde{Q}}$. The prover and the verifier share a common input y, while the prover also knows x, such that $g^x = y$. In the first message, the prover sends an encryption of x using the Paillier encryption key pk. Then a Σ-protocol is used to prove that the plaintext in the Paillier encryption is indeed the discrete log of y. A technical difficulty is that the discrete logarithm and the Paillier encryption work in different moduli. To overcome this we use the known technique of adding a commitment to x using two generators (h_1, h_2) over a third modulus $\tilde{N}$ of unknown factorization [7]–[9], [29], [43]. The detailed construction is presented in Appendix 6.4.

6.4. *An Efficient Generalized Ω-Protocol*

For an NP relation $R = \{(x, w)\}$ and a polynomial-time computable function f, let $R_f = \{(x, f(w)) : (x, w) \in R\}$. (Note that R_f may not itself be an NP relation.) Then we define an f-extracting Ω-protocol for R as an Ω-protocol for R except that

the extractor $\mathcal{E}_2$ outputs $f(w)$, instead of w. Similarly, we can define an f-extracting NMZK protocol in which the extractor $\mathcal{E}_2$ outputs $f(w)$, instead of w, and the extraction condition is changed appropriately.[34] It is easy to see that if we replace the Ω-protocol in our construction of NMZK protocols with an f-extracting Ω-protocol, our construction yields an f-extracting NMZK protocol. Note that the prover in both Ω-protocols and NMZK protocols still receives the "full" witness w. Also note that if f is the identity function, we have the normal definitions of an Ω-protocol and an NMZK protocol.

One application of these generalized definitions is in proving plaintext knowledge. See [39] for some applications of proof of plaintext knowledge. Consider a semantically secure encryption scheme. This scheme naturally induces a relation $R = \{(e, (x, r))\}$, where e is the encryption of plaintext x using random bits r. Then consider a function f defined as $x \leftarrow f(x, r)$. It is easy to see that an f-extracting Ω-protocol for R is essentially a proof of plaintext knowledge, so we call this function f a *plaintext knowledge function*.

We now present a very efficient f-extracting Ω-protocol for ElGamal encryption, where f is a plaintext knowledge function. Let (g, p, q) be public parameters, where q and p are primes with $q|(p-1)$, and $g \in \mathbb{Z}_p^*$ with $\mathsf{order}(g) = q$. Then the ElGamal encryption scheme can be formally defined as follows, with the message space being the subgroup generated by g:

- $\mathsf{enc\_gen}_{\mathsf{EG}}(g, p, q)$:
  $x \stackrel{R}{\leftarrow} \mathbb{Z}_q$; $y \leftarrow g^x \bmod p$; $sk \leftarrow x$; $pk \leftarrow y$;
  return (sk, pk).
- $\mathsf{encrypt}_{\mathsf{EG}}(vk, m)$:
  $r \stackrel{R}{\leftarrow} \mathbb{Z}_q$; $a \leftarrow g^r \bmod p$; $b \leftarrow my^r \bmod p$;
  return (a, b).
- $\mathsf{decrypt}_{\mathsf{EG}}(sk, (a, b))$:
  return b/a^x

The relation for the ElGamal system is $R = \{((a, b), (m, r))$: $(a \equiv g^r \bmod p) \wedge (b \equiv my^r \bmod p)\}$, and f is defined such that $m \leftarrow f(m, r)$. The f-restricted Ω-protocol is constructed as follows. The common reference string is a new public key y' for the ElGamal system, which is generated by running $(x', y') \leftarrow \mathsf{enc\_gen}_{\mathsf{EG}}(g, p, q)$ using fresh random bits. The corresponding decryption key x' is discarded. The prover takes $(a, b) = (g^r, my^r)$, which is an encryption of a message m (using random bits r), and then constructs a new encryption using the encryption key in the common reference string $(a', b') \leftarrow (g^{r'}, m(y')^{r'})$, where $r' \stackrel{R}{\leftarrow} \mathbb{Z}_q$. The prover then sends (a', b') to the verifier, and performs a Σ-protocol proving that the two ElGamal encryptions have the same plaintext. The Σ-protocol proceeds as follows. The prover picks $w, w' \leftarrow \mathbb{Z}_q$, computes $d \leftarrow g^w$, $d' \leftarrow g^{w'}$, and $e \leftarrow y^w/(y')^{w'}$, and outputs (d, d', e) as the first message. On challenge c, the prover computes $s \leftarrow rc + w \bmod q$ and $s' \leftarrow r'c + w' \bmod q$, and

[34] Note that the resulting NMZK protocols cannot necessarily be used to construct UCZK protocols (even with static corruptions), since UCZK protocols are, by definition, proofs of knowledge.

outputs (s, s') as the third message. Finally the verifier verifies that $g^s = a^c d$, $g^{s'} = (a')^c d'$, and $y^s/(y')^{s'} = (b/b')^c e$.

