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Chapter 1

Introduction

1.1 Security Amplification Theorems

Introduction to security amplification theorems and hardness implication statements. Example of classical results. Problems captured by weakly verifiable puzzles. Contribution of this thesis.

1.2 Weakly verifiable puzzles

1.3 Contribution of the Thesis

1.4 Organization of the Thesis

Overview of the content of the succeeding chapters.

Chapter 2

Preliminaries

In this chapter we set up notation and terminology used in the Thesis.

2.1 Notation and Definitions

(Functions) We write $\text{poly}(\alpha_1, \dots, \alpha_n)$ to denote a polynomial on variables $\alpha_1, \dots, \alpha_n$. We call a function $f : \mathbb{N} \rightarrow \mathbb{R}$ *negligible* if for every polynomial $\text{poly}(n)$ there exists $n_0 \in \mathbb{N}$ such that for all $n \in \mathbb{N} : n > n_0$ the following holds

$$f(n) < \frac{1}{\text{poly}(n)}.$$

On the other hand, we say that a function $f : \mathbb{N} \rightarrow \mathbb{R}$ is *non-negligible* if there exists a polynomial $\text{poly}(n)$ such that for some $n_0 \in \mathbb{N}$ and for all $n \in \mathbb{N} : n > n_0$ we have

$$f(n) \geq \frac{1}{\text{poly}(n)}.$$

We note that it is possible that a function $f : \mathbb{N} \rightarrow \mathbb{R}$ is neither negligible nor non-negligible.

We say that a function f is *efficiently computable* if there exists a polynomial time algorithm computing f .

(Algorithms, Bitstrings, and Circuits) We denote Boolean circuits using capital letters from the Greek or English alphabet. We define a *probabilistic circuit* as a Boolean circuit $C_{m,n} : \{0,1\}^m \times \{0,1\}^n \rightarrow \{0,1\}^*$. We write $C_{m,n}(x;r)$ to denote a probabilistic circuit taking as input $x \in \{0,1\}^m$ and as auxiliary input $r \in \{0,1\}^n$. If a probabilistic circuit does not take any input, we slightly abuse notation and write $C_n(r)$. Similarly, we use $\{C_n\}_{n \in \mathbb{N}}$ to denote a family of probabilistic circuits that takes only auxiliary input. We

make sure that it is clear from the context that probabilistic circuits (families of probabilistic circuits) with only auxiliary input are not confused with deterministic Boolean circuits (families of Boolean circuits respectively). For a (probabilistic) circuit C we write $\text{Size}(C)$ to denote the total number of vertices of C . A *(probabilistic) polynomial size circuit* is a (probabilistic) circuit of size polynomial in the number of input vertices (including auxiliary input vertices if the circuit is probabilistic). We define a *two phase circuit* $C := (C_1, C_2)$ as a circuit where in the first phase a circuit C_1 is used and in the second phase a circuit C_2 .¹ If C_1 and C_2 are probabilistic circuits we write $C(\delta) := (C_1, C_2)(\delta)$ to denote that in both phases C_1 and C_2 take as auxiliary input the same bitstring δ .

For an algorithm A we write $\text{Time}(A)$ to denote the number of steps it takes to execute A . We say that A runs in *polynomial time* if $\text{Time}(A)$ is bounded by some $\text{poly}(|x|)$ where $|x|$ denotes the length of the input that A takes.

Similarly, as for probabilistic circuits we often write randomness used by a probabilistic algorithm explicitly as a bitstring provided as an auxiliary input.

We denote a tuple (x_1, \dots, x_l) by $x^{(l)}$. The i -th element of $x^{(l)}$ is denote by $x_i^{(l)}$. Furthermore, for tuples $x^{(l)}, y^{(k)}$ we use $x^{(l)} \circ y^{(k)}$ to denote the concatenation of $x^{(l)}$ and $y^{(k)}$ which results in a tuple $(x_1, \dots, x_l, y_1, \dots, y_k)$.

(Probabilities and distributions) For a finite set \mathcal{R} we write $r \xleftarrow{\$} \mathcal{R}$ to denote that r is chosen from \mathcal{R} uniformly at random. For $\delta \in \mathbb{R} : 0 \leq \delta \leq 1$ we write μ_δ to denote the Bernoulli distribution where outcome 1 occurs with probability δ and 0 with probability $1 - \delta$. Moreover, we use μ_δ^k to denote the probability distribution over k -tuples where each element of a k -tuple is drawn independently according to μ_δ . Finally, let $u \leftarrow \mu_\delta^k$ denote that a k -tuple u is chosen according to μ_δ^k .

Let $(\Omega_n, \mathcal{F}_n, \text{Pr})$ be a probability space and $n \in \mathbb{N}$. Furthermore, let $E_n \in \mathcal{F}_n$ denote an event that probability depends on n . We say that E_n happens *almost surely* or with *high probability* if $\text{Pr}[E_n] \geq 1 - 2^{-n} \text{poly}(n)$.

TODO: Define non-negligible probability

(Interactive protocols) We are often interested in situations where two probabilistic circuits interact with each other according to some protocol. We limit ourselves to cases where messages are representable by bitstrings. Let $\{A_n\}_{n \in \mathbb{N}}$ and $\{B_n\}_{n \in \mathbb{N}}$ be families of circuits such that $A_n : \{0, 1\}^* \rightarrow \{0, 1\}^*$ and $B_n : \{0, 1\}^* \rightarrow \{0, 1\}^*$. An *interactive protocol* is defined by $\{A_n\}_{n \in \mathbb{N}}$ and $\{B_n\}_{n \in \mathbb{N}}$ where for random bitstrings $\rho_A \in \{0, 1\}^n$, $\rho_B \in \{0, 1\}^n$ in the first round $m_0 := A_n(\rho_A)$ and in the second round $m_1 := B_n(\rho_B, m_0)$. In

¹ Where what a *phase* means should be clear from the context in which C is used.

general in the $(2k-1)$ -th round we have $m_{2k-2} := A_n(\rho_A, m_1, \dots, m_{2k-3})$ and in the $2k$ -th round $m_{2k-1} := B_n(\rho_B, m_1, \dots, m_{2k-2})$. The protocol execution between two probabilistic circuits A and B is denoted by $\langle A, B \rangle$. The output of A in the protocol execution is denoted by $\langle A, B \rangle_A$ and of B by $\langle A, B \rangle_B$. A sequence of all messages send by A and B in the protocol execution is called a *communication transcript* and is denoted by $\langle A, B \rangle_{trans}$.

The time complexity of the protocol depends on the number of rounds of the protocol and time needed to evaluate A_n and B_n . Exemplary, the protocol runs in polynomial time if the number of rounds and time needed to evaluate A_n and B_n is bounded by some $poly(n)$.

(Oracle algorithms) We use notion of *oracle circuits* following the standard definition included in the literature [Gol04]. If a circuit A gains oracle access to a circuit B , we write A^B . If additionally B gains oracle access to a circuit C we write A^{B^C} . However, to shorten notation, we often write A^B instead and make sure that it is clear from the context which oracle is accessed by B .

In many situations when studying time complexity of algorithms with oracle access we count oracle call as a single step. We emphasize this by writing that an algorithm has a certain time complexity *with oracle calls*. On the other hand, in some settings we are interested in giving a more rigorous bounds on running time of an algorithm. In these situations we compute running time of an algorithm explicitly with regard to time needed for accessing oracle.

TODO: Exponential factor in what?
TODO: Cite someone?
TODO: Define statistical distance etc.

Definition 2.1 (Polynomial time sampleable distribution) *We say that a distribution is polynomial time sampleable if it can be approximated by an algorithm running in time $poly(\log |\mathcal{D}|, \log |\mathcal{R}|)$ up to an exponential factor.*

Definition 2.2 (Pairwise independent family of efficient hash functions) *Let \mathcal{D} and \mathcal{R} be finite sets and \mathcal{H} be a family of functions mapping values from \mathcal{D} to values in \mathcal{R} . We say that \mathcal{H} is a family of pairwise independent efficient hash functions if \mathcal{H} has the following properties.*

(Pairwise independent) *For $\forall x \neq y \in \mathcal{D}$ and $\forall \alpha, \beta \in \mathcal{R}$, it holds*

$$\Pr_{hash \leftarrow \mathcal{H}}[hash(x) = \alpha \mid hash(y) = \beta] = \frac{1}{|\mathcal{R}|}.$$

(Polynomial time sampleable) *For every $hash \in \mathcal{H}$ the function $hash$ is sampleable in time $poly(\log |\mathcal{D}|, \log |\mathcal{R}|)$.*

(Efficiently computable) For every $hash \in \mathcal{H}$ there exists an algorithm running in time $\text{poly}(\log |\mathcal{D}|, \log |\mathcal{R}|)$ which on input $x \in \mathcal{D}$ outputs $y \in \mathcal{R}$ such that $y = hash(x)$.

We note that the pairwise independence property is equivalent to

$$\Pr_{hash \leftarrow \mathcal{H}}[hash(x) = \alpha \wedge hash(y) = \beta] = \frac{1}{|\mathcal{R}|^2}.$$

It is well known [CW77] that there exists families of functions meeting the criteria stated in Definition 2.2.

Chapter 3

Weakly Verifiable Cryptographic Primitives

This chapter gives an overview of weakly verifiable cryptographic primitives. We start by formulating a definition of a *dynamic interactive weakly verifiable puzzle* in Section 3.1. To provide the Reader more intuition in Section 3.2 we describe a series of well known cryptographic primitives that are weakly verifiable. The Section 3.3 is devoted to the previous research concerning different types of weakly verifiable puzzles.

3.1 Dynamic Interactive Weakly Verifiable Puzzle

We define a *dynamic interactive weakly verifiable puzzle* as follows.

TODO: Γ_H and Γ_V should be probabilistic?

Definition 3.1 (Dynamic Interactive Weakly Verifiable Puzzle.) A dynamic interactive weakly verifiable puzzle (DIWVP) is defined by a family of probabilistic circuits $\{P_n\}_{n \in \mathbb{N}}$. A circuit belonging to $\{P_n\}_{n \in \mathbb{N}}$ is called the problem poser. The solver $C := (C_1, C_2)$ for P_n is a probabilistic two phase circuit. We write $P_n(\pi)$ to denote the execution of P_n with the randomness fixed to $\pi \in \{0, 1\}^n$ and $C(\rho) := (C_1, C_2)(\rho)$ to denote the execution of both C_1 and C_2 with the randomness fixed to $\rho \in \{0, 1\}^*$.

In the first phase, the problem poser $P_n(\pi)$ and the solver $C_1(\rho)$ interact. As the result of the interaction $P_n(\pi)$ outputs a verification circuit Γ_V and a hint circuit Γ_H . The circuit $C_1(\rho)$ produces no output. The circuit Γ_V takes as input $q \in Q$, an answer $y \in \{0, 1\}^*$ and outputs a bit. We say that an answer (q, y) is a correct solution if and only if $\Gamma_V(q, y) = 1$. The circuit Γ_H on input $q \in Q$ outputs a hint such that $\Gamma_V(q, \Gamma_H(q)) = 1$.

In the second phase, C_2 takes as input $x := \langle P_n(\pi), C_1(\rho) \rangle_{trans}$ and has oracle access to Γ_V and Γ_H . The execution of C_2 with the input x and the randomness fixed to ρ is denoted by $C_2(x, \rho)$. The queries of C_2 to Γ_V and Γ_H are called *verification queries* and *hint queries* respectively. We say that the circuit C_2 succeeds if and only if it makes a verification query (q, y) such that $\Gamma_V(q, y) = 1$ and it has not previously asked for a hint query on q .

We note that the above definition is very general and does not pose any constraints on size of the circuits or a time complexity of the interactive protocol.

We use the term *weakly verifiable* to emphasize that there is no easy way for the solver to check correctness of a solution except asking a verification query.

We call a weakly verifiable puzzle *dynamic* if the number of hint queries is greater than zero. Furthermore, we say that a weakly verifiable puzzle is *interactive* if in the first phase the number of messages exchanged between the problem poser and the solver is greater than one.

Definition 3.1 generalizes and combines previous approaches that studies *weakly verifiable puzzles* [CHS04], *dynamic weakly verifiable puzzles* [DIJK09], and *interactive weakly verifiable puzzles* [HS10].

There is no loss of generality in assuming that the problem poser and the solver are defined by probabilistic circuits. Definition 3.1 embraces also a case where the problem poser and the solver are probabilistic polynomial time algorithms. We use the well know fact [Hol13b] that a probabilistic polynomial time algorithm can be transformed into an equivalent family of probabilistic Boolean circuits of polynomial size¹.

3.2 Examples

In this section we give examples of cryptographic constructions that motivate studies of different types of weakly verifiable puzzles.

3.2.1 Message Authentication Codes

Let us consider a setting in which two parties a *sender* and a *receiver* communicate over an insecure channel. Messages of the sender may be intercepted, modified, and replaced by a third party called an *adversary*. The receiver needs a way to ensure that received messages have been indeed sent by the sender and have not been modified by the adversary. The solution is to use *message authentication codes*.

Loosely speaking, the message authentication codes may be explained as follows. Let sender, receiver, and adversary be polynomial time algorithms and

¹Theorem 6.10 from [Hol13b] is stated for probabilistic oracle programs with single bit of outputs, but it can be adopted to a case when an output is longer than a single bit.

messages be represented as bitstrings. Furthermore, we assume that the sender and the receiver share a secret key to which the adversary has no access. The sender appends to every message a tag which is computed as a function of the key and the message. The receiver, using the key, has a way to check whether an appended tag is valid for a received message. The receiver accepts a message if the tag is valid, otherwise it rejects. We require that it is hard for the adversary to find a tag and a message, not sent before, that is accepted by the receiver with non-negligible probability. We give the following formal definition of a *message authentication code* based on [Mau13] and [Gol04].

