Predictable Arguments of Knowledge

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Abstract. We initiate a formal investigation on the power of *predictability* for argument of knowledge systems for NP. Specifically, we consider private-coin argument systems where the answers of the prover can be predicted, given the private randomness of the verifier.
We show that predictable arguments of knowledge (PAoK) can be made extremely laconic, with the prover sending a single bit, and assumed to have only one round (two messages) without loss of generality. We then explore constructs of PAoK. For specific relations we obtain PAoK from Extractable Hash Proof systems (Wee, Crypto '10); we also show that PAoK are equivalent to Extractable Witness Encryption. Unfortunately, the latter poses serious doubts on the existence of PAoK for all NP.
Finally, we apply PAoK in the context of leakage-tolerant PKE protocols. At PKC '13 Nielsen *et al.* have shown that any leakage-tolerant PKE protocol requires long keys already when it tolerates super-logarithmic leakage. We strengthen their result proving a more fine-grained lower bound for any constant numbers bits of leakage. Moreover, we consider an adaptive-secure definition for PAoK. In this stronger model the challenge generation algorithm is independent of the instance proved, the extraction is guaranteed for adversarial chosen instances.

1 Introduction

Consider the classical proof system for Graphs Non-Isomorphism where, on common input a tuple of graphs (G_0, G_1), the verifier choses a uniformly random bit b, and sends a uniformly random permutation of the graph G_b to the prover. If the two graphs are not isomorphic the prover can reply correctly sending back the value b.

A peculiar property of the above proof system is that the verifier knows in advance the answer of the prover, i.e., the answer given by the prover is *predictable*. Another interesting property is that it uses only one round of communication and that the prover sends a single bit. Following the work of Goldreich *et al.* in [13] we call a proof system with this property *extremely laconic*. We study the notion of predictability in interactive proof systems for NP, specifically, we focus on the more cryptographic setting where the prover's strategy is efficiently computable and, moreover, we aim for the notion of knowledge soundness, where any convincing polynomial-time prover must "know" the witness relative to the instance. We formalise this notion of Predictable Arguments of Knowledge (PAoK) and explore their properties and applications.

1.1 Our Contributions

Characterizing PAoK. In Section 3, we show that PAoK can always be made extremely laconic both in term of round complexity and in term of the number of bits send by the prover (i.e. message

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complexity). For the former we show that we can collapse any multi-round PAoK into a one-round PAoK with higher message complexity. For the latter we show how to reduce the message length to a single bit using the Goldreich-Levin Hard-Core Theorem (see Goldreich and Levin [12]). Interestingly, we can wrap up the two results together showing that any PAoK, no matter of the round or message complexity can be made extremely laconic.

We sketch the main idea used in the first step. Consider a PAoK with round complexity ρ and a random sequence of challenges $(c_1, \ldots, c_\rho)$ where c_i is for round i. Without loss of generality the verifier can define all the challenge before any answers of the prover. Let $(a_1, \ldots, a_\rho)$ be the predicted answers. Consider for simplicity a cheating prover who can answer all of these challenges with probability $1/2$. Then for many possible challenges $(c_1, \ldots, c_\rho)$ there must exists a round i such that the prover cannot answer c_i with a_i with probability 1 even if it is given $(c_1, \ldots, c_{i-1})$ and $(a_1, \ldots, a_{i-1})$. Hence the verifier could sample $(c_1, \ldots, c_\rho)$ and $(a_1, \ldots, a_\rho)$, choose a uniformly random i, send $(c_1, \ldots, c_{i-1})$ and $(a_1, \ldots, a_{i-1})$ and expect to get back a_i. This would give non-zero probability of catching the cheating prover on many instances.

A main problem with this idea, however, is that if the round complexity of the protocol grows with the security parameter (which is the interesting case), the soundness will be vanishing. Therefore, we would like to amplify the soundness, however we cannot use sequential composition, as we are trying to construct a one-round protocol. Furthermore, in general, interactive arguments do not amplify soundness when compose in parallel (see Bellare *et al.* [1] and Pietrzak and Wikström [19]). It turns out that a more involved construction, inspired by parallel composition of the above simple protocol will actually allow to boost the soundness negligibly close to 1. We refer the reader to Section 3.2 for the details.

Constructions. In Section 4 we explore constructs of PAoK. In Section 4.1 we show how to obtain a PAoK, for some specific relations, using an Extractable Hash-Proof system (see Wee [20]).

Next, we show that PAoK for a relation R is equivalent to Extractable Witness Encryption (Ext-WE) (see Goldwasser *et al.* [14]) for R. The main ideas are the following. From Ext-WE to PAoK we encrypt a random bit a using the encryption scheme and then ask the prover to return a. For the other direction, namely from PAoK to Ext-WE, we first make the PAoK extremely laconic, generate a challenge/answer pair (c, a) for the PAoK, and then encrypt a single bit β as $(c, a \oplus \beta)$. We give the details in Section 4.2 .

Applications. We mainly investigate the notion of PAoK because we find it intriguing in its own right. We do, however, note that PAoK have a number of interesting properties and applications. Note for instance that a PAoK clearly is honest-verifier zero-knowledge, and therefore can be made zero-knowledge using standard reductions. In Appendix A we show an application of PAoK to proving lower bounds on the complexity of leakage-tolerant protocols. In particular, we show that if one uses a public-key encryption (PKE) scheme to implement secure message transmission, and the resulting protocol can tolerate leakage of even a constant number of bits, then the scheme needs to have secret keys which are as long as the total number of bits transmitted using that key.

Since the typical interpretation of PKE is that an unbounded number of messages can be encrypted using a fixed public key, this shows that typical PKE cannot be used to realize leakage-

tolerant secure message transmission even against a constant number of leaked bits. This strengthens an earlier result by Nielsen *et al.* [17] who showed a lower bound for schemes tolerating super-logarithmic leakage.

For the proof we need that there exists a PAoK for a relation associated to the PKE scheme, essentially proving knowledge of the secret key. The result can therefore be interpreted as showing that either a PKE does not give secure message transmission secure against a constant number bits of leakage or there does not exists a PAoK for the applied PKE scheme. Proving that such a PAoK does not exists seems very challenging using current techniques, so our result indicates that we probably cannot base leakage-tolerant message transmission secure against leaking a constant number of bits on PKE with our current proof techniques.

1.2 Related Work

A study of interactive proofs with laconic provers was done also in [11,13]. They do not investigate proofs of *knowledge*. As explained above our notion of PAoK is related to Ext-WE first proposed by Goldwasser *et al.* in [14], where they directly assume that the construction of Garg *et al.* in [9] is extractable. Our technique for weak PAoK from public-coin diO is related to the work of Boyle *et al.* [3].

In [8] Garg *et al.* showed that the notion of Ext-WE is "implausible", namely, assuming a virtual black-box obfuscation (VBB) for a specific function then Ext-WE for a specific NP relation is impossible. The reason relies on the auxiliary information that an adversary might have on the input, if such kind of VBB exists then the auxiliary input can be an obfuscated circuit that allows to decrypt ciphertexts but does not give any information about the witness. As stated by the authors of [8], this can be interpreted as an "implausibility" of the notion of Ext-WE (for all of NP). Notice that the result holds only when the adversary gets an arbitrary auxiliary input.

As mentioned above, in general, arguments do not compose nicely in parallel, however there are some exceptions like 3-messages arguments [1,5], public-coin arguments [18,7] and *simulatable* (a generalization of both 3-messages and public-coin) arguments [16,6]. (multi-round) PAoK are inherently private coins so the mentioned works do not apply directly. More relevant to ours is the work of Haitner on random-terminating arguments [15].

2 Notation

For $a, b \in \mathbb{R}$, we let $[a, b] = \{x \in \mathbb{R} \ : \ a \leq x \leq b\}$; for $a \in \mathbb{N}$ we let $[a] = \{1, 2, \ldots, a\}$. If x is a string, we denote its length by $|x|$; if $\mathcal{X}$ is a set, $|\mathcal{X}|$ represents the number of elements in $\mathcal{X}$. When x is chosen randomly in $\mathcal{X}$, we write $x \leftarrow\!\!\$\ \mathcal{X}$. When A is an algorithm, we write $y \leftarrow \mathsf{A}(x)$ to denote a run of A on input x and output y; if A is randomized, then y is a random variable and $\mathsf{A}(x; r)$ denotes a run of A on input x and randomness r. An algorithm A is *probabilistic polynomial-time* (PPT) if A is randomized and for any input $x, r \in \{0,1\}^*$ the computation of $\mathsf{A}(x; r)$ terminates in at most $\mathsf{poly}(|x|)$ steps.