SHVZK is satisfied since given input $(a, b) \in L_R$ and a challenge c, a simulator can generate an encryption (a', b') of an arbitrary value, and then use the perfect SHVZK property of the Σ-protocol to generate an accepting conversation. By the semantic security of ElGamal, the output of the simulator is computationally indistinguishable from that of an actual prover. Now we show the f-extraction property is satisfied. Let $\mathcal{E}_1(1^k)$ generate a fresh ElGamal key pair $(sk', vk') \leftarrow \mathsf{enc\_gen}_{\mathsf{EG}}(g, p, q)$, putting vk' in the common reference string, and passing the decryption key sk' to $\mathcal{E}_2$, which then interacts with the prover and obtains an accepting transcript tr. Finally $\mathcal{E}_2$ outputs $m' \leftarrow \mathsf{decrypt}_{\mathsf{EG}}(sk', (a', b'))$ where (a', b') is the encryption in the transcript tr. By the weak soundness property of the Σ-protocol, the probability that m' is not the plaintext in the encryption (a, b) is at most 2^{-k} (assuming k-bit challenges).

Acknowledgements

We thank the anonymous referees for their many useful comments and suggestions.

Appendix A. The Exclusive Collision Lemma

We prove the lemma used in the proof of Theorem 3.1.

Lemma A.1 (The Exclusive Collision Lemma). *Let A be a random variable and let B_a be a random variable whose distribution is parameterized by a value a in the support of A. For every a in the support of A, and for every b_1 and b_2 in the support of B_a, let $\mathsf{Coll}_a(b_1, b_2)$ be a predicate defining a* collision. *Let q be the maximum (over all a in the support of A) probability of a collision of two independent random variables B_a^1 and B_a^2, i.e., $q = \max_a\{\mathsf{Prob}[\mathsf{Coll}_a(B_a^1, B_a^2)]\}$. Let $\varphi(a, b)$ be a predicate, and let $p = \mathsf{Prob}[\varphi(A, B_A)]$. Let $\varphi'(a, b_1, b_2) = \varphi(a, b_1) \wedge \varphi(a, b_2) \wedge (\neg\mathsf{Coll}_a(b_1, b_2))$. Then we have $\mathsf{Prob}[\varphi'(A, B_A^1, B_A^2)] \geq p^2 - q$, where B_A^1 and B_A^2 are independent conditioned on A.*

Proof. We define a new predicate $\varphi''(a, b_1, b_2) = \varphi(a, b_1) \wedge \varphi(a, b_2)$, which is essentially predicate φ' without the requirement that $\neg\mathsf{Coll}_a(b_1, b_2)$. For every a in the support of A, let $p_a = \mathsf{Prob}[\varphi(a, B_a)]$. Let p_A be the function of random variable A taking value p_a when $A = a$. Then we have $p = \mathsf{Prob}[\varphi(A, B_A)] = E[p_A]$ and $\mathsf{Prob}[\varphi''(A, B_A^1, B_A^2)] = E[(p_A)^2] \geq (E[p_A])^2 = p^2$. (The inequality holds because $E[(p_A)^2] - (E[p_A])^2$ is the variance of p_A, and the variance is always positive.)

Finally we have

$$\mathsf{Prob}[\varphi'(A, B_A^1, B_A^2)] \geq \mathsf{Prob}[\varphi''(A, B_A^1, B_A^2)] - \mathsf{Prob}[\mathsf{Coll}_A(B_A^1, B_A^2)] \geq p^2 - q. \quad \square$$

We remark that, using a tighter analysis, the lower bound on $\mathsf{Prob}[\varphi'(A, B_A^1, B_A^2)]$ in Lemma A.1 can be improved to $p^2 - pq$.

Appendix B. Number-Theoretic Assumptions

We review some of the number-theoretic assumptions used in this paper.

The Strong RSA assumption. The Strong RSA assumption is a generalization of the standard RSA assumption which (informally) states that given an RSA modulus N and an exponent e, it is computationally infeasible to find the eth root of a random x. Informally, the Strong RSA assumption states that it is infeasible to find an *arbitrary* non-trivial root of a random x.