Definition 3.2 (Message Authentication Code) *Let \mathcal{M} be a set of messages, \mathcal{K} a set of keys and \mathcal{T} a set of tags. We define the message authentication code (MAC) as an efficiently computable function $\mathcal{M} \times \mathcal{K} \rightarrow \mathcal{T}$. Furthermore, we say that MAC is secure if it satisfies the following condition:*

Let $k \xleftarrow{\$} \mathcal{K}$ be fixed and H be a polynomial size circuit that takes as input a message $m \in \mathcal{M}$ and outputs a tag $t \in \mathcal{T}$ such that $f(m, k) = t$. We say that MAC is secure if there is no probabilistic polynomial time algorithm with oracle access to H that with non-negligible probability outputs a message $m \in \mathcal{M}$ as well as a corresponding tag $t \in \mathcal{T}$ such that $f(m, k) = t$, and Γ_H has not been queried for a tag of m .

We show how MAC is captured by notion of a dynamic weakly verifiable puzzles where at most one verification query is asked. For fixed f and $n \in \mathbb{N}$ the sender corresponds to the problem poser, the adversary to the problem solver, and the key is a bitstring $\pi \in \{0, 1\}^n$ taken as auxiliary input by the problem poser. In the first phase, which is non-interactive, the problem poser outputs a hint circuit Γ_H and a verification circuit Γ_V where both circuits have hard-coded π . The circuit Γ_H takes as input a message and outputs a tag for this message and corresponds to the circuit H from Definition 3.2. The circuit Γ_V that as input $m \in \mathcal{M}$ and $t \in \mathcal{T}$ and outputs a bit 1 if and only if $f(m, \pi) = t$. In the second phase the problem solver takes no input (x^* is empty string) and is given oracle access to Γ_H and Γ_V . We consider a case where at most one verification query is asked. Thus, a task of finding by an adversary a valid tag $t \in \mathcal{T}$ for a message $m \in \mathcal{M}$ such that a hint for m has not been asked before corresponds to asking a successful verification query by a problem poser to Γ_V .

3.2.2 Public Key Signature Scheme

First, we give a definition of a *public key encryption scheme*, and what it means for such a scheme to be secure. These definitions are based on [Gol04].

Definition 3.3 (Public key signature scheme) *Let \mathcal{Q} be the set of messages. A public key signature scheme is defined by a triple of probabilistic*

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polynomial time algorithms: G – the key generation algorithm, V – the verification algorithm, S – the signing algorithm, such that the following conditions are satisfied:

- $G(1^n)$ outputs a pair of bitstrings $k_{priv} \in \{0,1\}^n$ and $k_{pub} \in \{0,1\}^n$ where $n \in \mathbb{N}$ is a security parameter. We call k_{priv} a private key and k_{pub} a public key.
- The signing algorithm S takes as input $k_{priv} \in \{0,1\}^n$, $q \in \mathcal{Q}$ and outputs a signature $s \in \mathcal{S}$.
- The verification algorithm V takes as input $k_{pub} \in \{0,1\}^n$, $q \in \mathcal{Q}$, and $s \in \mathcal{S}$ and outputs a bit $b \in \{0,1\}$.
- For every k_{priv} , k_{pub} output by G and every $q \in \mathcal{Q}$ it holds

$$\Pr[V(k_{pub}, q, S(k_{priv}, q))] = 1,$$

where the probability is over random coins of V and S .

We say that $s \in \mathcal{S}$ is a valid signature for $q \in \mathcal{Q}$ if and only if $V(k_{pub}, q, s) = 1$.

Definition 3.4 (Security of public key signature scheme with respect to a chosen message attack) Let an adversary A be a probabilistic polynomial time algorithm that takes as input k_{pub} and has oracle access to S . We say that A succeeds if it finds a valid signature $s \in \mathcal{S}$ for a message $q \in \mathcal{Q}$ and the oracle S has not been queried for a signature of q . The public key encryption scheme is secure if there is no polynomial time adversary that succeeds with non-negligible probability.

We will show now that a public key signature scheme defined as above can be represented as a dynamic weakly verifiable puzzle. Let the problem poser correspond to an entity that generates k_{pub} , k_{priv} and the solver to the adversary. In the first phase, the problem poser uses algorithm $G(1^n)$ to obtain k_{pub} , k_{priv} and sends to the adversary the public key k_{pub} . Then the problem poser generates a hint circuit Γ_H and a verification circuit Γ_V . The hint circuit Γ_H takes as input $q \in \mathcal{Q}$ and outputs a signature for q . The verification circuit Γ_V takes as input $s \in \mathcal{S}$ and $q \in \mathcal{Q}$ and checks whether $s \in \mathcal{S}$ is a valid signature for $q \in \mathcal{Q}$. In the second phase, the problem solver takes as input a transcript of the messages from the first round which consists solely of k_{pub} . Additionally, it is given oracle access to Γ_V and Γ_H . It is clear that if the adversary asks a successful verification query (q, s) , then it succeeds also in finding a signature for q .

Thus, finding a valid signature by the adversary of a public key signature schemes is a weakly verifiable puzzle that is dynamic but non-interactive as in the first phase only a single message is sent.

3.2.3 Bit Commitments

Let us consider the following *bit commitment protocol* that involves two parties a *sender* and a *receiver*. We suppose that the sender and the receiver are polynomial time probabilistic algorithms. The protocol consists of a *commit phase* and a *reveal phase*. In the commit phase the sender and the receiver interact, as the result the sender commits to a value $b \in \{0, 1\}$. We require that after the commit phase it is hard for the receiver to correctly guess b . In the reveal phase the sender opens the commitment by sending to the receiver a pair (b', y) where $y \in \{0, 1\}^*$ is an information that helps the receiver to verify that the sender committed to the value $b' \in \{0, 1\}$. A desirable property of a bit commitment protocol is that in the reveal phase it should be hard for the sender to find two bitstrings y_0 and y_1 such that for the receiver both $(0, y_0)$ and $(1, y_1)$ are valid decommitments.

We base the following definition of a *bit commitment protocol* on [Hol13a].

Definition 3.5 (Bit Commitment Protocol) A bit commitment protocol is defined by a pair (S_n, R) where $S_n = (S_1, S_2)$ is a two phase probabilistic circuit, R is a probabilistic circuit, and $n \in \mathbb{N}$ is a security parameter. We call S_n the sender and R the receiver. The circuit S_1 takes as input a pair (b, ρ_S) where $b \in \{0, 1\}$ is interpreted as a bit to which S_n commits, and $\rho_S \in \{0, 1\}^n$ is the randomness used by the algorithm S_n . The receiver R takes as auxiliary input a bitstring $\rho_R \in \{0, 1\}^*$ that is the randomness used by R . The protocol consists of two phases. In the commit phase, circuits S_1 and R engage in the protocol execution. As the result S_1 commits to b and R generates a circuit V . The circuit V takes as input a bit $b' \in \{0, 1\}$ and a bitstring $y \in \{0, 1\}^*$ and outputs a bit. In the reveal phase the circuit S_2 takes as input a communication transcript from the commitment phase $\langle S_1, R \rangle_{\text{trans}}$, the bitstring ρ_S and returns (b', y) . We require a bit commitment protocol to have the following properties:

(Correctness) For a fixed $b \in \{0, 1\}$ we have

$$\Pr_{\substack{\rho_S \in \{0, 1\}^n, \rho_R \in \{0, 1\}^* \\ \Gamma_V := \langle S_1(b, \rho_S), R(\rho_R) \rangle_R \\ (b', y) := S_2(\langle S_1(b, \rho_S), R(\rho_R) \rangle_{\text{trans}}, \rho_S)}} \left[V(b', y) = 1 \right] \geq 1 - \varepsilon(n),$$

where $\varepsilon(n)$ is a negligible function of n .

(Hiding) Probability over random coins of S_n and R_n that any polynomial size circuit can guess bit b correctly after the commit phase is at most $\frac{1}{2} + \varepsilon(n)$ where $\varepsilon(n)$ is a negligible function of n .

(Binding) For every circuit $S_n := (S_1, S_2)$ such that S_1 and S_2 are of size

$\text{poly}(n)$ we have

$$\Pr_{\substack{\rho_S \in \{0,1\}^n, \rho_R \in \{0,1\}^* \\ V := \langle S_1(b, \rho_S), R(\rho_R) \rangle_R \\ ((0, y_0), (1, y_1)) := S_2(\rho_S)}} [V(0, y_0) = 1 \wedge V(1, y_1) = 1] \leq \varepsilon(k),$$

where $\varepsilon(n)$ is a negligible function in n .

Breaking the binding property of a bit commitment protocols can be generalized as an interactive weakly verifiable puzzle. The number of hint queries amounts to zero, and the number of the verification queries is at most one which means that the puzzle is non-dynamic. The receiver is the problem poser. The solver corresponds to the sender trying to break the binding property.

In the first phase the problem solver parses the auxiliary input such that the first bit b is the bit to which the problem solver. The remaining part of the auxiliary input is the randomness used by the problem solver. Then the solver interacts with the problem poser. After, the first phase the problem poser generates a circuit Γ_V that takes as input a y_0, y_1 and outputs 1 if and only if $V(0, y_0) = 1 \wedge V(1, y_1) = 1$. For the bit commitment protocols $|Q| = 1$, thus we do not write on which $q \in Q$ the solver asks a verification query.

In the second phase, the solver is given oracle access to Γ_V and is allowed to ask at most one verification query. We emphasize that the solver has only black box access to Γ_V . Therefore, it has no efficient way to check whether the solution is successful except asking a verification query.

Finally, we notice that for the problem poser and the solver defined as above asking a successful verification query corresponds to breaking the binding property of bit commitment protocol.

3.2.4 Automated Turing Tests

The goal of *Automated Turing Tests* is to distinguish humans from computers. An example of such a test is *CAPTCHA* formally defined in [VABHL03]. Loosely speaking, CAPTCHA is a test for which it is hard to write a computer program that has a success probability comparable or higher to the one achieved by most of humans. An example is an image depicting a distorted text. We note that the definition of hardness is defined by and bases on the opinions of AI community [VABHL03].

CAPTCHAs based on guessing the distorted text can be modeled as weakly verifiable puzzles. In the first round the problem poser and the solver engage in the execution of an interactive protocol such that after its execution the problem poser sends to the solver an image containing the distorted text. Furthermore, the problem poser generates a circuit Γ_V that takes as input a

bitstring y and outputs 1 if and only if y correctly encodes the text depicted on the distorted image.

The problem poser in the second round takes as input a distorted image, has oracle access to Γ_V and can ask at most a single verification query.

Standard CAPTCHAs are non-dynamic, as the problem solver does not gain access to the hint oracle.

As it is not known how good the possible algorithm can be to recognize CAPTCHA it is likely that the gap between human performance and a performance of computer programs may be small. Therefore, it is of interest to find a way to amplify this gap. Actually, in [HS10] it has been shown that it is possible for parallel repetition. The solver is given n independent puzzles. The verifier accepts when the solver succeeds on at least δn fraction of puzzles. It turns out that we can create weakly verifiable puzzles for which success probability of a computer program is at most ε and which is substantially easier for humans that succeeds with at least probability $1 - \varepsilon$ where $\varepsilon \in (0, 1)$.

In Chapter 4 we give a more general proof than the one in [HS10] that applies also to dynamic context.

3.3 Previous results

Different types of weakly verifiable cryptographic primitives have been studied in a series of works [CHS04, DIJK09, HS10]. This section is intended to give a short overview of techniques used in these works and aims to provide some intuition and insight into the problem of hardness amplification of dynamic interactive weakly verifiable puzzles.

3.3.1 Results of R.Canetti, S.Halevi, and M.Steiner

The notion of the *weakly verifiable puzzle* has been coined by R.Canetti, S.Halevi, and M.Steiner in the paper *Hardness amplification of weakly verifiable puzzles* [CHS04]. In comparison to Definition 3.1 the puzzles considered in [CHS04] are neither dynamic nor interactive. Moreover, the number of verification queries is limited to one. This constitutes a simplified case to the one considered in this Thesis. In this section we provide the definition of weakly verifiable puzzles (WVP) that closely follows the one contained in [CHS04] and state the theorem of hardness amplification of weakly verifiable puzzles in a similar vein as in [CHS04]. Finally, we give an intuition behind the proof of this theorem. It is noteworthy that the main proof of this Thesis, contained in Chapter 4, uses many ideas from the work of R.Canetti, S.Halevi, and M.Steiner [CHS04].

Definition 3.6 (Weakly Verifiable Puzzles) *A weakly verifiable puzzle is defined by a pair of polynomial time algorithms: a probabilistic puzzle-generation*

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algorithm G and a deterministic verification algorithm V . We write $G(1^k; \rho)$ to denote that G takes as input a bitstring 1^k , where k is a security parameter, and as auxiliary input a bitstring $\rho \in \{0, 1\}^*$ which is the randomness used by G . The algorithm G outputs a bitstring $p \in \{0, 1\}^*$ and a check information $c \in \{0, 1\}^*$. The verifier V is a deterministic algorithm that takes as input p , c , an answer $a \in \{0, 1\}^*$ and outputs $b \in \{0, 1\}$.

A solver S for G is a polynomial time probabilistic algorithm that takes as input p and outputs a . We denote the randomness used by S as $\pi \in \{0, 1\}^*$ and define the success probability of S in solving a puzzle defined by P as

$$\Pr_{\substack{\rho \in \{0, 1\}^*, \pi \in \{0, 1\}^* \\ (p, c) := G(1^k; \rho) \\ a := S(p, \pi)}} [V(p, c, a) = 1].$$

We write $P := (G, V)$ to denote a weakly verifiable puzzle P defined by algorithms G and V .