Throughout the paper we let $\kappa \in \mathbb{N}$ denote the security parameter. We say that a function $\mu : \mathbb{N} \to \mathbb{R}$ is negligible in the security parameter κ if $\mu(\kappa) = \kappa^{-\omega(1)}$. A positive function ν is

noticeable if there exist a positive polynomial $p(\cdot)$ and a number κ_0 such that $\nu(\kappa) \geqslant 1/p(\kappa)$ for all $\kappa \geqslant \kappa_0$. For two ensembles $\mathcal{X} = \{X_\kappa\}_{\kappa\in\mathbb{N}}$ and $\mathcal{Y} = \{Y_\kappa\}_{\kappa\in\mathbb{N}}$, we write $\mathcal{X} \equiv \mathcal{Y}$ if they are identically distributed and $\mathcal{X} \approx \mathcal{Y}$ to denote that they are statistically or computationally close.

Vectors and matrices are typeset in boldface. For a vector $\mathbf{v} = (v_1, \ldots, v_n)$ we sometimes write $\mathbf{v}[i]$ for the i-th element of $\mathbf{v}$ and $\mathbf{v}_{\downarrow i}$ for the vector $(v_1, \ldots, v_i)$.

3 Predictable Arguments of Knowledge

Let $R \subseteq \{0,1\}^* \times \{0,1\}^*$ be an NP-relation, naturally defining a language $L_R := \{x : \exists w \text{ s.t. } (x,w) \in R\}$. We are typically interested in efficiently samplable relations, for which there exists a ppt algorithm SamR taking as input the security parameter (and random coins r) and outputting a pair $(x,w) \in R$. An interactive protocol for R features a prover P (holding a value $x \in L_R$ together with a corresponding witness w) and a verifier V (holding x), where the goal of the prover is to convince the verifier that $x \in L_R$. At the end of the protocol execution the verifier outputs either $\texttt{acc}$ or $\texttt{rej}$; we write $\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle$ for the random variable corresponding to the verifier's verdict.

We start by defining Predictable Arguments of Knowledge (PAoK) in Section 3.1 as one-round interactive protocols in which the verifier generates a challenge (to be sent to the prover) and can at the same time predict the prover's answer to that challenge; we insist on (computational) extractable security, meaning that from any prover convincing a verifier with some probability we can extract a witness with probability related to the prover's success probability. The main result of this section is that PAoK can be assumed without loss of generality to be one-round and extremely laconic (i.e., the prover sends a single bit). In particular, in Section 3.2, we show that any multi-round PAoK can be squeezed into a one-round PAoK; In Section 3.3 we show that, for any $\ell \in \mathbb{N}$, there exists a laconic PAoK if and only if there exists a PAoK where the prover's answer is of length ℓ.

3.1 The Definition

We focus on one-round protocols where the verifier speaks first by sending a challenge message c, to which the prover returns an answer a. Importantly, we are interested in protocols where the verifier can predict the prover's answer at the time when it generates the challenge. Such protocols are fully specified by a tuple of two ppt algorithms $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ as described below:

1. V samples $(c, b) \leftarrow \mathsf{Chall}(1^\kappa, x)$ and sends c to P.
2. P samples $a \leftarrow \mathsf{Resp}(1^\kappa, x, w, c)$ and sends a to V.
3. V outputs $\texttt{acc}$ if and only if $a = b$.

We say that prover P and verifier V, running the protocol above, *execute a PAoK* Π upon input security parameter 1^κ, common input x and prover's private input w; we denote with $\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle_\Pi$ (or simply $\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle$ when Π is clear from the context) the output of such interaction. We say that a prover P *succeeds* on the instance x and auxiliary input w if $\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle = \texttt{acc}$.

Definition 1 (Predictable Arguments of Knowledge). *Let $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ be as specified above, and let R be an NP relation. Consider the properties below.*

Completeness: *There exists a negligible function μ such that for all $(x, w) \in R$, we have that:*

$$\Pr_{\mathsf{P},\mathsf{V}}[\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle = \mathtt{acc}] \geqslant 1 - \mu(\kappa).$$

f-Knowledge soundness with error ϵ: *For all* ppt *provers P^* there exists a non-uniform extractor K and a non-zero polynomial $q(\cdot)$ such that for any $x \in \{0,1\}^*$ and any auxiliary input $z \in \{0,1\}^*$ the following holds. Whenever $p(\kappa) = \Pr[\langle \mathsf{P}^*(1^\kappa, x, z), \mathsf{V}(1^\kappa, x)\rangle = \mathtt{acc}] > \epsilon(\kappa)$, then*

$$\Pr_{\mathsf{K}}\left[\begin{matrix}\exists w \text{ s.t. } f(w) = y \\ (x, w) \in R\end{matrix} : y \leftarrow\!\!\$\ \mathsf{K}(1^\kappa, x, z)\right] \geqslant q(p(\kappa) - \epsilon(\kappa)).$$

Let ℓ be the size of the prover's answer, we call Π a predictable argument of knowledge (PAoK) for R if Π satisfies completeness and f-knowledge soundness for any efficient computable function f, and moreover $\epsilon - 2^{-\ell}$ is negligible. We call it a laconic *PAoK if $\ell = 1$. We call it an f-PAoK if knowledge soundness holds for a specific function f.*

3.2 On Multi-Round PAoK

In this section we consider a natural extension of predictable arguments where there are $\rho > 1$ rounds. In particular, we show that multi-round PAoK can be squeezed into a one-round PAoK (maintaining knowledge soundness).

In a multi-round predictable argument the verifier produces many challenges $\mathbf{c} = (c_1, \ldots, c_\rho)$. W.l.o.g. we can assume that all the challenges are generated together and then forwarded one-by-one to the prover; this is because the answers are known *in advance*. Specifically a ρ-round PAoK is fully specified by a tuple of algorithms $\Pi = (\mathsf{Chall}, \mathsf{Resp})$, as described below:

1. V samples $(\mathbf{c}, \mathbf{b}) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa, x)$, where $\mathbf{c} := (c_1, \ldots, c_\rho)$ and $\mathbf{b} := (b_1, \ldots, b_\rho)$
2. For all $i \in [\rho]$ in increasing sequence:
   - V forwards c_i to P;
   - P computes $(a_1, \ldots, a_i) := \mathsf{Resp}(1^\kappa, x, w, c_1, \ldots, c_i)$ and forwards a_i to V;
   - V checks that $a_i = b_i$, and returns $\mathtt{rej}$ if this is not the case.
3. If all challenges are answered correctly, V returns $\mathtt{acc}$.

Notice that now algorithm Resp takes as input all challenges up-to round i in order to generate the i-th answer.[3]

Definition 2 (ρ-round PAoK). *Let R be an NP relation, $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ be as above, and denote by ℓ the size of each answer produced by the prover. We call Π a ρ-round PAoK for R if Π satisfies completeness and knowledge soundness with error ϵ, and moreover $\epsilon - 2^{-\rho\ell}$ is negligible.*

[3] In the description above we let Resp output also all previous answers $a_1, \ldots, a_{i-1}$; while this is not necessary it can be assumed w.l.o.g. and will simplify the proof of Theorem 1.

Let $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ be a ρ-round PAoK. Consider the following protocol between prover P' and verifier V'—let us call it the *collapsed protocol* for future reference—for a parameter $\varphi \in \mathbb{N}$ to be determined later:

1. For all $i \in [\varphi]$ and $j \in [\rho]$, V' samples $(\mathbf{c}^{i,j}, \mathbf{b}^{i,j}) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa, x)$. For each $j \in [\rho]$ let:
$$\mathbf{C}^j := (\mathbf{c}^{1,j}_{\downarrow j}, \mathbf{c}^{2,j}_{\downarrow j}, \ldots, \mathbf{c}^{\varphi,j}_{\downarrow j})$$
$$\mathbf{B}^j := (\mathbf{b}^{1,j}_{\downarrow j}, \mathbf{b}^{2,j}_{\downarrow j}, \ldots, \mathbf{b}^{\varphi,j}_{\downarrow j}).$$
V' sends the vectors $\mathbf{C} = \big(\mathbf{C}^1, \mathbf{C}^2, \ldots, \mathbf{C}^\rho\big)$.
2. For all $i \in [\varphi]$ and $j \in [\rho]$, P' computes the answer $\mathbf{a}^{i,j} \leftarrow\!\!\$\ \mathsf{Resp}(1^\kappa, x, w, \mathbf{C}^j[i])$. For each $j \in [\rho]$ let:
$$\mathbf{A}^j := (\mathbf{a}^{1,j}, \mathbf{a}^{2,j}, \ldots, \mathbf{a}^{\varphi,j}).$$
P' sends the vector $\mathbf{A} := (\mathbf{A}^1, \mathbf{A}^2, \ldots, \mathbf{A}^\rho)$.
3. V' sets $\mathbf{B} := (\mathbf{B}^1, \mathbf{B}^2, \ldots, \mathbf{B}^\rho)$ and outputs $\mathtt{acc}$ if and only if $\mathbf{A} = \mathbf{B}$.