More formally, we say that p is a *safe prime* if both p and $(p-1)/2$ are prime. Then let $\mathsf{RSA\text{-}Gen}(1^k)$ be a probabilistic polynomial-time algorithm that generates two random $k/2$-bit safe primes p and q, and outputs $N \leftarrow pq$.

Assumption B.1 (Strong-RSA). *For any non-uniform probabilistic polynomial-size circuit $\mathcal{A}$, the following probability is negligible in k*:

$$\Pr[N \leftarrow \mathsf{RSA\text{-}Gen}(1^k); x \leftarrow \mathbb{Z}_N^*; (y, e) \leftarrow \mathcal{A}(1^k, x, N) : y^e \equiv x \bmod N \wedge e \geq 2].$$

The Strong RSA assumption was introduced by Barić and Pfitzmann [4], and has been used in several applications (see [29], [30], and [17]). It is a stronger assumption than the "standard" RSA assumption, yet no method is known for breaking it other than factoring N.

The Paillier cryptosystem and the decision composite residuosity assumption. The Paillier encryption scheme [48] is defined as follows, where $\lambda(N)$ is the Carmichael function of N, and L is a function that takes input elements from the set $\{u < N^2 | u \equiv 1 \bmod N\}$ and returns $L(u) = (u-1)/N$. This definition differs from that in [48] only in that we define the message space for public key $pk = \langle N, g\rangle$ as $[-(N-1)/2, (N-1)/2]$ (versus $\mathbb{Z}_N$ in [48]), and we restrict h to be $1+N$. The security of this cryptosystem relies on the *decision composite residuosity assumption*, DCRA.

For key generation, choose random $k/2$-bit primes p, q, set $N = pq$, and set $h \leftarrow 1 + N$. The public key is $\langle N, h\rangle$ and the private key is $\langle N, h, \lambda(N)\rangle$. To encrypt a message m with public key $\langle N, h\rangle$, select a random $\alpha \in \mathbb{Z}_N^*$ and compute $c \leftarrow g^m \alpha^N \bmod N^2$. To decrypt a ciphertext c with secret key $\langle N, h, \lambda(N)\rangle$, compute $m = (L(c^{\lambda(N)} \bmod N^2)/L(g^{\lambda(N)}) \bmod N^2) \bmod N$, and the decryption is m if $m \leq (N-1)/2$, and otherwise the decryption is $m - N$. Paillier [48] shows that both $c^{\lambda(N)} \bmod N^2$ and $g^{\lambda(N)} \bmod N^2$ are elements of the form $(1+N)^d \equiv_{N^2} 1 + dN$, and thus the L function can be easily computed for decryption.

Appendix C. An Efficient Ω-Protocol for Proving Knowledge of Discrete Log

The detailed construction of the Ω-protocol for proving knowledge of the discrete logarithm is given in Fig. 6.

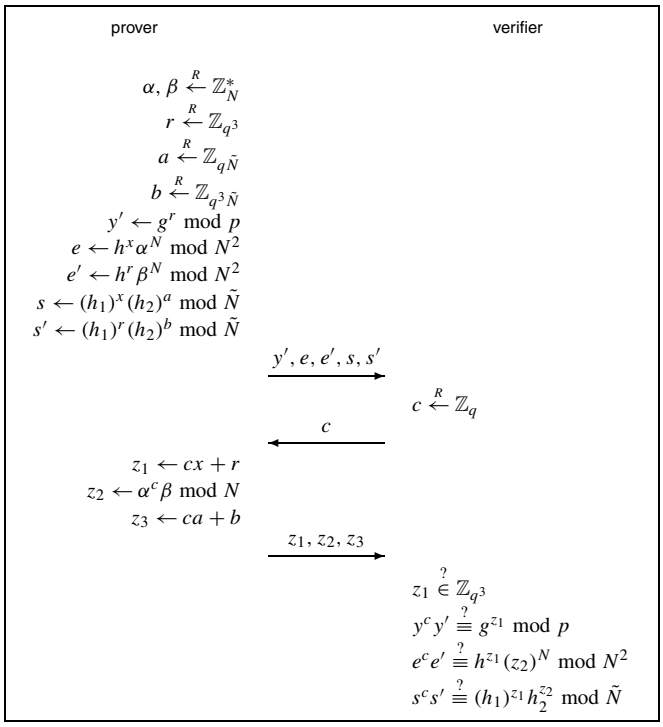

Fig. 6. Ω-protocol for the discrete log relation $\{(y, x) : y \equiv g^x \bmod p\}$. The common reference string is a Paillier public key and a Strong RSA modulus along with two generators $((N, h), (\tilde{N}, h_1, h_2))$.

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