We compare the above definition with Definition 3.1. First, we note that in Definition 3.1 we use the language of probabilistic circuits. This is more general approach as probabilistic polynomial time algorithms can be converted into families of polynomial size circuits. Next, we see that in Definition 3.6 the algorithm G is parameterized by a bitstring 1^k meaning that the length of a random bitstring taken by G is bounded by $\text{poly}(k)$. For a fixed k , without loss of generality, we can model the algorithm $G(1^k; \rho)$ as a polynomial size probabilistic circuit that does not take as input 1^k , but just a bitstring ρ of length $\text{poly}(k)$. The security parameters from Definition 3.6 and Definition 3.1 are not equivalent, as in the later definition the security parameter limits the length of the random bistring. Moreover, in Definition 3.6 a verification algorithm takes as input p , c , a . Again, without loss of generality, we can assume that bitstrings p and c are hard-coded in the circuit Γ_V from Definition 3.1. Hence, the algorithm V corresponds to Γ_V . The puzzles considered in Definition 3.6 are non-dynamic. Thus, there is not corresponding element for a hint circuit Γ_H from Definition 3.1. Finally, the puzzles described in Definition 3.6 are non-interactive.

We will formulate now the definition of n -fold repetition of Weakly Verifiable Puzzles along the lines of [CHS04]. Later in this Thesis, for other types of weakly verifiable puzzles, we use the notion of k -wise direct product of puzzles instead.²

Definition 3.7 (n -fold repetition of Weakly Verifiable Puzzles) *Let $n \in \mathbb{N}$ and a weakly verifiable puzzle $P = (G, V)$ be fixed. We define the n -fold repetition*

²We note that the terminology how to call the parallel repetition of n weakly verifiable puzzles is different in different works. In [CHS04] the term *n -fold direct product of puzzles* is used whereas in [DIJK09] the similar notion, but for dynamic puzzles, is defined as *k -wise repetition of puzzles*. In this Thesis we tend to use the latter formulation.

tition of P as a weakly verifiable puzzle where the puzzle-generation algorithm $G^{(n)}$ takes as input 1^k , as an auxiliary input a bitstring $\rho \in \{0, 1\}^*$ and outputs tuples $p^{(n)} := (p_1, \dots, p_n) \in \{0, 1\}^*$ and $c^{(n)} := (c_1, \dots, c_k) \in \{0, 1\}^*$ where for each $1 \leq i \leq n$ a pair (p_i, c_i) is an independent instance of a weakly verifiable puzzle defined by G and V with security parameter k . Finally, the verification algorithm $V^{(n)}$ takes as input $p^{(n)}$, $c^{(n)}$, an answer $a^{(n)}$, and outputs $b \in \{0, 1\}$ such that $b = 1$ if and only if for all $1 \leq i \leq n$ we have $V(p_i, c_i, a_i) = 1$. We write $P^{(n)} := (G^{(n)}, V^{(n)})$ to denote the n -fold repetition of P .

Let us lay down some terminology. For $P^{(n)} := (G^{(n)}, V^{(n)})$ we define a *puzzle on i -th coordinate* as i -th puzzle of the n -fold repetition (the puzzle that corresponds to $G_i^{(n)}, V_i^{(n)}$). The n -fold repetition of weakly verifiable puzzles is solved successfully if and only if all n puzzles are solved successfully. In contrast, in Chapter 4 we are interested in a more general situation where a monotone function $g : \{0, 1\}^n \rightarrow \{0, 1\}$ is used to decide which coordinates of the n -fold repetition of puzzles have to be solved correctly. A precise definition is given in Section 4.1. Clearly, we can assume that g is such that all coordinates have to be solved successfully which matches the case considered in Definition 3.7.

The main theorem proved in [CHS04] states that it is possible to turn a good solver for $P^{(n)}$ to a good solver for P .

Theorem 3.8 (Hardness amplification of Weakly Verifiable Puzzles)

Let $n : \mathbb{N} \rightarrow \mathbb{N}$ and $\delta : \mathbb{N} \rightarrow (0, 1)$ be efficiently computable functions and $q \in \mathbb{N}$ a slackness parameter. Moreover, let $P = (G, V)$ be a weakly verifiable puzzle. We denote the running time of the puzzle-generation algorithm G for P by T_G and of the verification algorithm V for P by T_V . If $S^{(n)}$ is a solver for the n -fold repetition of P that success probability is at least δ^n and running time is T , then there exists a solver S for G with oracle access to $S^{(n)}$ that success probability is at least $\delta(1 - \frac{1}{q})$ and running time $O\left(\frac{nq^3}{\delta^{2n}-1}(T + nT_G + nT_V)\right)$.

The following algorithm is used by R.Cannetti, S.Halevi, and M.Steiner in the proof of Theorem 3.8. It transforms $S^{(n)}$ for $P^{(n)}$ with success probability at least δ^n to solver for a single puzzle P with probability at least $\delta(1 - \frac{1}{q})$. The slackness parameter q has to be used as it is not possible to achieve the perfect hardness amplification. We note that in the analysis of running time of *CHS-solver* we explicitly take into account time needed for oracle calls to $S^{(n)}, V, G$.

We denote by $p \in \{0, 1\}^*$ a bitstring output by G and taken as input by *CHS-solver*. To make notation shorter in the following code excerpts we do not write randomness used by G explicitly.

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Algorithm: $CHS\text{-}solver^{S^{(n)},V,G}(p, n, k, q, \delta)$

Oracle: A solver $S^{(n)}$ for $P^{(n)}$, a verification algorithm V for P , a puzzle-generation algorithm G for P .

Input: A bistring $p \in \{0, 1\}^*$, parameters n, k, q, δ .

$prefix := \emptyset$

for $i = 1$ **to** $n - 1$ **do:**

$p^* := \text{ExtendPrefix}^{S^{(n)},V,G}(prefix, i, n, k, q, \delta)$

if $p^* = \perp$ **then return** $\text{OnlinePhase}^{S^{(n)},V,G}(prefix, p, i, n, k, q, \delta)$

else $prefix := prefix \circ p^*$

$a^{(n)} := S^{(n)}(prefix \circ p)$

return a_n

Algorithm: $\text{OnlinePhase}^{S^{(n)},V,G}(prefix, p, v, n, k, q, \delta)$

Oracle: A solver algorithm $S^{(n)}$ for $P^{(n)}$, a puzzle-generation algorithm G for P , a verification algorithm V for P .

Input: A $(v - 1)$ -tuple of bitstrings $prefix$, a bitstring $p \in \{0, 1\}^*$, parameters v, n, k, q, δ .

Repeat $\left\lceil \frac{6q \ln(6q)}{\delta^{n-v+1}} \right\rceil$ **times**

$((p_{v+1}, \dots, p_n), (c_{v+1}, \dots, c_n)) := G^{(n-v-1)}(1^k)$

$a^{(n)} := S^{(n)}(prefix, p, p_{v+1}, \dots, p_n)$

if $\forall_{v+1 \leq i \leq n} V(p_i, c_i, a_i) = 1$ **then return** a_v

return \perp

Algorithm: $\text{ExtendPrefix}^{S^{(n)},V,G}(prefix, i, n, k, q, \delta)$

Oracle: A solver algorithm $S^{(n)}$ for $P^{(n)}$, a puzzle-generation algorithm G for P , a verification algorithm V for P .

Input: A $(i - 1)$ -tuple of puzzles $prefix$, parameters i, n, k, q, δ .

Repeat $\left\lceil \frac{6q}{\delta^{n-v+1}} \ln\left(\frac{18qn}{\delta}\right) \right\rceil$ **times do:**

$(p^*, c^*) := G(1^k)$

$\bar{v}_i := \text{EstimateResSuccProb}^{G,V}(prefix \circ p^*, i, n, k, q, \delta)$

if $\bar{v}_i \geq \delta^{n-i}$ **then return** p^*

return \perp

Algorithm: $EstimateResSuccProb^{S^{(n)}, V, G}(prefix, i, n, k, q, \delta)$

Oracle: A solver algorithm for $P^{(n)}$, a verification algorithm V for P , a generation algorithm G for P

Input: A i -tuple of puzzles $prefix$, parameters i, n, k, q, δ .

$successes := 0$

Repeat $M := \left\lceil \frac{84q^2}{\delta^{n-i}} \ln \left(\frac{18qn \cdot N_i}{\delta} \right) \right\rceil$ times

$((p_{i+1}, \dots, p_n), (c_{i+1}, \dots, c_n)) := G^{(n-i)}(1^k)$

$a^{(n)} := A(prefix, p_{i+1}, \dots, p_n)$

if $\forall_{i+1 \leq j \leq n} : V(p_j, c_j, a_j) = 1$ **then** $successes := successes + 1$

return $successes/M$

A detail proof of Theorem 3.8 is presented in [CHS04]. We limit ourselves to providing an intuition why the above algorithm transforms a good solver for the n -wise direct product of P to a good solver for P .

Let us consider the n -fold repetition of P , and for simplicity a deterministic solver $S^{(n)}$ for $P^{(n)} := (G^{(n)}, V^{(n)})$. Furthermore, we write $p^{(n)}, c^{(n)}$ to denote the output of $G^{(n)}$. We define a matrix M as follows. The columns of M are labeled with all possible bitstrings p_1 whereas the rows are labeled with all possible tuples (p_2, \dots, p_n) output by algorithm $G^{(n)}$ executed with different randomness. A cell of M contains a binary n -tuple such that the i -th bit equals 1 if and only if $V_i(p_i, c_i, a_i) = 1$ where $a^{(n)} := S^{(n)}(p^{(n)})$ and $p^{(n)}$ is a tuple of bitstring inferred by a column and a row of the cell. We make the following observation.

Observation 3.9 *For a deterministic polynomial time algorithm $S^{(n)}$ that successfully solves the n -fold repetition of P with probability at least δ^n the matrix M defined as above has either a column with $\delta^{(n-1)}$ fractions of cells that are all one vectors, or a conditional probability that a cell is of the form 1^n given that the last $(n-1)$ bits of the cell are equal 1 is at least δ .*

We show, at least intuitively using Observation 3.9, how the algorithm *CHS-solver* can be used to solve P with substantial probability given oracle access to $S^{(n)}$ for $P^{(n)}$. The algorithm starts with the first position and tries to fix a bitstring p^* on this position such that the success probability of $S^{(n)}$ on the remaining $(n-1)$ position is at least $\delta^{(n-1)}$. If it is possible to find p^* such that this condition is satisfied, then we fix p^* on this position and repeat the whole procedure again in the consecutive iteration for the next position. If *CHS-solver* fails to find a bitstring p^* , then we assume that there is no column of M that contains $\delta^{(n-1)}$ fraction of cells that are of the form 1^n . We use Observation 3.9 and hope that the conditional probability of solving the first puzzle given that all puzzles on the remaining position are solved successfully is at least

δ . We place the input puzzle p on this position and note that all remaining puzzles are generated by *CHS-solver*. Thus, it is possible to efficiently verify whether these puzzles are successfully solved by $S^{(n)}$.

Obviously, the algorithm *CHS-solver* can still fail. First, it may happen that it does not find a column with a high fraction of puzzles that are solved successfully, although such a column exists. Secondly, we cannot exclude a situation where no such column exists, but the algorithm fails to find a cell such that last $(n-1)$ bits are 1. Finally, it is also possible that an estimate returned by *EstimateResSuccProb* is incorrect.

It is possible to show that all these events happen with negligible probability. Therefore, at least intuitively we see that the algorithm *CHS-solver* solves a single WVP puzzle successfully with probability at least $\delta(1-\frac{1}{q})$ almost surely.

In Chapter 4 we study a more general class of puzzles that are not only weakly verifiable but also dynamic and interactive. Furthermore, we allow a more general situation where a solver successfully solves the n -fold repetition of puzzles³ although it is successful only on some coordinates of the n -fold repetition of P . It turns out that it is possible to use a similar technique of fixing puzzles on consecutive positions of the n -fold repetition of puzzles to prove the hardness amplification in this more general setting.

3.3.2 Results of Y.Dodis, R.Impagliazzo, R.Jaiswal, V.Kabanets

Some of the cryptographic constructions presented in Section 3.2 are not only weakly verifiable but also dynamic (MAC and SIG). This type of puzzles are defined and studied in [DIJK09]. We give a short overview of this work, state the definition of a *dynamic weakly verifiable puzzle* that closely follows the one included in [DIJK09]. Finally, we provide intuition for the proof of the hardness amplification of DWVP included in [DIJK09].

Definition 3.10 (Dynamic Weakly Verifiable Puzzle.) A dynamic weakly verifiable puzzle (DWVP) is defined by a distribution \mathcal{D} on pairs (x, α) where $\alpha \in \{0, 1\}^*$ is an advice used to generate and evaluate responses and $x \in \{0, 1\}^*$ is a bitstring taken as input by the solver. Furthermore, we consider a set \mathcal{Q} and a probabilistic polynomial time computable relation R such that $R(\alpha, q, r) = 1$ if and only if r is a correct answer to $q \in \mathcal{Q}$ on the set of puzzle determined by α . Finally, let $H(\alpha, q)$ be a probabilistic polynomial time computable hint relation.

A solver S takes as input x and can ask hint queries on $q \in \mathcal{Q}$ which are answered using $H(\alpha, q)$ and verification queries of the form (q, r) answered by means of $R(\alpha, q, r)$. We say that S succeeds if and only if it makes a

³Actually, in Chapter 4 we define the k -wise repetition of puzzles following the terminology used in [DIJK09] which is equivalent to the n -fold repetition of puzzles.

verification query on (q, r) such that $R(\alpha, q, r) = 1$ and it has not previously asked for a hint query on this q . We write $P := (\mathcal{D}, R, H)$ to denote a DWVP with a distribution \mathcal{D} of pairs (x, α) , and R, H being a verification and hint relations respectively.

We show now how the above definition is generalized by Definition 3.1. First, instead of considering a distribution on pairs (x, α) in Definition 3.1 we use a probabilistic problem poser that outputs circuits Γ_H and Γ_V that corresponds to hint and verification relations respectively. Furthermore, the problem poser may interact in the first phase with the problem solver. In particular, the problem poser can send a bitstring x as a message in the first phase. Thus, Definition 3.1 captures a more general case where the distribution of puzzles is defined by both the problem poser and the problem solver.