We write $\Pi' := (\mathsf{Resp}'_\varphi, \mathsf{Chall}'_\varphi)$ for the algorithms describing the generation of the challenge $\mathbf{C}$ and the predicted answer $\mathbf{B}$ by the prover, and the generation of the answer $\mathbf{A}$ by the verifier in the collapsed protocol (respectively).

Theorem 1. *For any polynomial $\rho(\cdot)$ and any function f if $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ is a $\rho(\kappa)$-round f-PAoK with knowledge error $\epsilon(\kappa)$, then the above collapsed protocol $\Pi' = (\mathsf{Chall}'_\varphi, \mathsf{Resp}'_\varphi)$ with parameter $\varphi = \omega(\log \kappa) \cdot \rho$ is an f-PAoK with knowledge error $\epsilon^{\omega(\log \kappa)}$.*

Proof. We start showing knowledge soundness of Π'. Let P' be a prover for the collapsed protocol, and define $p' := \Pr[\langle \mathsf{P}'(1^\kappa, x, z), \mathsf{V}'(1^\kappa, x)\rangle = \mathtt{acc}]$. We construct a prover P^* (with oracle access to P') that convinces the verifier V of the ρ-round PAoK Π with probability greater than $1/3$.

We start by setting up some notation. For any vector $\mathbf{c}$, any challenge $\mathbf{C}$ as defined in the collapsed protocol, and any pair of indexes i, j, let $\mathbf{C}_{|\mathbf{c}^{i,j}=\mathbf{c}}$ be defined as $\mathbf{C}$ but with the vector $\mathbf{c}^{i,j}$ set to $\mathbf{c}$. Let $Z_i(\mathbf{B}, \mathbf{C})$ be the indicator random variable for the event

$$\forall j \in [\rho] :\ \mathbf{a}^{i,j} = \mathbf{b}^{i,j}, \text{ where } \mathbf{A} = \mathsf{P}'(1^\kappa, \mathbf{C}).$$

For any $i \in [\varphi]$ define δ_i to be the following probability:

$$\delta_i := \Pr\left[\sum_{k=1}^{\varphi} Z_k(\mathbf{B}, \mathbf{C}) = i :\ (\mathbf{B}, \mathbf{C}) \leftarrow\!\!\$\ \mathsf{Chall}'_\varphi(1^\kappa, x)\right].$$

Finally, let $k^* \in [0, \varphi - 1]$ be the first index for which the following holds:

$$\frac{k^* + 1}{\varphi}\delta_{\varphi - k^* - 1} \leqslant \frac{1}{2\rho} \sum_{i \geqslant \varphi - k^*} \delta_i. \tag{1}$$

Before proceeding with the proof, we establish that for any P' that succeeds with noticeable probability p' there exists an index k^* having the property required in Eq. (1). This can be seen as follows. Suppose that Eq. (1) does not hold then:

$$\begin{aligned}\delta_1 &> \frac{\varphi}{2\rho}\sum_{i=2}^{\varphi}\delta_i = \frac{\varphi}{2\rho}\left(\delta_2 + \sum_{i=3}^{\varphi}\delta_i\right)\\ &> \frac{\varphi}{2\rho}\left(\left(\frac{\varphi}{2\cdot 2\rho}\sum_{i=3}^{\varphi}\delta_i\right) + \sum_{i=3}^{\varphi}\delta_i\right) = \frac{\varphi}{2\rho}\left(\frac{\varphi}{2\cdot 2\rho}+1\right)\sum_{i=3}^{\varphi}\delta_i\\ &> \frac{\varphi}{2\rho}\cdot\prod_{i=2}^{\varphi-1}\left(1+\frac{\varphi}{i2\rho}\right)\cdot p' > \frac{\varphi}{2\rho}\cdot\left(1+\frac{1}{2\rho}\right)^{\varphi-3}\cdot p' > 1,\end{aligned}$$

where the last inequality holds if we set $\varphi = \omega(\log\kappa)\rho$. We reached a contradiction.

For any $i \in [\varphi]$, and for a parameter $\tau \in \mathbb{N}$, we define the procedure $\overline{\mathsf{Resp}}_\tau^{\mathsf{P}'}(1^\kappa, x, c_1, \ldots, c_j)$ that upon input an instance x and challenges $\mathbf{c} = (c_1, \ldots, c_j)$, and given oracle access to a prover P', outputs answers $a_1, \ldots, a_j$. The description of the procedure follows:

1. For τ many trials:
   - Sample $(\mathbf{C}, \mathbf{B})$ as in the collapsed protocol and choose $i^* \leftarrow\!\!\$\ [\varphi]$.
   - Using oracle access to prover $\mathsf{P}'(1^\kappa, \cdot)$, compute the value:

$$Z = \sum_{i\in[\varphi]\setminus\{i^*\}} Z_i(\mathbf{B}, \mathbf{C}_{|\mathbf{c}^{i^*,j}=\mathbf{c}}).$$

   If $Z \leqslant k^*$ break the loop and output $\mathbf{a}^{i^*,j}$.
2. Output $\perp$ if no answer has been found.

Notice that the index k^* depends only on the description of P', therefore we can assume that k^* is passed to P^* as auxiliary input. Consider now an execution of the ρ-round protocol between $\mathsf{P}^{*\mathsf{P}'}$ and V on common input x and prover's auxiliary input $z^* := (z, k^*)$, where the prover executes the procedure $\overline{\mathsf{Resp}}_\tau$ with oracle access to $\mathsf{P}'(1^\kappa, x, z)$ and the verifier samples $(\mathbf{b}, \mathbf{c}) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa, x)$. We can upper bound the probability that P^* fails in convincing V as follows:

$$\begin{aligned}\Pr\left[\langle \mathsf{P}^{*\mathsf{P}'}(1^\kappa, x, z^*), \mathsf{V}(1^\kappa, x)\rangle \neq \mathtt{acc}\right] &\leqslant \Pr\left[\exists j : \overline{\mathsf{Resp}}_\tau^{\mathsf{P}'}(1^\kappa, x, \mathbf{c}_{\downarrow j}) \neq \mathbf{b}_{\downarrow j}\right]\\ &\leqslant \sum_{j\in[\rho]} \Pr_{\overline{\mathsf{Resp}}_\tau, \mathbf{c}}\left[\overline{\mathsf{Resp}}_\tau^{\mathsf{P}'}(1^\kappa, x, \mathbf{c}_{\downarrow j}) \neq \mathbf{b}_{\downarrow j}\right]. \qquad (2)\end{aligned}$$

Next, we analyze the probability that $\overline{\mathsf{Resp}}_\tau$ forwards the right answer in a generic round $j \in [\rho]$. The analysis is facilitated by first looking at the algorithm $\overline{\mathsf{Resp}}_\infty$ that executes the main loop until it finds a valid answer; later we will show that the probability that the number of executions is less than τ is overwhelming in κ. In case the algorithm $\overline{\mathsf{Resp}}_\infty$ ends there are two possible situations:

- The algorithm accepts a wrong answer, let us call this event Fail. In this case, the number of errors made by P' is $k^* + 1$ (including the answer $\mathbf{a}^{i^*,j}$). Observe that, since i^* is uniformly random, at any iteration we have $\Pr[\mathsf{Fail}] = \delta_{\varphi-k^*-1}\frac{k^*+1}{\varphi}$.
- The algorithm accepts a correct answer, let us call this event $\mathsf{Success}$. In this case, the number of errors made by P' is at most k^*. Observe that, by definition, at any iteration $\Pr[\mathsf{Success}] = \sum_{i \geqslant \varphi-k^*} \delta_i$.

Recalling the definition of the index k^* from Eq. 1, we obtain:

$$\Pr[\mathsf{Fail}] \leqslant \frac{k^*+1}{\varphi}\delta_{\varphi-k^*-1} \leqslant \frac{1}{2\rho}\sum_{i \geqslant \varphi-k^*}\delta_i = \frac{1}{2\rho}\Pr[\mathsf{Success}],$$

and thus the probability that $\overline{\mathsf{Resp}}_\infty$ fails is less than $\frac{1}{2\rho}$. (We are conditioning on the event that $\overline{\mathsf{Resp}}_\infty$ terminates here.)