We define the n -wise direct product of DWVPs, which is conceptually similar to the n -fold repetition of WVPs.

Definition 3.11 (n -wise direct product of DWVPs) For a dynamic weakly verifiable puzzle $P := (\mathcal{D}, R, H)$ we define the n -wise direct product of P as a DWVP with a distribution $\mathcal{D}^{(n)}$ on tuples $(x_1, \alpha_1), \dots, (x_n, \alpha_n)$. Furthermore, the hint relation is defined by $H^{(n)}(q, \alpha_1, \dots, \alpha_n) := (H(\alpha_1, q), \dots, H(\alpha_n, q))$ and the verification relation $R^{(n)}(\alpha_1, \dots, \alpha_n, r_1, \dots, r_n, q)$ evaluates to 1 if and only if for $1 \leq i \leq n$ at least $n - (1 - \gamma)\delta n$ is such that $R(\alpha_i, q, r_i) = 1$ where $0 \leq \gamma, \delta \leq 1$. We write $P^{(n)} := (D^{(n)}, H^{(n)}, R^{(n)})$ to denote the n -wise direct product of $P := (D, H, R)$.

In contrast to the n -fold repetition of puzzles defined in previous section, here we require a solver to succeed only on a fraction of puzzles.

Dynamic weakly verifiable puzzles generalize breaking message authentication codes and public signature schemes. In case of MAC the adversary takes x which is an empty string. For the public signature schemes x is the public key.

We write $\mathcal{H}_{\text{hint}}, \mathcal{V}_{\text{verif}} \leftarrow S(x; \delta)$ to denote the execution of the solver S with input $x \in \{0, 1\}^*$ and using randomness $\delta \in \{0, 1\}^*$ where $\mathcal{H}_{\text{hint}}$ is the set of all hint queries asked by S and by $\mathcal{V}_{\text{verif}}$ is a set of all pairs (q, a) of verification queries asked in the execution of S .

With no loss of generality we make an assumption that the solver does not make hint queries on the successful verification queries. We define the *success probability* of a solver S for $P := (G, V)$ as

$$\Pr_{\substack{\delta \in \{0, 1\}^* \\ (x, \alpha) \leftarrow \mathcal{D} \\ (\mathcal{H}_{\text{hint}}, \mathcal{V}_{\text{verif}}) \leftarrow S(x, \delta)}} [\exists (q, a) \in \mathcal{V}_{\text{verif}} : q \notin \mathcal{H}_{\text{hint}} \wedge V(q, a) := 1]$$

Theorem 3.12 (Hardness amplification for dynamic weakly verifiable puzzle). *Let $S^{(n)}$ be a probabilistic algorithm for the $P^{(n)}$ that succeeds with probability at least ε , where $\varepsilon \geq (800/\gamma\delta) \cdot (h + v) \cdot e^{-\gamma^2\delta n/40}$, and h and v denote the number of hint and verification queries asked by $S^{(n)}$ respectively. Then there exists a probabilistic algorithm S that succeeds in solving P with probability at least $1 - \delta$ making $O(h(h + v)/\varepsilon) \cdot \log(1/\gamma\delta)$ hint queries and at most one verification query. Furthermore, the running time $\text{poly}(h, v, \frac{1}{\varepsilon}, t, \omega, \log(1/\gamma\delta))$ where ω is time needed to ask a single hint query.*

It is worth seeing why the approach presented in the previous section that works well for n -fold repetition of WVP can not be applied in the setting of the k -wise direct product of DWVP (moving aside for a moment the issue of solving only a fraction of puzzle successfully). For DWVP the algorithm *CHS-solver* breaks in the *OnlinePhase* where the solver $S^{(n)}$ can be called multiple times. It is possible that in one of these runs $S^{(n)}$ asks a hint query on q for which in one of the later runs a verification query (q, r) is asked for which algorithm would return an answer for the input puzzle (in other words the condition $\forall_{v+1 \leq i \leq n} V(p_i, c_i, a_i) = 1$ is satisfied). However, the fact that a hint query on this q has been asked makes it impossible to ask a successful verification query on this q . Thus, we can not dismiss a situation where the success probability of $S^{(n)}$ decreases with the number of iterations.

The solution proposed in [DIJK09] is to partition the set \mathcal{Q} into a set of *attacking queries* $\mathcal{Q}_{\text{attack}}$ and a set of *advice queries* \mathcal{Q}_{adv} . The idea is to allow a solver for the n -wise direct product to ask hint queries only on $q \in \mathcal{Q}_{\text{adv}}$, and to halt the execution whenever a hint query is asked on $q \in \mathcal{Q}_{\text{attack}}$.

It is possible, for a solver S that asks at most h hint queries and v verification queries, to find a function $\mathcal{Q} \rightarrow \{0, 1, \dots, 2(h + v)\}$ such that the success probability of S with respect to $\mathcal{Q}_{\text{attack}}$ and \mathcal{Q}_{adv} is multiplied by $\frac{1}{8(h+v)}$. If h and v are not too big, then the success probability of S can still be substantial. More formally, for a function $\text{hash} : \mathcal{Q} \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$ we define $\mathcal{Q}_{\text{attack}} := \{q \in \mathcal{Q} : \text{hash}(q) = 0\}$ and $\mathcal{Q}_{\text{adv}} := \{q \in \mathcal{Q} : \text{hash}(q) \neq 0\}$

In [DIJK09] the following lemma is proved.

Lemma 3.13 *Let S be a solver for DWVP which success probability is at least ε , the running time is at most t , and the number of hint and verification queries is at most h and v respectively. There exists a probabilistic algorithm that runs in time $\text{poly}(h, v, \frac{1}{\varepsilon}, t)$ that outputs a function $\text{hash} : \mathcal{Q} \rightarrow \{0, 1, \dots, 2(h + v)\}$ that partitions \mathcal{Q} to $\mathcal{Q}_{\text{attack}}$ and \mathcal{Q}_{adv} such that with probability at least $\frac{\varepsilon}{8(h+v)}$ the first successful verification query (q', a) asked by S is such that $q' \in \mathcal{Q}_{\text{attack}}$ and all previous hint and verification queries has been asked on $q \in \mathcal{Q}_{\text{adv}}$.*

A function *hash* can be found by using a natural sampling technique. We follow exactly the same approach of the partitioning the domain in the main proof of this Thesis in Section 4.1.3.

Let $H_\alpha(q)$ denote a polynomial time probabilistic algorithm that takes as input q , has hard-coded α and outputs $H(\alpha, q)$. Similarly, we use $R_\alpha(q, r)$ to denote a polynomial time probabilistic algorithm that computes relations $R(\alpha, q, r)$ and has hard-coded bitstring α . The following algorithm is used in [DIJK09] in the proof of lemma 3.12. It gains oracle access to R_α and H_α as well as a function *hash* from Lemma 3.13.

Algorithm: $DWVP_solver^{S^{(n)}, hash, H_\alpha^{(n)}, R_\alpha^{(n)}}(x)$

Oracle: A solver $S^{(n)}$ for $P^{(n)}$, a function $hash : Q \rightarrow \{0, 1, \dots, 2(h+v)\}$.

Input: A bistring $x \in \{0, 1\}^*$.

Repeat at most $O(\frac{h+v}{\epsilon} \cdot \log(\frac{1}{\gamma\delta}))$ times

Let $i \xleftarrow{\$} \{1, \dots, n\}$ be a position for x .

Generate $(x_1, \alpha_1), \dots, (x_{i-1}, \alpha_{i-1}), (x_{i+1}, \alpha_{i+1}), \dots, (x_n, \alpha_n)$ using $(n-1)$ calls to P each time with fresh randomness.

run $S^{(n)}(x_1, \dots, x_{i-1}, x, x_{i+1}, \dots, x_n)$

if $S^{(n)}$ asks a hint query on q **then**

if $hash(q) \neq 0$ **then** abort current run of $S^{(n)}$

 Ask a verification query $r := H(q)$

 Let $(r_1, \dots, r_{i-1}, r_{i+1}, \dots, r_n)$ be hints for query q for puzzle sets $(x_1, \dots, x_{i-1}, x_{i+1}, x_n)$

 Answer the hint query of $S^{(n)}$ using $(r_1, \dots, r_{i-1}, r, r_{i+1}, r_n)$

if $S^{(n)}$ asks a verification query (q, r_1, \dots, r_n) **then**

if $hash(q) = 0$ **then** answer the query with 0

 Let $m := |j : V(q, r_j) = 1, j \neq i|$

if $m \geq n - n(1 - \gamma)\delta$ **then**

 make a verification query (q, r_i) and halt.

else with probability $\rho^{m-n(1-\gamma)\delta}$ ask a verification query (q, r_i) and halt.

 Halt the current run of $S^{(n)}$ and go to the next iteration.

return \perp

In the above algorithm we execute multiple times a solver $S^{(n)}$ for the k -wise direct product of DWVPs. In each iteration the position for $x \in \{0, 1\}^*$ is chosen uniformly at random. The remaining $(n-1)$ puzzles are generated by the algorithm, thus it is possible to answer all hint and verification queries for these puzzles. We use a function *hash* to partition the query domain. We assume that *hash* is such that the success probability of $S^{(n)}$ with respect to *hash* is at least $\frac{\delta}{8(h+v)}$. We check on which q the solver $S^{(n)}$ asks hint and verification queries. If a hint query is asked on q such that $hash(q) = 0$ then the execution of $S^{(n)}$ is aborted and we go to the next iteration. This way

me make sure that the algorithm never asks a hint query that could prevent a verification query to succeed.

If a verification query is asked on q such that $\text{hash}(q) \neq 0$ we answer such a verification query with 0.

Finally, in case when $S^{(n)}$ asks a verification query using index q such that $\text{hash}(q) = 0$, then we use a soft decision system to decide whether to ask a verification query. The idea is that if there are many puzzles among the ones generated by the algorithm is solved correctly then it is likely that also the input puzzle is solved successfully. We also have to discount $\gamma\delta n$ to take into account that we require not all puzzles to be solved successfully. The detail calculations provided in [DIJK09] shows that this approach yields a demanded result. We do not give a more detail description of the proof of [DIJK09] as in Chapter 4, except the domain partitioning, we use a different prove technique.

In case of weakly verifiable primitives like CAPTCHAs we assume that most people have at least slightly higher probability of solving these kind of puzzles than the best computer programs. Still, it may happen that humans do not solve all puzzles. This is a motivation to introduce a threshold function such that on average solutions of humans are treated as solved successfully but the ones of computer programs on average are classified as not successful. This is a motivation to study the situations where only a fraction of puzzles is solved successfully.

In Chapter 4 we consider a weakly verifiable puzzles that are interactive and dynamic. We use a very similar technique to partition domain Q into advice and hint queries as presented in [DIJK09]. Instead of the requirement to succeed only on the fraction of puzzles we consider an arbitrary, monotone function $g : \{0, 1\}^n \rightarrow \{0, 1\}$ that determine on which coordinates the solver has to succeed in order to successfully solve the n -wise direct product of puzzles.

To show the hardness amplification for the n -wise direct product with the domain partitioned we use the approach similar to the one presented in Section 3.3.1. Namely, we try to find a good position for the input puzzle instead of choosing the position uniformly on random as in [DIJK09].

3.3.3 Results of T.Holenstein and G.Scheonebeck

TODO: Justify why you do not give a more detail description of the algorithm as in the previous sections .

TODO: Explain in detail what is in your work and what is here. What are my contributions.

The hardness amplification of interactive weakly verifiable puzzles has been studied by T.Holenstein and G.Schoenebeck in [HS10]. We will give now an

overview of this work and compare it with our approach.

The following definition of an *interactive weakly verifiable puzzle* closely follow the one from [HS10].

Definition 3.14 *An interactive weakly verifiable puzzle is defined by a protocol given by two probabilistic algorithms P and S . The algorithm P is called the problem poser and produces as output a verification circuit Γ . The algorithm S , called the problem solver, produces no output. Furthermore, the success probability of the algorithm S in solving an interactive weakly verifiable puzzle defined by (P, S) is:*

$$\Pr_{\rho, \pi} \left[\Gamma^{(g)}(\langle P(\rho), S(\pi) \rangle_S) = 1 \right].$$

$\Gamma^{(g)} := \langle P(\rho), S(\pi) \rangle_P$

We are interested in the hardness amplification of interactive weakly verifiable puzzles. Thus, similarly as in the previous sections we define the k -wise direct product of puzzles.

Definition 3.15 (k -wise direct product of interactive weakly verifiable puzzles) *Let $g : \{0, 1\}^k \rightarrow \{0, 1\}$ be a monotone function and (P, S) be a fixed interactive weakly verifiable puzzle. The k -wise direct product of (P, S) defined by $(P^{(g)}, S^{(g)})$ is an interactive weakly verifiable puzzle in which the sender and the receiver sequentially interact in k rounds where in each round (P, S) is used to generate an instance of interactive weakly verifiable puzzle. As the result circuits $\Gamma^{(1)}, \dots, \Gamma^{(k)}$ for P are generated. Finally, $P^{(g)}$ outputs the circuit $\Gamma^{(g)}(y_1, \dots, y_k) := g(\Gamma^{(1)}(y_1), \dots, \Gamma^{(k)}(y_k))$.*

Similarly as in Definition 3.1 the puzzles considered in [HS10] have an interactive phase. Furthermore, a monotone binary function is used to determine whether the k -wise repetition has been successfully solved. Unlike, puzzles in Definition 3.1, the puzzles studied by T.Holenstein and G.Schoenebeck are non-dynamic. Thus, only a verification circuit Γ is generated and no hint circuit is ever used.

The following hardness amplification theorem is proved in [HS10].