Since $\Pr[\mathsf{Success}] \geq \delta_\varphi = p'$, we have $\Pr[\mathsf{Fail}]+\Pr[\mathsf{Success}] \geqslant p'$. Hence, by the tail inequality of the geometric distribution the probability that $\overline{\mathsf{Resp}}_\infty$ makes more than $\tau = {p'}^{-1}\omega(\log\kappa)$ iterations is negligible. Therefore, we can upper bound the probability that $\overline{\mathsf{Resp}}_\tau$ fails by $\frac{1}{2\rho} + \mathsf{negl}(\kappa)$. Combining this with Eq. (2) we get that P^* is successful in the ρ-round protocol with probability $1 - \rho(\frac{1}{2\rho} + \mathsf{negl}(\kappa)) \geq 1/3$.

We finally define the f-knowledge extractor of protocol Π' on prover P' to be the same f-knowledge extractor of protocol Π on prover $\mathsf{P}^{*\mathsf{P}'}$.[4] Recall that the knowledge soundness of Π is $\epsilon(\kappa)$; moreover prover P^* succeeds with probability greater than $1/3$ whenever executed with oracle access to a prover P' for Π' that succeeds with noticeable probability. It follows that the knowledge soundness of Π' is $\epsilon^{\omega(\log\kappa)}$.

To complete the proof we show that the completeness error of the collapsed protocol is negligible. Notice that for any $j \in [\rho]$, given an instance $(x, w) \in R$:

$$1 - \mu(\kappa) \geqslant \Pr\left[\langle \mathsf{P}(1^\kappa, x, w), \mathsf{V}(1^\kappa, x)\rangle = \mathtt{acc}\right] = \Pr\left[\wedge_{i\in[\rho]}(a_i = b_i)\right] \geqslant \Pr\left[\wedge_{i\in[j]}(a_i = b_i)\right].$$

Therefore:

$$\Pr\left[\left\langle \mathsf{P}'(1^\kappa, x, w), \mathsf{V}'(x)\right\rangle = \mathtt{acc}\right] \geqslant (1-\mu)^{\rho\cdot\varphi}.$$

Recall that we set $\varphi = \omega(\log\kappa)\rho$ hence the right-hand side of the equation above is overwhelming.

3.3 On Laconic PAoK

We show that laconic PAoK (where the size of the prover's answer is $\ell = 1$ bit) are equivalent to PAoK.

Theorem 2. *Let R be an* NP *relation. There exists a PAoK for R if and only if there exists a* laconic *PAoK for R.*

[4] Strictly speaking, the definition of f-knowledge soundness does not consider provers with oracle access, however, since P' is efficient, we can define a new prover $\hat{\mathsf{P}}$ that executes P^* and emulates each invocation to the oracle using the code of P'.

Proof (Proof sketch.). Consider $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ to be a PAoK for R, with $\ell = \mathsf{poly}(\kappa)$. We write $(\mathbf{c}, \mathbf{b})$ for the output of $\mathsf{Chall}(1^\kappa, x)$ and $\mathbf{a}$ for the output of $\mathsf{Resp}(1^\kappa, x, w, \mathbf{c})$; note that $|\mathbf{a}| = |\mathbf{b}| = \ell$. Define the following laconic protocol $\Pi' = (\mathsf{Chall}', \mathsf{Resp}')$:

- Upon input $1^\kappa, x$, define $\mathsf{Chall}'(1^\kappa, x) := (c', b')$ where $c' = (\mathbf{c}, \mathbf{r})$ and $b' = \langle \mathbf{b}, \mathbf{r} \rangle$ for a random $\mathbf{r} \leftarrow\!\!\$\ \{0,1\}^\ell$.
- Upon input $1^\kappa, x, w, c'$, define $\mathsf{Resp}'(1^\kappa, x, w, c') := \langle \mathbf{a}, \mathbf{r} \rangle$ where $c' = (\mathbf{c}, \mathbf{r})$ and $\mathbf{a} = \mathsf{Resp}(1^\kappa, x, w, \mathbf{c})$.

Clearly, Π' is laconic. Notice that the bit b' is an hard-core predicate [12] for the relation $R_{\mathsf{ip}} = \{(\mathbf{b}, (\mathbf{c}, \mathbf{r})) : (\mathbf{c}, \mathbf{b}) \in \mathsf{Chall}(1^\kappa, x), \mathbf{r} \in \{0,1\}^\ell\}$. Moreover, the inverter of the Goldreich-Levin Theorem [10] works for any surjective function and, for any fixed function $g(b) := (\mathbf{c}, \mathbf{r})$ such that $(\mathbf{b}, (\mathbf{c}, \mathbf{r})) \in R_{\mathsf{ip}}$, it does not require the function $g(\cdot)$ to be efficiently computable in order to invert it. With this in mind, it is not hard to reduce knowledge soundness of Π' to the hardness of the Goldreich-Levin predicate. We skip the technical details.

We proceed to prove the other direction. Let $\Pi' = (\mathsf{Chall}', \mathsf{Resp}')$ be a laconic PAoK for the relation R, and consider the protocol Π^ℓ below which, for a parameter $\ell = \mathsf{poly}(\kappa)$, simply repeats Π' sequentially for $\rho(\kappa)$ times:

1. For $i \in [\rho(\kappa)]$ where $\rho(\kappa) = \omega(\ell^{1/2} / \log^{1/2} \kappa)$, proceed as follows:
   - Execute protocol Π';
   - If the execution aborts then output `rej`, otherwise proceed to the next iteration.
2. If all the ρ executions were accepting, then return `acc`.

It is well known that sequential repetition amplifies the knowledge soundness error of private-coin arguments of knowledge; thus Π^ℓ is a ρ-round PAoK with knowledge soundness error ϵ^ℓ. We can now apply Theorem 1 to obtain a one-round PAoK.

4 Constructing PAoK

We give two constructions of PAoK. In Section 4.1 we show that we can construct a PAoK from any extractable hash-proof system [20] (Ext-HPS); if the Ext-HPS is defined w.r.t. a relation R, we obtain a PAoK for a related relation R' where R and R' share the same x and the witness for x w.r.t. R' is the randomness used to sample the instance $(x, w) \in R$.

In Section 4.2 we investigate the relationship between PAoK and Extractable Witness Encryption [9,14];

4.1 Construction from Extractable Hash-Proof Systems

The definition below is adapted from [20].

Definition 3 (Ext-HPS). *Let $\mathcal{H} = \{h_{pk}\}$ be a set of hash functions indexed by a public key pk, and let R be an* NP*-relation. An extractable hash-proof system for R is a tuple of* ppt *algorithms $\Pi_{\mathsf{HPS}} := (\mathsf{SetupHash}, \mathsf{SetupExt}, \mathsf{Ext}, \mathsf{Pub}, \mathsf{Priv})$ such that the following properties are satisfied.*

Public evaluation: *For all* $(pk, sk) \leftarrow \mathsf{SetupExt}(1^\kappa)$*, and* $(x, w) \leftarrow \mathsf{SamR}(1^\kappa; r)$*, we have* $\mathsf{Pub}(1^\kappa, pk, r) = h_{pk}(u)$.

Extraction mode: *For all* $(pk, sk) \leftarrow \mathsf{SetupExt}(1^\kappa)$ *and all* (x, y)*, we have that* $\pi = h_{pk}(x) \Leftrightarrow (x, \mathsf{Ext}(1^\kappa, sk, x, \pi)) \in R$.

Hashing mode: *For all* $(pk, sk) \leftarrow \mathsf{SetupHash}(1^\kappa)$*, and for all* $(x, w) \in R$*, we have that* $\mathsf{Priv}(1^\kappa, sk, x) = h_{pk}(x)$.

Indistinguishability: *The ensembles* $\{pk : (pk, sk) \leftarrow \mathsf{SetupHash}(1^\kappa)\}_{\kappa\in\mathbb{N}}$ *and* $\{pk : (pk, sk) \leftarrow \mathsf{SetupExt}(1^\kappa)\}_{\kappa\in\mathbb{N}}$ *are statistically indistinguishable.*

Let R be an efficiently samplable relation with sampling algorithm SamR. Define the relation R', such that $(x, w') \in R'$ if and only if $(x, w) \in R$ where $(x, w) := \mathsf{SamR}(1^\kappa; w')$. Consider the following pair of ppt algorithms $\Pi = (\mathsf{Chall}, \mathsf{Resp})$, defining a one-round interactive argument for R' (as described in Section 3.1).

1. Algorithm $\mathsf{Chall}(1^\kappa, x)$ runs $(pk, sk) \leftarrow \mathsf{SetupHash}(1^\kappa)$, and defines $c := pk$ and $b := \mathsf{Priv}(1^\kappa, sk, x)$.
2. Algorithm $\mathsf{Resp}(1^\kappa, x, w', c)$ defines $a := \mathsf{Pub}(1^\kappa, pk, w')$.

Theorem 3. *Let* R*,* R' *and* SamR *be as above. Assume that* Π_{HPS} *is an Ext-HPS for the relation* R*. Then* $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ *as defined above is an* f*-extractable PAoK for the relation* R' *and where* $f(\cdot)$ *returns the second output of* $\mathsf{SamR}(\cdot)$.

Proof. Completeness follows by the correctness property of the hashing mode of the underlying HPS. In order to show knowledge soundness, we consider a mental experiment where algorithm Chall is defined differently. In particular, the verifier samples $(pk, sk) \leftarrow \mathsf{SetupExt}(1^\kappa)$ using the extraction mode instead of the hashing mode. By the indistinguishability property of the HPS this results in a statistically close distribution.

Now, we can define the extractor K of the PAoK as follows. Let P^* be a ppt algorithm such that $\langle \mathsf{P}^*(1^\kappa, x, z), \mathsf{V}(1^\kappa, x)\rangle_\Pi = \mathtt{acc}$ with probability $p(\kappa)$, where P^* uses auxiliary input $z \in \{0, 1\}^*$. Define $\mathsf{K}(1^\kappa, x, z) := \mathsf{Ext}(1^\kappa, sk, x, a)$ where a is the message sent by P^*. By definition of protocol Π we get that whenever P^* succeeds then $c = h_{pk}(x) = a$. Thus the extraction property of the HPS implies that $w \leftarrow\!\!\$\ \mathsf{Ext}(1^\kappa, sk, x, a)$ is a valid witness for x, i.e. $R(x, w) = 1$ with probability 1. The proof now follows by the fact that for all w' such that $\mathsf{SamR}(1^\kappa; w') = (x, w)$, we also have $R'(x, w') = 1$.

Instantiations. We consider two instantiations of Theorem 3, based on the constructions of Ext-HPS given in [20].

- The first construction is for the Diffie-Hellman relation $R_{\mathsf{DH}} := \{(g^r, g^{\alpha r}) : g \in \mathbb{G}, \alpha, r \in \mathbb{Z}_q\}$, for a group $\mathbb{G}$ of prime order q. Note that $\mathsf{SamR}(r) := (g^r, g^{\alpha r})$. The corresponding relation R'_{DH} is $R'_{\mathsf{DH}} := \{(g^r, r) : g \in \mathbb{G}, r \in \mathbb{Z}_q\}$, and $f(r) = f_\alpha(r) := g^{\alpha r}$.
- The second construction is based on factoring. Let $R_{\mathsf{QR}} := \{(g^{2^k r}, g^r) : g \leftarrow\!\!\$\ \mathbb{QR}_N^+, r \in [(N-1)/4]\}$, where N is a Blum integer. Note that $\mathsf{SamR}(r) := (g^{2^k r}, g^r)$. The corresponding relation R'_{QR} is $R'_{\mathsf{QR}} := \{(g^{2^k r}, r) : g \leftarrow\!\!\$\ \mathbb{QR}_N^+, r \in [(N-1)/4]\}$, and $f(r) := g^r$.

4.2 Equivalence to Extractable Witness Encryption

We show that full-fledged PAoK imply extractable witness encryption (Ext-WE), and viceversa. We start by recalling the definition of Ext-WE, taken from [8].

Extractable Witness Encryption. Let R be an NP-relation. A WE scheme $\Pi = (\mathsf{Encrypt}, \mathsf{Decrypt})$ for R (with message space $\mathcal{M} = \{0,1\}$) consists of two ppt algorithms, specified as follows:[5] (i) Algorithm $\mathsf{Encrypt}$ takes as input a security parameter 1^κ, a value $x \in \{0,1\}^*$, and a message $\beta \in \{0,1\}$, and outputs a ciphertext γ; (ii) Algorithm $\mathsf{Decrypt}$ takes as input a security parameter 1^κ, a ciphertext γ, a value $w \in \{0,1\}^*$, and outputs a message $\beta \in \{0,1\}$ or a special symbol $\perp$.

Definition 4 (Ext-WE). *Let R be an NP-relation, and $\Pi_{\mathsf{WE}} = (\mathsf{Encrypt}, \mathsf{Decrypt})$ be a WE scheme for R. We say that Π_{WE} is an Ext-WE scheme for R if the following requirements are met.*

Correctness: *For any $x \in L$ and $\beta \in \{0,1\}$, we have that $\mathsf{Decrypt}(1^\kappa, w, \mathsf{Encrypt}(1^\kappa, x, \beta)) = \beta$ with probability one, where $(x,w) \in R_L$.*

Extractable security: *For any ppt adversary $\mathsf{A} = (\mathsf{A}_0, \mathsf{A}_1)$ and for any noticeable function $\epsilon(\cdot)$, there exists a non-uniform extractor K and a non-zero polynomial $q(\cdot)$ such that the following holds. For any auxiliary information $z \in \{0,1\}^*$ and for any tuple $(x, st) \leftarrow_\$ \mathsf{A}_0(1^\kappa, z)$, whenever*

$$\Pr\left[\mathsf{A}_1(1^\kappa, st, x, \mathsf{Encrypt}(1^\kappa, x, \beta), z) = \beta : \ \beta \leftarrow_\$ \{0,1\}\right] \geqslant \frac{1}{2} + \epsilon(\kappa)$$

then we have $\Pr[(x, \mathsf{K}(1^\kappa, x, z)) \in R_L] \geqslant q(\epsilon(\kappa))$.

Theorem 4. *Let R be an NP-relation. There exists a PAoK for R if and only if there exists a Ext-WE scheme for R.*

Proof. Let $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ be a PAoK for the relation R. Without loss of generality, by our analysis in Section 3, we can assume that the PAoK is laconic (i.e., the output of Resp is a single bit $a \in \{0,1\}$). Consider the following construction of an Ext-WE scheme $\Pi_{\mathsf{WE}} = (\mathsf{Encrypt}, \mathsf{Decrypt})$ for R (with message space $\mathcal{M} = \{0,1\}$):

- Upon input $1^\kappa, x$ and message β, define $\mathsf{Encrypt}(1^\kappa, x, \beta) := (c, \beta \oplus b) := \gamma$ where $(c,b) \leftarrow_\$ \mathsf{Chall}(1^\kappa, x)$.
- Upon input $1^\kappa, w, \gamma$, where $\gamma = (\gamma_1, \gamma_2)$, define $\mathsf{Decrypt}(1^\kappa, w, \gamma) = \gamma_2 \oplus a$ where $a \leftarrow \mathsf{Resp}(1^\kappa, x, w, c)$.

Let $\mathsf{A} = (\mathsf{A}_0, \mathsf{A}_1)$ be an adversary for the WE scheme. Assume that there exists a noticeable function $\epsilon(\cdot)$ such that

$$\Pr\left[\mathsf{A}_1(1^\kappa, st, x, \gamma, z) = \beta : \ \beta \leftarrow_\$ \{0,1\}; \gamma \leftarrow_\$ \mathsf{Encrypt}(1^\kappa, x, \beta)\right] \geqslant \frac{1}{2} + \epsilon(\kappa)$$

for $(st, x) \leftarrow_\$ \mathsf{A}_0(1^\kappa, z)$ (where $z \in \{0,1\}^*$ is the auxiliary input). We use A to construct a prover P^* attacking knowledge soundness of Π. Prover P^* first runs $(st, x) \leftarrow_\$ \mathsf{A}_0(1^\kappa, z)$, and then interacts with the honest verifier of Π on common input x, as follows:

[5] WE for arbitrary-length messages can be obtained encrypting each bit of the plaintext independently.

1. Receive challenge c from the verifier.
2. Sample $\beta' \leftarrow\!\!\$\ \{0,1\}$ and run $\mathsf{A}_1(1^\kappa, st, x, \gamma, z)$ on $\gamma := (c, \beta')$, obtaining a bit β.
3. Send $a := \beta \oplus \beta'$ to the verifier.