Theorem 3.16 *There exists an algorithm $\text{Gen}(C, g, \varepsilon, \delta, n)$ which takes as input a solver circuit C for the k -wise direct product of P , a monotone function $g : \{0, 1\}^k \rightarrow \{0, 1\}$, and parameters ε, δ, n . The algorithm Gen outputs a solver circuit D for P such that the following holds. If C is such that*

$$\Pr \left[\Gamma^{(g)}(\langle P^{(g)}, C \rangle_C) = 1 \right] \geq \Pr_{u \leftarrow \mu_\delta^{(k)}} \left[g(u) = 1 \right] + \varepsilon,$$

then, D satisfies almost surely,

$$\Pr \left[\Gamma(\langle P, D \rangle_D) = 1 \right] \geq \delta + \frac{\varepsilon}{6k}.$$

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Additionally, Gen and D only require oracle access to g and C . Furthermore, $Size(D) \leq Size(C) \cdot \frac{6k}{\varepsilon} \log(\frac{6k}{\varepsilon})$, and $Time(Gen) = poly(k, \frac{1}{\varepsilon}, n)$ with oracle calls to C .

First, we notice that the above definition does not impose any restrictions on the time complexity of the poser and the solver. We consider a general approach where Gen is used to define a polynomial time reduction between a solver for the k -wise direct product of puzzles to a solver for a single puzzle. Furthermore, in the previous sections we considered solvers for the k -wise direct product that were compared with the algorithms that either solves all puzzles ([CHS04]) or allowed a fraction of puzzles to be solved incorrectly ([DIJK09]). In the above definition a more general case is considered where we use a binary monotone function g . More precisely we are interested functions that are binary monotonously non-decreasing.

We emphasize that the monotone restriction on g is essential. For $g(b) := 1 - b$ an algorithm that deliberately gives incorrect answers satisfies g with probability 1 whereas an algorithm that solves a puzzle successfully with probability $\gamma > 0$ succeeds only with probability $1 - \gamma$. A desirable property for the solver for the k -wise direct product of IWVP is that an algorithm that solves puzzles on all coordinates with the higher probability does not the k -wise direct product with probability lower than an algorithm that success probability on these coordinates is lower.

The proof technique used by T.Holenstein and G.Schoenebeck is similar to the one presented in Section 3.3.1. In chapter 4 we use very similar approach and fix puzzles on consecutive coordinates of the n -wise direct product.

TODO: Compare to the work of CHS i.e. what we use there to compare the puzzles

In order to estimate how much better a solver circuit C for the n -wise direct product performs when a puzzle on the first position is fixed a notion of a surplus $S_{\pi^*, b}$ is introduced:

$$S_{\pi^*, b} := \Pr_{\pi^{(k)}} [c \in \mathcal{G}_b | \pi_1 = \pi^*] - \Pr_{u \leftarrow \mu_\delta^k} [u \in \mathcal{G}_b],$$

which intuitively tells us how much better a solver C performs when a first puzzle is always solved correctly (case when $b = 1$) or is always solved incorrectly (when $b = 0$). Now we observe the following fact. If there exists a puzzle which is fixed on the first position for which the surplus is bigger than $(1 - \frac{1}{k})\varepsilon$ then we can fix a this first puzzle and inductively solve the problem for the $(k - 1)$ -direct product of puzzles.

TODO: what do it mean

On the other hand, if we there is no such puzzle to fix on the first position it means that when we fix the first bit of g then the performance between the solver C and an algorithm that solves puzzle on each position independently with probability δ is similar. However, we know that when the first bit of a function g is not fixed then the solver C is better. Thus, we draw a conclusion that the puzzle on the first position has to be solved unusually often.

In a case it would be possible to fix all $(k - 1)$ puzzles except the last one, the proof become trivial as we know that a function g with the first $k - 1$ bits fixed is either the identity or a constant function.

Thus, it is enough to show that if it is not possible to find an estimate that is low then if we place an input puzzle on this position and we can find remaining $k - 1$ puzzle such that $c \in \mathcal{G}_1 \setminus \mathcal{G}_0$ then this puzzle is solved with substantial probability. The whole proof is given in [HS10], and requires some probability manipulations.

Our proof of the hardness amplification for dynamic interactive weakly verifiable puzzles closely follows the one given in [HS10].

TODO: Why do we consider fixing 0/1 on the first position **TODO:** How the technique is generalized to approach of CHS

TODO: Explain why we compare to such a probability i.e. why we consider with μ

TODO: Why the requirement for g being a monotone function is interesting

TODO: Give the intuition behind the proof.

TODO: Give a proof under the simplified assumptions?

1) The algorithm always output an answer

2) For every π the surpluses $S_{\pi^*,0}$ and $S_{\pi^*,1}$ are less than $(1 - \frac{1}{k})\varepsilon$.

3.4 Limitations of Security Amplification

Chapter 4

Security amplification for dynamic weakly verifiable puzzles

In this chapter we show that it is possible to amplify security of dynamic weakly verifiable puzzles. In section 4.1 we state the theorem, which is next proved in three steps. First, in Section 4.1.3, we show how to use to partition the domain on which hint and verification queries are asked, next we give a prove of security amplification under the simplifying assumption that there is no collisions of hint and verification queries. Finally, in Section 4.1.5 we combine both former steps which yields the desirable result.

4.1 Main theorem

We start by giving the definition of the k -wise direct product of weakly verifiable puzzles.

4.1.1 The k -wise direct product of weakly verifiable puzzle

Definition 4.1 (k -wise direct-product of DWVPs.) Let $g : \{0, 1\}^k \rightarrow \{0, 1\}$ be a monotone function and $P_n^{(1)}$ a problem poser as in Definition 3.1. The k -wise direct product of $P_n^{(1)}$ is a DWVP defined by a circuit $P_{kn}^{(g)}$. We write $P_{kn}^{(g)}(\pi^{(k)})$ to denote the execution of $P_{kn}^{(g)}$ with the randomness fixed to $\pi^{(k)} := (\pi_1, \dots, \pi_k)$ where for each $1 \leq i \leq k : \pi_i \in \{0, 1\}^n$. Let $(C_1, C_2)(\rho)$ be a solver for $P_{kn}^{(g)}$ as in Definition 3.1. In the first phase, the algorithm $C_1(\rho)$ sequentially interacts in k rounds with $P_{kn}^{(g)}(\pi^{(k)})$. In the i -th round $C_1(\rho)$ interacts with $P_n^{(1)}(\pi_i)$, and as the result $P_n^{(1)}(\pi_i)$ generates circuits Γ_V^i, Γ_H^i . Finally, after k rounds $P_{kn}^{(g)}(\pi^{(k)})$ outputs a verification circuit

$$\Gamma_V^{(g)}(q, y_1, \dots, y_k) := g(\Gamma_V^1(q, y_1), \dots, \Gamma_V^k(q, y_k))$$

and a hint circuit

$$\Gamma_H^{(k)}(q) := (\Gamma_H^1(q), \dots, \Gamma_H^k(q)).$$

If it is clear from the context, we omit the subscript n and write $P(\pi)$ instead of $P_n(\pi)$ where $\pi \in \{0, 1\}^n$.

A verification query (q, y) of a solver C for which a hint query on this q has been asked before cannot be a verification query for which C succeeds. Therefore, without loss of generality, we make the assumption that C does not ask verification queries on q for which a hint query has been asked before. Furthermore, we assume that once C asked a verification query that succeeds, it does not ask any further hint or verification queries.

Experiment $\text{Success}^{P,C}(\pi, \rho)$

Oracle: A problem poser P , a solver $C = (C_1, C_2)$ for P .

Input: Bitstrings $\pi \in \{0, 1\}^n$, $\rho \in \{0, 1\}^*$.

Output: A bit $b \in \{0, 1\}$.

```

run  $\langle P(\pi), C_1(\rho) \rangle$ 
       $(\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P$ 
       $x := \langle P(\pi), C_1(\rho) \rangle_{\text{trans}}$ 

run  $C_2^{\Gamma_V, \Gamma_H}(x, \rho)$ 
      if  $C_2^{\Gamma_V, \Gamma_H}(x, \rho)$  asks a verification query  $(q, y)$  s.t.  $\Gamma_V(q, y) = 1$  then
        return 1
return 0
    
```

We define the *success probability* of C in solving a puzzle defined by P as

$$\Pr_{\pi, \rho}[\text{Success}^{P,C}(\pi, \rho) = 1]. \quad (4.1)$$

Furthermore, we say that C *succeeds* for π, ρ if $\text{Success}^{P,C}(\pi, \rho) = 1$.

Theorem 4.2 (Security amplification for dynamic weakly verifiable puzzles.)

Let $P_n^{(1)}$ be a fixed problem poser as in Definition 3.1 and $P_{kn}^{(g)}$ a problem poser for the k -wise direct product of $P_n^{(1)}$. Additionally, let C be a problem solver for $P_{kn}^{(g)}$ asking at most h hint queries and v verification queries. There exists a probabilistic algorithm Gen with oracle access to a solver circuit C , a monotone function $g : \{0, 1\}^k \rightarrow \{0, 1\}$ and problem posers $P_n^{(1)}, P_{kn}^{(g)}$. Furthermore, Gen takes as input parameters $\varepsilon, \delta, n, k, h, v$, and outputs a solver circuit D

for $P_n^{(1)}$ such that the following holds:
If C is such that

$$\Pr_{\substack{\pi^{(k)} \in \{0,1\}^{kn} \\ \rho \in \{0,1\}^*}} \left[\text{Success}^{P_{kn}^{(g)}, C}(\pi^{(k)}, \rho) = 1 \right] \geq 16(h + v) \left(\Pr_{u \leftarrow \mu_\delta^k} [g(u) = 1] + \varepsilon \right)$$

then D is a two phase probabilistic circuit and satisfies almost surely

$$\Pr_{\substack{\pi \in \{0,1\}^n \\ \rho \in \{0,1\}^*}} \left[\text{Success}^{P_n^{(1)}, D}(\pi, \rho) = 1 \right] \geq \delta + \frac{\varepsilon}{6k}.$$

Additionally, D requires oracle access to g , $P_n^{(1)}$, C , hint and verification circuits and asks at most $\frac{6k}{\varepsilon} \log \left(\frac{6k}{\varepsilon} \right) h$ hint queries and one verification query. Finally, $\text{Time}(\text{Gen}) = \text{poly}(k, \frac{1}{\varepsilon}, n, v, h)$ with oracle access to C .

4.1.2 Intuition

TODO: add intuition for hashing

The idea of the algorithm Gen is to output a circuit D that solves the input puzzle often. We know that C has high success probability in solving the k -wise direct product of $P^{(1)}$. The algorithm Gen tries to find a puzzle such that when C runs with this puzzle fixed on the first position and disregards whether this puzzle is correctly solved then the assumptions of Theorem 4.2 are true for the $(k - 1)$ -wise direct product. If it was possible to find such a puzzle, then Gen could recurse and solve a smaller problem. In the optimistic case we can reach $k = 1$, which means that we found a good circuit for solving a single puzzle by just fixing the initial puzzles of C .

Otherwise, when the first position is disregarded then the success probability of C is not substantially better. This is remarkable, as we know that C performs good for the k -wise direct product. It means that the first position is important, in the sense that C solves the puzzle on that position unusually often. Therefore, it is reasonable to construct the circuit D using C by placing the input puzzle of D on that position, and then finding remaining $k - 1$ puzzles. The $(k - 1)$ remaining puzzles are generated by the circuit D , hence it is possible to check whether they are correctly solved by the circuit C . We know that circuit C has good success probability, and the puzzle on the first position is important. Therefore, if we are able to find a $(k - 1)$ puzzles such that the fact whether the k -wise direct product is correctly solved depends on whether the puzzle on the first position is correctly solved then we can assume that C is often correct on this first position.

There are some problems with this approach, first we have to ensure that we can make a decision when the algorithm Gen should recurse and when not

correctly with high probability. Then, we have to show that it is possible to find a puzzles such that C is often correct on the first position. Finally, we also have to be sure that we do not ask a hint query, on the final verification query to the oracle. To satisfy the last requirement we split the set Q .

4.1.3 Domain partitioning

Let $hash : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$, the idea is to partition Q such that the set of preimages of 0 for $hash$ contains $q \in Q$ on which C is not allowed to ask hint queries, and the first successful verification query (q, y) of C is such that $hash(q) = 0$. Therefore, if C makes a verification query (q, y) such that $hash(q) = 0$, then we know that no hint query is ever asked on this q .

We denote the i -th query of C by q_i if it is a hint query, and by (q_i, y_i) if it is a verification query. We define the following experiment *CanonicalSuccess* in which the set Q is partitioned using a function $hash$. We say that a solver circuit *succeeds* in the experiment *CanonicalSuccess* if it asks a successful verification query (q_j, y_j) such that $hash(q_j) = 0$, and no hint query q_i is asked before (q_j, y_j) such that $hash(q_i) = 0$.

Experiment $CanonicalSuccess^{P,C,hash}(\pi, \rho)$

Oracle: A problem poser P , a solver circuit $C = (C_1, C_2)$ for P ,
a function $hash : Q \rightarrow \{0, \dots, 2(h + v) - 1\}$.

Input: Bitstrings $\pi \in \{0, 1\}^n$, $\rho \in \{0, 1\}^*$.

Output: A bit $b \in \{0, 1\}$.

```

run  $\langle P(\pi), C_1(\rho) \rangle$ 
       $(\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P$ 
       $x := \langle P(\pi), C_1(\rho) \rangle_{trans}$ 

run  $C_2^{\Gamma_V, \Gamma_H}(x, \rho)$ 
      if  $C_2^{\Gamma_V, \Gamma_H}(x, \rho)$  does not succeed for any verification query then
        return 0
      Let  $(q_j, y_j)$  be the first verification query such that  $\Gamma_V(q_j, y_j) = 1$ .

if  $(\forall i < j : hash(q_i) \neq 0)$  and  $(hash(q_j) = 0)$  then
  return 1
else
  return 0
    
```

We define the *canonical success probability* of a solver circuit C for P with

respect to a function *hash* as

$$\Pr_{\pi, \rho}[CanonicalSuccess^{P, C, hash}(\pi, \rho) = 1]. \quad (4.2)$$

For fixed *hash* and *P* a *canonical success* of *C* for bistrings π, ρ is a situation where $CanonicalSuccess^{P, C, hash}(\pi, \rho) = 1$.