For the analysis, note that the ciphertext simulated by P^* has the right distribution (in particular, the second component is a random bit). Since $\beta' = \beta \oplus b$ we get that P^* outputs $a = b$ with probability at least $1/2 + \epsilon(\kappa)$ and thus P^* convinces V with probability $p(\kappa) \geq 1/2 + \epsilon(\kappa)$. We are now in a position to run the extractor K of Π, and hence we obtain a valid witness $w \leftarrow\!\!\$\ \mathsf{K}(1^\kappa, x, z)$ with probability $q(\epsilon(\kappa))$. The statement follows. A similar argument shows that Π_{WE} is a weak Ext-WE whenever Π is a weak PAoK.

Conversely, let $\Pi_{\mathsf{WE}} = (\mathsf{Encrypt}, \mathsf{Decrypt})$ be an Ext-WE scheme for the relation R, with message space $\mathcal{M} = \{0,1\}$. Consider the following construction of a PAoK $\Pi = (\mathsf{Chall}, \mathsf{Resp})$:

- Upon input $1^\kappa, x$, define $\mathsf{Chall}(1^\kappa, x) := (\mathsf{Encrypt}(1^\kappa, x, b), b)$ where $b \leftarrow\!\!\$\ \{0,1\}$.
- Upon input $1^\kappa, x, w, c$, define $\mathsf{Resp}(1^\kappa, x, w, c) := \mathsf{Decrypt}(1^\kappa, w, c)$.

Fix any x and let P^* be a malicious prover for the PAoK. Assume that there exists a polynomial $p(\cdot)$ such that

$$p(\kappa) := \Pr\left[\langle \mathsf{P}^*(1^\kappa, x, z), \mathsf{V}(1^\kappa, x)\rangle = \mathtt{acc}\right] \geq \epsilon(\kappa).$$

where $z \in \{0,1\}^*$ is the auxiliary input. We use P^* to construct an adversary $\mathsf{A} := (\mathsf{A}_0, \mathsf{A}_1)$ attacking extractable security of Π_{WE}. Adversary $\mathsf{A}_0(1^\kappa, z)$ outputs x, and then A_1 is given a challenge ciphertext γ that is either an encryption of $\beta = 0$ or an encryption of $\beta = 1$ (under x), and its goal is to guess β. To do so A proceeds as follows:

1. Forward γ to P^*.
2. Let a be the answer sent by P^*; output $\beta := a$.

For the analysis, note that the challenge simulated by A_1 has the right distribution (in particular, it is a witness encryption of a random bit). Since $a = b = \mathsf{Decrypt}(1^\kappa, x, w, \gamma)$ with probability at least $p(\kappa)$, we get that A_1 guesses β with at least the same probability. We are now in a position to run the extractor K of Π_{WE}, and hence we obtain a valid witness $w \leftarrow\!\!\$\ \mathsf{K}(1^\kappa, x, z)$ with probability $q(p(\kappa) - \epsilon(\kappa))$. The statement follows. A similar argument shows that Π is a weak PAoK whenever Π_{WE} is a weak Ext-WE.

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A Application to Leakage-Tolerant Secure Message Transmission

In this section we show an application of PAoK. We show that if you use a public-key encryption scheme to implement secure message transmission and the resulting protocol can tolerate leakage of even a constant number of bits, then the scheme will have to have secret keys which are as long as the total number of bits transmitted using that key. Since the typical interpretation of public-key encryption is that an unbounded number of messages can be encrypted using a fixed public key, this shows that typical public-key encryption cannot be used to realize leakage-tolerant secure message transmission even against a constant number of leaked bits.

This strengthens an earlier result which showed this was the case for schemes tolerating a super-logarithmic number of bits of leakage. For the proof we need that there exists a PAoK for a relation associated to the public-key encryption scheme, essentially proving knowledge of the secret key. The result can therefore be interpreted as showing that either a public-key encryption does not give secure message transmission secure against a constant number of bits of leakage or there does not exists a PAoK for the applied public-key encryption scheme. Proving that such a PAoK does not exists seems very challenging using current techniques, so the result indicates that we probably cannot base leakage-tolerant message transmission secure against leaking a constant number of bits on public-key encryption with our current proof techniques.

Functionality $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$

Running with parties $\mathsf{Rec}, \mathsf{Sen}$ and adversary S, the functionality $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$ is parametrized by the security parameter κ, message space $\mathcal{M}$ and the set of all admissible leakage functions Φ. Hence, $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$ behaves as follows:

- Upon input $(\mathtt{send}, \mathsf{Sen}, \mathsf{Rec}, m)$ send a message $(\mathtt{send}, \mathsf{Sen}, \mathsf{Rec}, |m|)$ to S. Once S allows to forward the message, send $(\mathtt{sent}, \mathsf{Sen}, m)$ to Rec.
- Upon input $(\mathtt{leak}, X, \phi_{\mathsf{Z}})$ for $X \in \{\mathsf{Sen}, \mathsf{Rec}\}$ and $\phi_{\mathsf{Z}} \in \Phi$ send a message $(\mathtt{leak}, X)$ to S. Receive $(\mathtt{leak}, X', \phi_{\mathsf{S}})$ from S, check that $\phi_{\mathsf{S}} \in \Phi$, and that $|\phi_{\mathsf{Z}}(\cdot)| = |\phi_{\mathsf{S}}(\cdot)|$ and $X' = X$. Send $(\mathtt{leak}, \phi_{\mathsf{S}}(m))$ to S and $(\mathtt{leaked}, |\phi_{\mathsf{S}}(m)|)$ to X'.

Fig. 1. Ideal functionality $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$ for secure message transmission with leakage

Syntax of PKE. A tuple of algorithms $(\mathsf{KGen}, \mathsf{Enc}, \mathsf{Dec})$ is said to be a PKE scheme for message space $\mathcal{M}$ if the following holds: (i) Algorithm KGen takes as input the security parameter and returns a pair (pk, sk); (ii) Algorithm Enc takes as input a public key pk and a message $m \in \mathcal{M}$, and outputs a ciphertext γ; (iii) Algorithm Dec takes as input a secret key sk and a ciphertext γ, and outputs a message $m \in \mathcal{M}$ or $\perp$. We require that for all $m \in \mathcal{M}$ one has $\mathsf{Dec}(sk, \mathsf{Enc}(pk, m)) = m$ with overwhelming probability over the randomness of $(\mathsf{KGen}, \mathsf{Enc}, \mathsf{Dec})$.

Message transmission with leakage. We recall the notion of leakage tolerance—introduced by Bitansky *et al.* [2]—for secure message transmission. Informally, in a leakage-tolerant message transmission protocol leakage queries from an adversary A are viewed as a form of partial corruptions, where A does not receive the complete state of the chosen party but just some function of it. Security is then achieved if such an adversary can be simulated in the UC framework [4]. Without loss of generality we will consider only dummy adversaries—adversaries which just carry out the commands of the environment. I.e., it is the environment which specifies all leakage queries. We will therefore completely drop the adversary in the notation for clarity.

Let Π_{PKE} be a protocol between a sender Sen and a receiver Rec. The ideal-world functionality for secure message transmission with leakage is depicted in Fig. 1. We say that Π_{PKE} is a leakage-tolerant secure implementation of $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$ if there exists a simulator S such that no environment Z can distinguish between the real life protocol Π_{PKE} and S interacting with the ideal functionality $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$. We denote with $\mathbf{IDEAL}_{\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}, \mathsf{S}, \mathsf{Z}}(\Phi, \kappa)$ the output of the environment Z when interacting with simulator S in the simulation.

Consider the following protocol Π_{PKE} between a sender Sen and a receiver Rec, supposed to realize $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$ via a public-key encryption scheme $(\mathsf{KGen}, \mathsf{Enc}, \mathsf{Dec})$ with message space $\mathcal{M}$, assuming authenticated channels:

1. Sen transmits to Rec that it wants to forward a message $m \in \mathcal{M}$;
2. Rec samples $(pk, sk) = \mathsf{KGen}(1^{\kappa}; r_G)$, and sends pk to Sen;

3. Sen computes $\gamma = \mathsf{Enc}(pk, m; r_E)$ and forwards the result to Rec;
4. Rec outputs $m' = \mathsf{Dec}(sk, \gamma; r_D)$.

Note that at the end of the execution of Π_{PKE} the state of Sen is $\sigma_S = (m, r_E)$ whereas the state of Rec is $\sigma_R = (sk, r_G, r_D, m')$. Denote with $\mathbf{REAL}_{\Pi_{\mathsf{PKE}},\mathsf{Z}}(\Phi, \kappa)$ the output of the environment Z after interacting with parties Rec, Sen in a real execution of Π_{PKE}.