We show that if a solver circuit *C* for *P* often succeeds in the experiment *Success*, then there exists a function *hash* such that *C* also often succeeds in the experiment *CanonicalSuccess*.

Lemma 4.3 (*Success probability in solving DWVP with respect to a function hash.*) *For fixed P_n let C be a solver for P_n with success probability at least γ , asking at most h hint queries and v verification queries. Let \mathcal{H} be an efficient family of pairwise independent hash functions $Q \rightarrow \{0, 1, \dots, 2(h+v)-1\}$. There exists a probabilistic algorithm *FindHash* that takes as input parameters γ, n, h, v , and has oracle access to C and P_n . Furthermore, *FindHash* runs in time $\text{poly}(h, v, \frac{1}{\gamma}, n)$, and with high probability outputs a function $hash \in \mathcal{H}$ such that the canonical success probability of C with respect to $hash$ is at least $\frac{\gamma}{16(h+v)}$.*

Proof (4.3). We fix a problem poser *P* and a solver *C* for *P* in the whole proof of Lemma 4.3. For $k, l \in \{1, \dots, (h+v)\}$ and $\alpha, \beta \in \{0, 1, \dots, 2(h+v)-1\}$ by the pairwise independence property, we have

$$\begin{aligned} \forall q_k \neq q_l \in Q : \Pr_{hash \leftarrow \mathcal{H}}[hash(q_k) = \alpha \mid hash(q_l) = \beta] &= \Pr_{hash \leftarrow \mathcal{H}}[hash(q_k) = \alpha] \\ &= \frac{1}{2(h+v)}. \end{aligned} \quad (4.3)$$

We write $\mathcal{P}_{Success}$ to denote a set containing all (π, ρ) for which $Success^{P, C}(\pi, \rho) = 1$. Let us fix $(\pi^*, \rho^*) \in \mathcal{P}_{Success}$. We are interested in the probability over a choice of function *hash* of the event $CanonicalSuccess^{P, C, hash}(\pi^*, \rho^*) = 1$. Let

(q_j, y_j) denote the first query such that $\Gamma_V(q_j, y_j) = 1$. We have

$$\begin{aligned}
 & \Pr_{hash \leftarrow \mathcal{H}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi^*, \rho^*) = 1 \right] \\
 &= \Pr_{hash \leftarrow \mathcal{H}} [hash(q_j) = 0 \wedge (\forall i < j : hash(q_i) \neq 0)] \\
 &= \Pr_{hash \leftarrow \mathcal{H}} [\forall i < j : hash(q_i) \neq 0 \mid hash(q_j) = 0] \Pr_{hash \leftarrow \mathcal{H}} [hash(q_j) = 0] \\
 &\stackrel{(4.3)}{=} \frac{1}{2(h+v)} \left(1 - \Pr_{hash \leftarrow \mathcal{H}} [\exists i < j : hash(q_i) = 0 \mid hash(q_j) = 0] \right) \\
 &\stackrel{(*)}{\geq} \frac{1}{2(h+v)} \left(1 - \sum_{i < j} \Pr_{hash \leftarrow \mathcal{H}} [hash(q_i) = 0 \mid hash(q_j) = 0] \right) \\
 &\stackrel{(4.3)}{=} \frac{1}{2(h+v)} \left(1 - \sum_{i < j} \Pr_{hash \leftarrow \mathcal{H}} [hash(q_i) = 0] \right) \\
 &\stackrel{(4.3)}{\geq} \frac{1}{4(h+v)}, \tag{4.4}
 \end{aligned}$$

where in $(*)$ we used the union bound. Let us denote the set of those (π, ρ) for which $\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1$ by $\mathcal{P}_{\text{Canonical}}$. If for π, ρ the circuit C succeeds canonically, then for the same π, ρ we also have $\text{Success}^{P,C}(\pi, \rho) = 1$. Hence, $\mathcal{P}_{\text{Canonical}} \subseteq \mathcal{P}_{\text{Success}}$, and we conclude

$$\begin{aligned}
 & \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1 \right] \\
 &= \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1 \mid (\pi, \rho) \in \mathcal{P}_{\text{Success}} \right] \Pr_{\pi, \rho} [(\pi, \rho) \in \mathcal{P}_{\text{Success}}] \\
 &\quad + \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1 \mid (\pi, \rho) \notin \mathcal{P}_{\text{Success}} \right] \Pr_{\pi, \rho} [(\pi, \rho) \notin \mathcal{P}_{\text{Success}}] \\
 &= \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1 \mid (\pi, \rho) \in \mathcal{P}_{\text{Success}} \right] \Pr_{\pi, \rho} [(\pi, \rho) \in \mathcal{P}_{\text{Success}}] \\
 &\geq \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} \left[\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1 \mid (\pi, \rho) \in \mathcal{P}_{\text{Success}} \right] \cdot \gamma \\
 &= \mathbb{E}_{(\pi, \rho) \in \mathcal{P}_{\text{Success}}} \left[\Pr_{hash \leftarrow \mathcal{H}} [\text{CanonicalSuccess}^{P,C,hash}(\pi, \rho) = 1] \right] \cdot \gamma \\
 &\stackrel{(4.4)}{\geq} \frac{\gamma}{4(h+v)}. \tag{4.5}
 \end{aligned}$$

Algorithm FindHash^{*P,C*}(γ, n, h, v)

Oracle: A problem poser P , a solver circuit C for P .

Input: Parameters γ, n . The number of hint queries h and of verification queries v .

Output: A function $hash : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$.

```

for  $i := 1$  to  $32n(h + v)^2/\gamma^2$  do:
   $hash \leftarrow \mathcal{H}$ 
   $count := 0$ 
  for  $j := 1$  to  $32n(h + v)^2/\gamma^2$  do:
     $\pi \leftarrow \{0, 1\}^n$ 
     $\rho \leftarrow \{0, 1\}^*$ 
    if  $CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1$  then
       $count := count + 1$ 
  if  $count \geq \frac{\gamma}{12(h+v)} \frac{32(h+v)^2}{\gamma^2} n$  then
    return  $hash$ 
return  $\perp$ 

```

We show that FindHash chooses $hash \in \mathcal{H}$ such that the canonical success probability of C with respect to $hash$ is at least $\frac{\gamma}{16(h+v)}$ almost surely. Let \mathcal{H}_{Good} denote a family of functions $hash \in \mathcal{H}$ for which

$$\Pr_{\pi, \rho} \left[CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1 \right] \geq \frac{\gamma}{8(h+v)}, \quad (4.6)$$

and \mathcal{H}_{Bad} be a family of functions $hash \in \mathcal{H}$ such that

$$\Pr_{\pi, \rho} \left[CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1 \right] \leq \frac{\gamma}{16(h+v)}. \quad (4.7)$$

Let N denote the number of iterations of the inner loop of FindHash. We consider a single iteration of the outer loop of FindHash in which $hash$ is fixed. We define independent, identically distributed, binary random variables X_1, \dots, X_N such that

$$X_i = \begin{cases} 1 & \text{if in the } i\text{-th iteration of the inner loop } count \text{ is increased} \\ 0 & \text{otherwise.} \end{cases}$$

We now turn to the case when $hash \in \mathcal{H}_{Bad}$ and show that it is unlikely that $hash$ is returned by FindHash. From (4.7) it follows that $\mathbb{E}_{\pi, \rho}[X_i] \leq \frac{\gamma}{16(h+v)}$. Therefore, for any fixed $hash \in \mathcal{H}_{Bad}$ using the Chernoff bound we get

$$\begin{aligned} \Pr_{\pi, \rho} \left[\frac{1}{N} \sum_{i=1}^N X_i \geq \frac{\gamma}{12(h+v)} \right] &\leq \Pr_{\pi, \rho} \left[\frac{1}{N} \sum_{i=1}^N X_i \geq \left(1 + \frac{1}{3}\right) \mathbb{E}[X_i] \right] \\ &\leq e^{-\frac{\gamma}{16(h+v)} N/27} \leq e^{-\frac{2}{27} \frac{(h+v)}{\gamma} n} \stackrel{(*)}{\leq} e^{-\frac{2}{27} n}, \end{aligned}$$

where in (*) we used the trivial facts that $h+v \geq 1$ and $\gamma \leq 1$. The probability that $hash \in \mathcal{H}_{Good}$, when picked, is not returned amounts

$$\begin{aligned} \Pr_{\pi, \rho} \left[\frac{1}{N} \sum_{i=1}^N X_i \leq \frac{\gamma}{12(h+v)} \right] &\leq \Pr_{\pi, \rho} \left[\frac{1}{N} \sum_{i=1}^N X_i \leq \left(1 - \frac{1}{3}\right) \mathbb{E}[X_i] \right] \\ &\leq e^{-\frac{\gamma}{8(h+v)} N/18} \leq e^{-\frac{2}{9} \frac{(h+v)}{\gamma} n} \stackrel{(*)}{\leq} e^{-\frac{2}{9} n}, \end{aligned}$$

where we once more used the Chernoff bound. We now show that the probability of picking $hash \in \mathcal{H}_{Good}$ is at least $\frac{\gamma}{8(h+v)}$. To obtain a contradiction suppose that

$$\Pr_{hash \leftarrow \mathcal{H}} [hash \in \mathcal{H}_{Good}] < \frac{\gamma}{8(h+v)}. \quad (4.8)$$

From this it follows that we can bound probability of canonical success as follows

$$\begin{aligned} &\Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} [CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1] \\ &= \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} [CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1 \mid hash \in \mathcal{H}_{Good}] \Pr_{hash \leftarrow \mathcal{H}} [hash \in \mathcal{H}_{Good}] \\ &\quad + \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} [CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1 \mid hash \notin \mathcal{H}_{Good}] \Pr_{hash \leftarrow \mathcal{H}} [hash \notin \mathcal{H}_{Good}] \\ &\leq \Pr_{hash \leftarrow \mathcal{H}} [hash \in \mathcal{H}_{Good}] + \Pr_{\substack{hash \leftarrow \mathcal{H} \\ \pi, \rho}} [CanonicalSuccess^{P,C,hash}(\pi, \rho) = 1 \mid hash \notin \mathcal{H}_{Good}] \\ &\stackrel{(4.6)}{<} \stackrel{(4.8)}{<} \frac{\gamma}{8(h+v)} + \frac{\gamma}{8(h+v)} = \frac{\gamma}{4(h+v)}, \end{aligned}$$

which contradicts (4.5). Therefore, we conclude that the probability of choosing a $hash \in \mathcal{H}_{Good}$ amounts at least $\frac{\gamma}{8(h+v)}$.

We show that FindHash picks in one of its iteration $hash \in \mathcal{H}_{Good}$ almost surely. Let K be the number of iterations of the outer loop of FindHash and Y_i be a random variable for the event that in the i -th iteration of the outer loop $hash \notin \mathcal{H}_{Good}$ is picked. We use $\Pr_{hash \leftarrow \mathcal{H}} [hash \in \mathcal{H}_{Good}] \geq \frac{\gamma}{8(h+v)}$ and $K \leq \frac{32(h+v)^2}{\gamma^2} n$, and conclude

$$\begin{aligned} \Pr_{hash \leftarrow \mathcal{H}} \left[\bigcap_{1 \leq i \leq K} Y_i \right] &\leq \left(1 - \frac{\gamma}{8(h+v)} \right)^{\frac{32(h+v)^2}{\gamma^2} n} \leq e^{-\frac{\gamma}{8(h+v)} \frac{32(h+v)^2}{\gamma^2} n} \\ &\leq e^{-\frac{4(h+v)}{\gamma} n} \leq e^{-n}. \end{aligned}$$

It is clear that running time of FindHash is $poly(n, h, v, \gamma)$ with oracle access. This finishes the proof of Lemma 4.3. \square

4.1.4 Amplification proof for partitioned domain

Let $C = (C_1, C_2)$ be a solver circuit for a dynamic weakly verifiable puzzle as in definition 3.1. We write $C_2^{(\cdot, \cdot)}$ to emphasize that C_2 does not obtain direct access to hint and verification circuits. Instead, whenever C_2 ask hint or verification queries, then it is answered explicitly as in the following code excerpt of the circuit \tilde{C}_2 .

Circuit $\tilde{C}_2^{\Gamma_H, C_2, hash}(x, \rho)$

Oracle: A hint circuit Γ_H , a circuit C_2 ,
a function $hash : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$.

Input: Bitstrings $x \in \{0, 1\}^*$, $\rho \in \{0, 1\}^*$.

Output: A pair (q, y) .

```

run  $C_2^{(\cdot, \cdot)}(x, \rho)$ 
  if  $C_2^{(\cdot, \cdot)}(x, \rho)$  asks a hint query on  $q$  then
    if  $hash(q) = 0$  then
      return  $\perp$ 
    else
      answer the query of  $C_2^{(\cdot, \cdot)}(x, \rho)$  using  $\Gamma_H(q)$ 

  if  $C_2^{(\cdot, \cdot)}(x, \rho)$  asks a verification query  $(q, y)$  then
    if  $hash(q) = 0$  then
      return  $(q, y)$ 
    else
      answer the verification query of  $C_2^{(\cdot, \cdot)}(x, \rho)$  with 0

return  $\perp$ 

```

Given $C = (C_1, C_2)$ we define the circuit $\tilde{C} = (C_1, \tilde{C}_2)$. Every hint query q asked by \tilde{C} is such that $hash(q) \neq 0$. Furthermore, \tilde{C} asks no verification queries. Instead, it returns (q, y) such that $hash(q) = 0$ or \perp .