Definition 5 (Leakage-tolerant PKE protocol). *We say that Π_{PKE} is a (λ, ϵ)-leakage-tolerant PKE protocol (w.r.t. a set of leakage functions Φ) if Π_{PKE} securely implements $\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}}$, i.e., there exists a probabilistic polynomial-time simulator* S *such that for any environment* Z *leaking at most λ bits of information it holds that*

$$\{\mathbf{IDEAL}_{\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}},\mathsf{S},\mathsf{Z}}(\Phi, \kappa)\}_{\kappa\in\mathbb{N}} \approx_{\epsilon} \{\mathbf{REAL}_{\Pi_{\mathsf{PKE}},\mathsf{Z}}(\Phi, \kappa)\}_{\kappa\in\mathbb{N}}.$$

Necessity of long keys. Nielsen *et al.* [17] have shown that any leakage-tolerant PKE protocol (as per Definition 5) requires long keys already when Π_{PKE} tolerates super-logarithmic leakage. Below we strengthen their result proving a more fine-grained lower bound for any $\lambda = O(1)$ bits of leakage.

Consider the following NP relation, depending upon a PKE scheme $(\mathsf{KGen}, \mathsf{Enc}, \mathsf{Dec})$:

$$R_{\mathsf{PKE}} := \{((pk, \gamma, m), (sk, r_G)) : \ (pk, sk) = \mathsf{KGen}(1^{\kappa}; r_G) \wedge \mathsf{Dec}(sk, \gamma) = m\} \tag{3}$$

We show the following theorem.

Theorem 5. *Let $\Pi = (\mathsf{Chall}, \mathsf{Resp})$ be a PAoK for the above relation R_{PKE}, with knowledge soundness error $2^{-\ell} + \epsilon$, perfect completeness, and prover's answers of length $\ell \geq 1$. Let Π_{PKE} be an (ℓ, ϵ')-leakage-tolerant PKE protocol. Then it must be that $|\mathcal{SK}| \geqslant (1 - 2^{-\ell} - \epsilon - \mu - 2\epsilon')|\mathcal{M}|$, where $\mathcal{SK}$ is the space of all secret keys.*

Before we sketch the proof, we give an interpretation of the Theorem 5. When the knowledge soundness of the PAoK is negligible and the PKE protocol is (ℓ, negl)-leakage-tolerant it means that $|\mathcal{SK}| \geqslant (1 - 2^{-\ell+1})|\mathcal{M}|$

Proof (Proof sketch.). We make the proof for the case where the decryption algorithm is deterministic and has perfect correctness. Probabilistic decryption can be handled as in [17].

We construct an environment Z which uses ℓ bits of leakage on the receiver's state after the execution of Π, for which the existence of a simulator S implies our bound. The environment Z works as follows:

1. Input a uniformly random $m \in \mathcal{M}$ to Sen.
2. Let the protocol terminate without any leakage queries or any corruptions, i.e., simply deliver all messages between Sen and Rec. As part of this Z learns pk and γ from observing the authenticated channel between Sen and Rec.

3. After the protocol terminates, let Rec prove via leakage queries that $x := (pk, c, m) \in L_{R_{\mathsf{PKE}}}$ using $\Pi = (\mathsf{Chall}, \mathsf{Resp})$. Notice that Rec can do this as it knows a valid witness $w := (sk, r_G)$. More in detail, Z runs $(c, b) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa, x)$ and specifies a single leakage query defined as $\phi_{\mathsf{Z}}^{x,c}(\sigma_{\mathsf{Rec}}) = \phi_{\mathsf{Z}}^{x,c}(w) = \mathsf{Resp}(1^\kappa, x, w, c)$ (the values x and c are hard-wired into the function).
4. Finally Z outputs 1 if and only if $a = b$, where a is the output of Z's leakage query.

Note that the total amount of leaked information is equal to the communication complexity of the prover in Π, i.e., ℓ bits. By completeness of the PAoK, we know that $\mathbf{REAL}_{\Pi_{\mathsf{PKE}},\mathsf{Z}}(\Phi, \kappa) = 1$. From this we conclude that $\mathbf{IDEAL}_{\mathbf{F}_{\mathsf{SMT}}^{+\mathsf{lk}},\mathsf{S},\mathsf{Z}}(\Phi, \kappa) = 1$ except with negligible probability, by security of the protocol. We write out what this means. The simulation proceeds as follows:

1. First Z inputs a uniformly random $m \in \mathcal{M}$ to the ideal functionality on behalf of Sen. As a result S is given $(\mathtt{send}, \mathsf{Sen}, \mathsf{Rec}, |m|)$.
2. Then S must simulate the communication of the protocol, which in particular means that it must output some pk and γ to Z.
3. After the simulation of the protocol terminates, the environment runs $(c, b) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa, x)$ and makes the leakage query $\phi_{\mathsf{Z}}^{x,c}$ with which Rec proves that $x = (pk, \gamma, m) \in L_{R_{\mathsf{PKE}}}$. Such leakage query is answered by S using another function ϕ_{S} producing a value a.
4. Finally Z outputs 1 if and only if $a = b$

Since $\mathbb{Z}$ is computing (c, b) as the verifier of Π would have done, and the value a is computed by S which is PPT, and since Z outputs 1, it follows from soundness that $x \in L_{R_{\mathsf{PKE}}}$ except with probability $\epsilon = 2^{-\ell} + \mathsf{negl}(\kappa)$. This means that there exist (sk, r_G) such that $(pk, sk) = \mathsf{KGen}(1^\kappa; r_G)$ and $m = \mathsf{Dec}(sk, \gamma)$. In particular, there exists $sk \in \mathcal{SK}$ such that $m = \mathsf{Dec}(sk, \gamma)$. Let $M_{pk,\gamma} \subset \mathcal{M}$ denote the subset of $m' \in \mathcal{M}$ for which there exist $sk' \in \mathcal{SK}$ such that $m' = \mathsf{Dec}(sk', \gamma)$. An argument identical to the one in [17, Theorem 1] shows that $m \in M_{pk,\gamma}$ except with probability $\epsilon + \epsilon'$. Combined with the above, this implies that $|M_{pk,\gamma}| \geq (1 - 2^{-\ell} - \mathsf{negl}(\kappa) - 2\epsilon')|\mathcal{M}|$. Take two $m_0 \neq m_1 \in M_{pk,\gamma}$. By definition there exist $sk_0, sk_1 \in \mathcal{SK}$ such that $m_0 = \mathsf{Dec}(sk_0, \gamma)$ and $m_1 = \mathsf{Dec}(sk_1, \gamma)$. From $m_0 \neq m_1$, we conclude that $sk_0 \neq sk_1$, so $|\mathcal{SK}| \geq |M_{pk,\gamma}|$. From this we get the theorem.

B Adaptive-Secure PAoK

In this section we consider adaptive-secure PAoK. Such protocols are fully specified by a tuple of three ppt algorithms $\Pi = (\mathsf{Chall}, \mathsf{Resp}, \mathsf{TResp})$ as described below:

1. V samples $(c, tp) \leftarrow\!\!\$\ \mathsf{Chall}(1^\kappa)$ and sends c to P.
   (Notice that the generated tuple is independent of the instance.)
2. P samples $a \leftarrow\!\!\$\ \mathsf{Resp}(1^\kappa, x, w, c)$ and sends a to V.
3. V computes $b := \mathsf{TResp}(tp, x)$ and outputs `acc` if $a = b$, else `rej`.

Before giving the definition, we set up some useful notation. Consider the following experiment $\mathsf{Exp}_{\mathsf{P}^*,\Pi}^{\mathsf{adp}}(\kappa, z)$ between a challenger and a prover P^*:

1. The challenger picks $(c, tp) \leftarrow\!\!\$\ \Pi.\mathsf{Chall}(1^\kappa)$;

2. The prover P^* upon inputs c, the auxiliary input z and randomness $r \leftarrow\!\!\$\ \{0,1\}^*$ produces an instance x such that $|x| = \kappa$ and an answer a;
3. The challenger computes $b := \mathsf{TResp}(tp, x)$, if $a = b$ then $\mathtt{Exp}^{\mathsf{adp}}_{\mathsf{P}^*,\Pi}(\kappa, z)$ outputs $\mathtt{acc}$, else 0.