For fixed π , ρ , and $hash$ we say that the circuit \tilde{C} *succeeds* if for $x := \langle P(\pi), C_1(\rho) \rangle_{trans}$, $(\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P$, we have

$$\Gamma_V(\tilde{C}_2^{\Gamma_H, C_2, hash}(x, \rho)) = 1.$$

Lemma 4.4 *For fixed P , C , and $hash$ the following statement is true*

$$\Pr_{\pi, \rho}[CanonicalSuccess^{P, C, hash}(\pi, \rho) = 1] \leq \Pr_{\substack{\pi, \rho \\ x := \langle P(\pi), C_1(\rho) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P}}[\Gamma_V(\tilde{C}_2^{\Gamma_H, C_2, hash}(x, \rho)) = 1]$$

Proof. If for fixed π , ρ , and $hash$ the circuit C succeeds canonically, then for the same π , ρ , and $hash$ also \tilde{C} succeeds. Using this observation, we conclude that

$$\begin{aligned}
 \Pr_{\pi, \rho} \left[CanonicalSuccess^{P, C, hash}(\pi, \rho) = 1 \right] & \\
 & \leq \mathbb{E}_{\substack{\pi, \rho \\ x := \langle P(\pi), C_1(\rho) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P}} [\Gamma_V(\tilde{C}_2^{\Gamma_H, C_2, hash}(x, \rho)) = 1] \\
 & = \Pr_{\substack{\pi, \rho \\ x := \langle P(\pi), C_1(\rho) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P(\pi), C_1(\rho) \rangle_P}} [\Gamma_V(\tilde{C}_2^{\Gamma_H, C_2, hash}(x, \rho)) = 1] \quad \square
 \end{aligned}$$

Lemma 4.5 (Security amplification for dynamic weakly verifiable puzzles with respect to hash.) Let $g : \{0, 1\}^k \rightarrow \{0, 1\}$ be a monotone function and $P_n^{(1)}$ a fixed problem poser as in Definition 3.1 and $\tilde{C} := (C_1, \tilde{C}_2)$ a circuit with oracle access to a function $hash : Q \rightarrow \{0, 1, \dots, 2(h + v - 1)\}$ and a solver circuit $C := (C_1, C_2)$ for $P_{kn}^{(g)}$ which asks at most h hint queries and v verification queries. There exists an algorithm Gen that takes as input parameters ε , δ , n , k , has oracle access to $P_n^{(1)}$, \tilde{C} , $hash$, g , and outputs a circuit $D := (D_1, D_2)$ such that the following holds:
If \tilde{C} is such that

$$\Pr_{\substack{\pi^{(k)} \in \{0, 1\}^{kn}, \rho \in \{0, 1\}^* \\ x := \langle P^{(g)}(\pi^{(k)}), C_1(\rho) \rangle_{trans} \\ (\Gamma_H^{(k)}, \Gamma_V^{(g)}) := \langle P^{(g)}(\pi^{(k)}), C_1(\rho) \rangle_{P^{(g)}}}} [\Gamma_V^{(g)}(\tilde{C}_2^{\Gamma_H^{(k)}, C_2, hash}(x, \rho)) = 1] \geq \Pr_{u \leftarrow \mu_\delta^k} [g(u) = 1] + \varepsilon,$$

then D satisfies almost surely

$$\Pr_{\substack{\pi \in \{0, 1\}^n, \rho \in \{0, 1\}^* \\ x := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{trans} \\ (\Gamma_H, \Gamma_V) := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{P^{(1)}}}} [\Gamma_V(D_2^{P^{(1)}, \tilde{C}, hash, g, \Gamma_H}(x, \rho)) = 1] \geq \delta + \frac{\varepsilon}{6k}.$$

Furthermore, D asks at most $\frac{6k}{\varepsilon} \log\left(\frac{6k}{\varepsilon}\right) h$ hint queries and no verification queries. Finally, $Time(Gen) = poly(k, \frac{1}{\varepsilon}, n)$ with oracle access to \tilde{C} .

Before we give the proof of Lemma 4.5 we define additional algorithms. First, we are interested in the probability that for $u \leftarrow \mu_\delta^k$ and a bit b we have $g(b, u_2, \dots, u_k) = 1$. The estimate of this probability is calculated by `EstimateFunctionProbability`.

Algorithm EstimateFunctionProbability^g(b, k, ε, δ, n)

Oracle: A function $g : \{0, 1\}^k \rightarrow \{0, 1\}$.

Input: A bit $b \in \{0, 1\}$, parameters $k, \varepsilon, \delta, n$.

Output: An estimate \tilde{g}_b of $\Pr_{u \leftarrow \mu_\delta^k}[g(b, u_2, \dots, u_k) = 1]$.

for $i := 1$ **to** $N := \frac{64k^2}{\varepsilon^2}n$ **do:**
 $u \leftarrow \mu_\delta^k$
 $g_i := g(b, u_2, \dots, u_k)$
return $\frac{1}{N} \sum_{i=1}^N g_i$

Lemma 4.6 *The algorithm EstimateFunctionProbability^g(b, k, ε, δ, n) outputs an estimate \tilde{g}_b such that $|\tilde{g}_b - \Pr_{u \leftarrow \mu_\delta^k}[g(b, u_2, \dots, u_k) = 1]| \leq \frac{\varepsilon}{8k}$ almost surely.*

Proof. We fix the notation as in the algorithm EstimateFunctionProbability. Let us define independent, identically distributed binary random variables K_1, K_2, \dots, K_N such that for each $1 \leq i \leq N$ the random variable K_i takes value g_i . We use the Chernoff bound to obtain

$$\begin{aligned} \Pr \left[\left| \tilde{g}_b - \Pr_{u \leftarrow \mu_\delta^k}[g(b, u_2, \dots, u_k) = 1] \right| \geq \frac{\varepsilon}{8k} \right] \\ = \Pr \left[\left| \left(\frac{1}{N} \sum_{i=1}^N K_i \right) - \mathbb{E}_{u \leftarrow \mu_\delta^k}[g(b, u_2, \dots, u_k)] \right| \geq \frac{\varepsilon}{8k} \right] \leq 2 \cdot e^{-n/3}. \square \end{aligned}$$

The algorithm EvaluatePuzzles^{P⁽¹⁾, \tilde{C} , hash}($\pi^{(k)}, \rho, n, k$) evaluates which of the k puzzles of the k -wise direct product defined by $P^{(g)}$ are solved successfully by $\tilde{C}(\rho) := (C_1, \tilde{C}_2)(\rho)$. To decide whether the i -th puzzle of the k -wise direct product is solved successfully we need to gain access to the verification circuit for the puzzle generated in the i -th round of the interaction between $P^{(g)}$ and \tilde{C} . Therefore, the algorithm EvaluatePuzzles runs k times $P^{(1)}$ to simulate the interaction with $C_1(\rho)$ each time with a fresh random bitstring $\pi_i \in \{0, 1\}^n$ where $1 \leq i \leq k$.

Let us introduce some additional notation. We denote by $\langle P^{(1)}(\pi_i), C_1(\rho) \rangle^i$ the execution of the i -th round of the sequential interaction. We use $\langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{P^{(1)}}^i$ to denote the output of $P^{(1)}(\pi_i)$ in the i -th round. Finally, we write $\langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{trans}^i$ to denote the transcript of communication in the i -th round. We note that the i -th round of the interaction between $P^{(1)}$ and C_1 is well defined only if all previous rounds have been executed before.

To make the notation easier in the code excerpts of circuits C_2 , D_2 and EvaluatePuzzles we omit superscripts of some oracles. Exemplary, we write $\tilde{C}_2^{\Gamma_H^{(k)}, hash}$ instead of $\tilde{C}_2^{\Gamma_H^{(k)}, C, hash}$ where the superscript of the oracle circuit C

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is omitted. We make sure that it is clear from the context which oracles are used.

Algorithm EvaluatePuzzles $^{P^{(1)}, \tilde{C}, hash}(\pi^{(k)}, \rho, n, k)$

Oracle: A problem poser $P^{(1)}$, a solver circuit $\tilde{C} = (C_1, \tilde{C}_2)$ for $P^{(g)}$,
a function $hash : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$.

Input: Bitstrings $\pi^{(k)} \in \{0, 1\}^{kn}$, $\rho \in \{0, 1\}^*$, parameters n, k .

Output: A tuple $(c_1, \dots, c_k) \in \{0, 1\}^k$.

for $i := 1$ **to** k **do:** *//simulate k rounds of interaction*
 $(\Gamma_V^i, \Gamma_H^i) := \langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{P^{(1)}}^i$
 $x_i := \langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{trans}^i$
 $x := (x_1, \dots, x_k)$
 $\Gamma_H^{(k)} := (\Gamma_H^1, \dots, \Gamma_H^k)$
 $(q, y_1, \dots, y_k) := \tilde{C}_2^{\Gamma_H^{(k)}, hash}(x, \rho)$
if $(q, y_1, \dots, y_k) = \perp$ **then**
 return $(0, \dots, 0)$
 $(c_1, \dots, c_k) := (\Gamma_V^1(q, y_1), \dots, \Gamma_V^k(q, y_k))$
return (c_1, \dots, c_k)

All puzzles used by EvaluatePuzzles are generated internally thus the algorithm has access to hint circuit, and can answer itself all queries of \tilde{C}_2 .

We are interested in the success probability of \tilde{C} with the bitstring π_1 fixed to π^* where the fact whether \tilde{C} succeeds in solving the first puzzle defined by $P^{(1)}(\pi_1)$ is neglected, and instead a bit b is used. More formally, we define the surplus $S_{\pi^*, b}$ as

$$S_{\pi^*, b} = \Pr_{\pi^{(k)}, \rho} [g(b, c_2, \dots, c_k) = 1 \mid \pi_1 = \pi^*] - \Pr_{u \leftarrow \mu_\delta^k} [g(b, u_2, \dots, u_k) = 1], \quad (4.9)$$

where (c_2, c_3, \dots, c_k) is obtained as in EvaluatePuzzles.

The algorithm EstimateSurplus returns an estimate $\tilde{S}_{\pi^*, b}$ for $S_{\pi^*, b}$.

Algorithm EstimateSurplus $^{P^{(1)}, \tilde{C}, g, hash}(\pi^*, b, k, \varepsilon, \delta, n)$

Oracle: A problem poser $P^{(1)}$, a circuit \tilde{C} for $P^{(g)}$, functions
 $g : \{0, 1\}^k \rightarrow \{0, 1\}$ and $hash : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$.

Input: A bistring $\pi^* \in \{0, 1\}^n$, a bit $b \in \{0, 1\}$, parameters $k, \varepsilon, \delta, n$.

Output: An estimate $\tilde{S}_{\pi^*, b}$ for $S_{\pi^*, b}$.

```

for  $i := 1$  to  $N := \frac{64k^2}{\varepsilon^2}n$  do:
     $(\pi_2, \dots, \pi_k) \xleftarrow{\$} \{0, 1\}^{(k-1)n}$ 
     $\rho \xleftarrow{\$} \{0, 1\}^*$ 
     $(c_1, \dots, c_k) := \text{EvaluatePuzzles}^{P^{(1)}, \tilde{C}, \text{hash}}((\pi^*, \pi_2, \dots, \pi_k), \rho, n, k)$ 
     $\tilde{s}_{\pi^*, b}^i := g(b, c_2, \dots, c_k)$ 
 $\tilde{g}_b := \text{EstimateFunctionProbability}^g(b, k, \varepsilon, \delta, n)$ 
return  $\left( \frac{1}{N} \sum_{i=1}^N \tilde{s}_{\pi^*, b}^i \right) - \tilde{g}_b$ 

```

Lemma 4.7 *The estimate $\tilde{S}_{\pi^*, b}$ returned by *EstimateSurplus* differs from $S_{\pi^*, b}$ by at most $\frac{\varepsilon}{4k}$ almost surely.*

Proof. We use the union bound and similar argument as in Lemma 4.6 which yields that $\frac{1}{N} \sum_{i=1}^N \tilde{s}_{\pi^*, b}^i$ differs from $\mathbb{E}[g(b, c_2, \dots, c_k)]$ by at most $\frac{\varepsilon}{8k}$ almost surely. Together, with Lemma 4.6 we conclude that the surplus estimate returned by *EstimateSurplus* differs from $S_{\pi^*, b}$ by at most $\frac{\varepsilon}{4k}$ almost surely. \square

We define the following solver circuit $C' = (C'_1, C'_2)$ for the $(k-1)$ -wise direct product of $P^{(1)}$.

Circuit $C'_1{}^{\tilde{C}, P^{(1)}}(\rho)$

Oracle: A solver circuit $\tilde{C} = (C_1, \tilde{C}_2)$ for $P^{(g)}$, a poser $P^{(1)}$.

Input: A bitstring $\rho \in \{0, 1\}^*$.

Hard-coded: A bitstring $\pi^* \in \{0, 1\}^n$.

Simulate $\langle P^{(1)}(\pi^*), C_1(\rho) \rangle^1$

Use $C_1(\rho)$ for the remaining $k-1$ rounds of interaction.

Circuit $\tilde{C}_2{}^{\Gamma_H^{(k-1)}, \tilde{C}, \text{hash}}(x^{(k-1)}, \rho)$

Oracle: A hint oracle $\Gamma_H^{(k-1)} := (\Gamma_H^2, \dots, \Gamma_H^k)$,
a solver circuit $\tilde{C} = (C_1, \tilde{C}_2)$ for $P^{(g)}$,
a function $\text{hash} : Q \rightarrow \{0, 1, \dots, 2(h+v)-1\}$.