Similarly, let $\mathtt{Exp}^{\mathsf{adp-e}}_{\mathsf{P}^*,\mathsf{K}}(\kappa, z)$ be the following experiment:

1. The challenger picks $(c, tp) \leftarrow\!\!\$\ \Pi.\mathsf{Chall}(1^\kappa)$;
2. The prover P^* upon inputs c, the auxiliary input z and randomness $r \leftarrow\!\!\$\ \{0,1\}^*$ produces an instance x such that $|x| = \kappa$ and an answer a;
3. The challenger computes $w \leftarrow\!\!\$\ \mathsf{K}(tp, z)$, if $(x, w) \in R$ then $\mathtt{Exp}^{\mathsf{adp-e}}_{\mathsf{P}^*,\mathsf{K}}(\kappa, z)$ outputs $\mathtt{acc}$, else 0.

Definition 6 (Adaptive-Secure PAoK). *Let $\Pi = (\mathsf{Chall}, \mathsf{Resp}, \mathsf{TResp})$ be as specified above, and let R be an* NP *relation. Consider the properties below.*

Completeness: *There exists a negligible function μ such that:*

$$\Pr_{c,tp}\left[\forall (x,w) \in R : \mathsf{TResp}(tp, x) = \mathsf{Resp}(1^\kappa, x, w, c)\right] \geqslant 1 - \mu(\kappa),$$

where the probability is taken over the outcomes of $\mathsf{Chall}(1^\kappa)$.

Knowledge soundness with error ϵ: *For all* ppt *provers P^* there exists a non-uniform extractor K and a non-zero polynomial $q(\cdot)$ such that for any auxiliary input $z \in \{0,1\}^*$ the following holds. Whenever $p(\kappa, z) := \Pr[\mathtt{Exp}^{\mathsf{adp}}_{\mathsf{P}^*,\Pi}(\kappa, z) = \mathtt{acc}] > \epsilon(\kappa)$, then*

$$\Pr\left[\mathtt{Exp}^{\mathsf{adp-e}}_{\mathsf{P}^*,\mathsf{K}}(\kappa, z) = \mathtt{acc}\right] \geqslant q(p(z, \kappa) - \epsilon(\kappa)).$$

Let ℓ be the size of the prover's answer, we call Π a adaptive-secure predictable argument of knowledge (PAoK) for R if Π satisfies completeness and knowledge soundness for any efficient computable function f, and moreover $\epsilon - 2^{-\ell}$ is negligible. We call it a laconic *adaptive-secure PAoK if $\ell = 1$.*

B.1 A protocol based on Extractable Witness PRF.

An Extractable Witness Pseudo Random Function (Ext-WPRF) (see Zhandry []) is a tuple of algorithms $\Pi_{\mathsf{WPRF}} := (\mathsf{KGen}, \mathsf{F}, \mathsf{Eval})$ defined as follow:

- Upon input 1^κ and a relation R the key generation algorithm KGen samples a public evaluation key ek and a secret key fk.
- Upon inputs fk and an instance $x \in \{0,1\}^\kappa$ the PRF F produces as output y.
- Upon input pk, x, w the public evaluation algorithm Eval computes the value $F(fk, x)$ if w is a valid witness for x, $\perp$ otherwise. public evaluation key pk and a secret evaluation key fk.

We say that Π_{WPRF} is complete if $\mathsf{F}(fk, x) = \mathsf{Eval}(ek, x, w)$ for all $(x, w) \in R$ and all fk, ek generated by KGen. Consider the following experiment $\mathsf{WPRF}.\mathsf{Exp}^R_{\mathsf{A}}(b, \kappa)$ parameterized by a relation $R : \mathcal{X} \times \mathcal{W} \to \{0,1\}$, the security parameter κ and a bit b:

1. Run $fk, pk \leftarrow\!\!\$\ \mathsf{KGen}(1^\kappa, R)$;
2. A on input pk can make a single challenge query on instance $x^* \in \mathcal{X}$. The challenger computes $y_0 := \mathsf{F}(fk, x^*)$ and $y_1 \leftarrow\!\!\$\ \{0,1\}^\kappa$ and responds with y_b.
3. A produces a bit b', The experiment outputs b'

Let $\mathsf{Adv}^{\mathsf{WPRF}}_{R,\mathsf{A}} := \big|\Pr[\mathsf{WPRF.Exp}^R_{\mathsf{A}}(0,\kappa) = 1] - \Pr[\mathsf{WPRF.Exp}^R_{\mathsf{A}}(1,\kappa) = 1]\big|$. Consider the following experiment $\mathsf{WPRF.Exp}^{R,e}_{\mathsf{A},\mathsf{K}}(\kappa)$ parameterized by a relation R and the security parameter κ :

1. Run $fk, pk \leftarrow\!\!\$\ \mathsf{KGen}(1^\kappa, R)$;
2. A on input pk, auxiliary z and randomness r can make a single challenge query on instance $x^* \in \mathcal{X}$. The challenger computes $w \leftarrow\!\!\$\ \mathsf{K}(ek, z, r)$ and $y_1 \leftarrow\!\!\$\ \{0,1\}^\kappa$ and responds with y_b.
3. A produces a bit b', The experiment outputs b'

Definition 7. *$\Pi_{\mathsf{WPRF}} = (\mathsf{KGen}, \mathsf{F}, \mathsf{Eval})$ is* adaptive instance extracting non-interactively secure *for a relation R if, for all* ppt *adversaries* A *such that* $\mathsf{Adv}^{\mathsf{WPRF}}_{R,\mathsf{A}} \geqslant p(\kappa)$ *for an inverse polynomial p, there is a* ppt *extractor* K *and an inverse polynomial q such that* $\Pr[\mathsf{WPRF.Exp}^{R,e}_{\mathsf{A},\mathsf{K}}(\kappa)] \geqslant q(\kappa)$.

Let $\Pi_{\mathsf{WPRF}} = (\mathsf{KGen}, \mathsf{F}, \mathsf{Eval})$ be an Ext-WPRF scheme. Consider the following construction of an adaptive-secure PAoK $\Pi = (\mathsf{Chall}, \mathsf{Resp}, \mathsf{TResp})$ for a NP relation R:

- Upon input 1^κ, define $\mathsf{Chall}(1^\kappa) := \Pi_{\mathsf{WPRF}}.\mathsf{KGen}(1^\kappa, R)$.
- Upon input $1^\kappa, x, w, c := ek$, define $\mathsf{Resp}(1^\kappa, x, w, c) := \Pi_{\mathsf{WPRF}}.\mathsf{Eval}(c, x, w)$.
- Upon input $tp, x := fk$, define $\mathsf{TResp}(tp, x) := \Pi_{\mathsf{WPRF}}.\mathsf{F}(tp, x)$.

Theorem 6. *Assume that Π_{WPRF} is an Ext-WPRF for the relation R. Then $\Pi = (\mathsf{Chall}, \mathsf{Resp}, \mathsf{TResp})$ as defined above is an adaptive-secure PAoK for the relation R.*

Proof. Suppose there exists an prover P^* and an auxiliary input z such that $p(\kappa, z) := \Pr[\mathbf{Exp}^{\mathsf{adp}}_{\mathsf{P}^*,\Pi}(\kappa, z) = \mathtt{acc}] > \mathsf{negl}(\kappa)$ but for any non-uniform K we have

$$\Pr\left[\mathbf{Exp}^{\mathsf{adp-e}}_{\mathsf{P}^*,\mathsf{K}}(\kappa, z) = \mathtt{acc}\right] < \mathsf{negl}(\kappa). \tag{4}$$

Notice that $p(\kappa, z)$ is a polynomial. Consider the adversary A against the WPRF experiment that upon input pk runs $(x, a) \leftarrow\!\!\$\ \mathsf{P}^*(c := pk)$ and forwards x. Then, on challenge y checks if $a = y$ and output 1, else 0. Notice that :

$$\begin{aligned}\mathsf{Adv}^{\mathsf{WPRF}}_{R,\mathsf{A}} &= \big|\Pr[\mathsf{WPRF.Exp}^R_{\mathsf{A}}(0,\kappa) = 1] - \Pr[\mathsf{WPRF.Exp}^R_{\mathsf{A}}(1,\kappa) = 1]\big| \\ &= \big|\Pr[\mathbf{Exp}^{\mathsf{adp}}_{\mathsf{P}^*,\Pi}(\kappa, z) = \mathtt{acc}] - 2^{-\kappa}\big| = p(\kappa, z) - 2^{-\kappa}\end{aligned}$$

Note that the equation together with Eq. (4) is in contradiction with the hypothesis of the theorem[6].

[6] Specifically, the class of extractor for extractable WPRF is a subset of the class of extractor for adaptive-secure PAoK.