Input: A transcript of $k-1$ rounds of interaction

$x^{(k-1)} := (x_2, \dots, x_k) \in \{0, 1\}^*$, a bitstring $\rho \in \{0, 1\}^*$

Hard-coded: A bitstring $\pi^* \in \{0, 1\}^n$

Simulate $\langle P^{(1)}(\pi^*), C_1(\rho) \rangle^1$

$(\Gamma_H^*, \Gamma_V^*) := \langle P^{(1)}(\pi^*), C_1(\rho) \rangle_{P^{(1)}}^1$

$x^* := \langle P^{(1)}(\pi^*), C_1(\rho) \rangle_{\text{trans}}^1$

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 $\Gamma_H^{(k)} := (\Gamma_H^*, \Gamma_H^2, \dots, \Gamma_H^k)$ 
 $x^{(k)} := (x^*, x_2, \dots, x_k)$ 
 $(q, y_1, \dots, y_k) := \tilde{C}_2^{\Gamma_H^{(k)}, hash}(x^{(k)}, \rho)$ 
return  $(q, y_2, \dots, y_k)$ 

```

We are ready to define the solver circuit $D = (D_1, D_2)$ for $P^{(1)}$ and the algorithm Gen.

Circuit $D_1^{\tilde{C}}(r)$

Oracle: A solver circuit $\tilde{C} = (C_1, \tilde{C}_2)$ for $P^{(g)}$.

Input: A pair $r := (\rho, \sigma)$ where $\rho \in \{0, 1\}^*$ and $\sigma \in \{0, 1\}^*$.

Interact with the problem poser $\langle P^{(1)}, C_1(\rho) \rangle^1$.

Let $x^* := \langle P^{(1)}, C_1(\rho) \rangle_{trans}^1$.

Circuit $D_2^{P^{(1)}, \tilde{C}, hash, g, \Gamma_H}(x^*, r)$

Oracle: A poser $P^{(1)}$, a solver circuit $\tilde{C} = (C_1, \tilde{C}_2)$ for $P^{(g)}$,
functions $hash : Q \rightarrow \{0, 1, \dots, 2(h+v)-1\}$, $g : \{0, 1\}^k \rightarrow \{0, 1\}$,
a hint circuit Γ_H for $P^{(1)}$.

Input: A communication transcript $x^* \in \{0, 1\}^*$, a bitstring $r := (\rho, \sigma)$
where $\rho \in \{0, 1\}^*$ and $\sigma \in \{0, 1\}^*$

Output: A pair (q, y^*) .

for at most $\frac{6k}{\epsilon} \log(\frac{6k}{\epsilon})$ iterations **do:**

$(\pi_2, \dots, \pi_k) \leftarrow$ read next $(k-1) \cdot n$ bits from σ

Use x^* to simulate the first round of interaction of $C_1(\rho)$ with the problem poser $P^{(1)}$

for $i := 2$ **to** k **do:**

run $\langle P^{(1)}(\pi_i), C_1(\rho) \rangle^i$
 $(\Gamma_V^i, \Gamma_H^i) := \langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{P^{(1)}}^i$

$x_i := \langle P^{(1)}(\pi_i), C_1(\rho) \rangle_{trans}^i$

$\Gamma_H^{(k)}(q) := (\Gamma_H(q), \Gamma_H^2(q), \dots, \Gamma_H^k(q))$

$(q, y^*, y_2, \dots, y_k) := \tilde{C}_2^{\Gamma_H^{(k)}, hash}((x^*, x_2, \dots, x_k), \rho)$

$(c_2, \dots, c_k) := (\Gamma_V^2(q, y_2), \dots, \Gamma_V^k(q, y_k))$

if $g(1, c_2, \dots, c_k) = 1$ **and** $g(0, c_2, \dots, c_k) = 0$ **then**

return (q, y^*)

return \perp

Algorithm $\text{Gen}^{P^{(1)}, \tilde{C}, g, \text{hash}}(\varepsilon, \delta, n, k)$

Oracle: A poser $P^{(1)}$, a solver circuit \tilde{C} for $P^{(g)}$, functions $g : \{0, 1\}^k \rightarrow \{0, 1\}$,

$\text{hash} : Q \rightarrow \{0, 1, \dots, 2(h + v) - 1\}$.

Input: Parameters $\varepsilon, \delta, n, k$.

Output: A circuit D .

for $i := 1$ **to** $\frac{6k}{\varepsilon}n$ **do:**

$\pi^* \xleftarrow{\$} \{0, 1\}^n$

$\tilde{S}_{\pi^*, 0} := \text{EstimateSurplus}^{P^{(1)}, \tilde{C}, g, \text{hash}}(\pi^*, 0, k, \varepsilon, \delta, n)$

$\tilde{S}_{\pi^*, 1} := \text{EstimateSurplus}^{P^{(1)}, \tilde{C}, g, \text{hash}}(\pi^*, 1, k, \varepsilon, \delta, n)$

if $\exists b \in \{0, 1\} : \tilde{S}_{\pi^*, b} \geq (1 - \frac{3}{4k})\varepsilon$ **then**

Let C'_1 have oracle access to \tilde{C} , and have hard-coded π^*

Let \tilde{C}'_2 have oracle access to \tilde{C} , and have hard-coded π^* .

$\tilde{C}' := (C'_1, \tilde{C}'_2)$

$g'(b_2, \dots, b_k) := g(b, b_2, \dots, b_k)$

return $\text{Gen}^{P^{(1)}, \tilde{C}', g', \text{hash}}(\varepsilon, \delta, n, k - 1)$

// all estimates are lower than $(1 - \frac{3}{4k})\varepsilon$

return $D^{P^{(1)}, \tilde{C}, \text{hash}, g}$

Proof (Lemma 4.5). First let us consider the case where $k = 1$. The function $g : \{0, 1\} \rightarrow \{0, 1\}$ is either the identity or a constant function. If g is the identity function, then the circuit D returned by Gen directly uses \tilde{C} to find a solution. From the assumptions of Lemma 4.5 it follows that \tilde{C} succeeds with probability at least $\delta + \varepsilon$. Hence, D trivially satisfies the statement of Lemma 4.5. In the latter case g is a constant function, and the statement is vacuously true.

For the general case, we consider two possibilities. Either Gen in one of the iterations finds an estimate $\tilde{S}_{\pi^*, b} \geq (1 - \frac{3}{4k})\varepsilon$ or it fails and returns the circuit D .

In the former case we define a new monotone function $g'(b_2, \dots, b_k) := g(b, b_2, \dots, b_k)$ and a new circuit $\tilde{C}' = (C'_1, \tilde{C}'_2)$ with oracle access to $\tilde{C} := (C_1, \tilde{C}_2)$. By Lemma 4.7 it follows that $S_{\pi^*, b} \geq \tilde{S}_{\pi^*, b} - \frac{\varepsilon}{4k} \geq (1 - \frac{1}{k})\varepsilon$ almost surely. Therefore, the circuit \tilde{C}' succeeds in solving the $(k-1)$ -wise direct product of puzzles with probability at least $\Pr_{u \leftarrow \mu_\delta^{(k-1)}}[g'(u_1, \dots, u_{k-1})] + (1 - \frac{1}{k})\varepsilon$. In this case \tilde{C}' satisfies the conditions of Lemma 4.5 for the $(k-1)$ -wise direct product of puzzles. Therefore, the recursive call to Gen with access to g' and \tilde{C} returns

a circuit $D = (D_1, D_2)$ that with high probability satisfies

$$\Pr_{\pi, \rho}[\Gamma_V(D_2^{P^{(1)}, \tilde{C}, hash, g, \Gamma_H}(x, \rho)) = 1] \geq \delta + \left(1 - \frac{1}{k}\right) \frac{\varepsilon}{6(k-1)} = \delta + \frac{\varepsilon}{6k}. \quad (4.10)$$

$x := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{trans}$
 $(\Gamma_H, \Gamma_V) := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{P^{(1)}}$

The only point remaining concerns the behavior of Gen when none of the estimates is greater than $(1 - \frac{3}{4k})\varepsilon$. Assume that...

TODO : Give more intuition (and the correct one with references to the equations)

Intuitively this means that \tilde{C} does not succeed on the remaining $k-1$ puzzles with much higher probability than an algorithm that correctly solves each puzzle with probability δ . However, from the assumptions of Lemma 4.5 we know that on all k puzzles the success probability of \tilde{C} is higher. Therefore, it is likely that the first puzzle is correctly solved unusually often. It remains to prove that this intuition is indeed correct.

We fix the notation used in the code excerpt of the circuit D_2 . We consider a single iteration of the outer loop of D_2 , in which values π_2, \dots, π_k are fixed. Additionally, we write π_1 to denote the randomness of the problem poser and define $c_1 := \Gamma_V(q, y_1)$, where Γ_V is the verification circuit generated by $P^{(1)}(\pi_1)$ in the first phase when interacting with $D_1(r)$. Finally, we introduce the additional notation $\mathcal{G}_b := \{(b_1, b_2, \dots, b_k) : g(b, b_2, \dots, b_k) = 1\}$ and $c = (c_1, c_2, \dots, c_k)$. Using the new notation the following holds

$$\begin{aligned} \Pr_{u \leftarrow \mu_\delta^k}[u \in \mathcal{G}_b] &= \Pr_{u \leftarrow \mu_\delta^k}[g(b, u_2, \dots, u_k) = 1] \\ \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_b] &= \Pr_{\pi^{(k)}, \rho}[g(b, c_2, \dots, c_k) = 1]. \end{aligned} \quad (4.11)$$

We fix the randomness π_1 of the problem poser $P^{(1)}$ to π^* and use (4.9), (4.11) to obtain

$$\begin{aligned} &\Pr_{u \leftarrow \mu_\delta^k}[u \in \mathcal{G}_1] - \Pr_{u \leftarrow \mu_\delta^k}[u \in \mathcal{G}_0] \\ &= \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_1 \mid \pi_1 = \pi^*] - \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_0 \mid \pi_1 = \pi^*] - (S_{\pi^*, 1} - S_{\pi^*, 0}) \end{aligned} \quad (4.12)$$

We know that the function g is monotone, hence it holds $\mathcal{G}_0 \subseteq \mathcal{G}_1$, and we write (4.12) as

$$\Pr_{u \leftarrow \mu_\delta^k}[u \in \mathcal{G}_1 \setminus \mathcal{G}_0] = \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] - (S_{\pi^*, 1} - S_{\pi^*, 0}). \quad (4.13)$$

Still having $\pi_1 = \pi^*$ fixed we multiply both sides of (4.13) by

$$\Pr_r[\Gamma_V(D_2(x^*, r)) = 1] / \Pr_{u \leftarrow \mu_\delta^k}[u \in \mathcal{G}_1 \setminus \mathcal{G}_0],$$

$x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}$
 $(\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P^{(1)}}$

which yields

$$\begin{aligned}
 & \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1] \\
 &= \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1] \Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] \frac{1}{\Pr_{u \leftarrow \mu_\delta^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \\
 &\quad - \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1] (S_{\pi^*, 1} - S_{\pi^*, 0}) \frac{1}{\Pr_{u \leftarrow \mu_\delta^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0]}.
 \end{aligned} \tag{4.14}$$

We analyze the first summand of (4.14). First, we have

$$\begin{aligned}
 & \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1] \\
 &= \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1 \mid D_2(x^*, r) \neq \perp] \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}}} [D_2(x^*, r) \neq \perp] \\
 &\stackrel{(*)}{=} \Pr_{\substack{\pi^{(k)}, \rho \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}}} [c_1 = 1 \mid c \in \mathcal{G}_1 \setminus \mathcal{G}_0, \pi_1 = \pi^*] \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}}} [D_2(x^*, r) \neq \perp], \tag{4.15}
 \end{aligned}$$

where in $(*)$ we use the observation that the event $D_2(x^*, r) \neq \perp$ happens if and only if the circuit $D_2(x^*, r)$ finds $\pi^{(k)}$ such that $c \in \mathcal{G}_1 \setminus \mathcal{G}_0$. Furthermore, conditioned on $c \in \mathcal{G}_1 \setminus \mathcal{G}_0$ we have that $\Gamma_V(D_2(x^*, r)) = 1$ occurs if and only if $c_1 = 1$. Inserting (4.15) to the numerator of the first summand of (4.14) yields

$$\begin{aligned}
 & \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}}} [\Gamma_V(D_2(x^*, r)) = 1] \Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] \\
 &= \Pr_{\substack{r \\ x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}}} [D_2(x^*, r) \neq \perp] \Pr_{\pi^{(k)}, \rho} [c_1 = 1 \mid c \in \mathcal{G}_1 \setminus \mathcal{G}_0, \pi_1 = \pi^*] \Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*].
 \end{aligned} \tag{4.16}$$

We consider the following two cases. If $\Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] \leq \frac{\varepsilon}{6k}$ then

$$\Pr_{\pi^{(k)}, \rho} [c_1 = 1 \mid c \in \mathcal{G}_1 \setminus \mathcal{G}_0, \pi_1 = \pi^*] \Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] \leq \frac{\varepsilon}{6k}. \tag{4.17}$$

When $\Pr_{\pi^{(k)}, \rho} [c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] > \frac{\varepsilon}{6k}$ the circuit D_2 outputs \perp if and only if it fails in all $\frac{6k}{\varepsilon} \log(\frac{6k}{\varepsilon})$ iterations to find $\pi^{(k)}$ such that $c \in \mathcal{G}_1 \setminus \mathcal{G}_0$ which

happens with probability

$$\Pr_r[D_2(x^*, r) = \perp] \leq (1 - \frac{\varepsilon}{6k})^{\frac{6k}{\varepsilon} \log(\frac{6k}{\varepsilon})} \leq \frac{\varepsilon}{6k}. \quad (4.18)$$

$x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}$

We conclude that in both cases by (4.17) and (4.18) we have

$$\begin{aligned} & \Pr_r[D_2(x^*, r) \neq \perp] \Pr_{\pi^{(k)}, \rho}[c_1 = 1 \mid c \in \mathcal{G}_1 \setminus \mathcal{G}_0, \pi_1 = \pi^*] \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] \\ & \geq \Pr_{\pi^{(k)}, \rho}[c_1 = 1 \mid c \in \mathcal{G}_1 \setminus \mathcal{G}_0, \pi_1 = \pi^*] \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] - \frac{\varepsilon}{6k} \\ & = \Pr_{\pi^{(k)}, \rho}[c_1 = 1 \wedge c \in \mathcal{G}_1 \setminus \mathcal{G}_0 \mid \pi_1 = \pi^*] - \frac{\varepsilon}{6k} \\ & = \Pr_{\pi^{(k)}, \rho}[g(c) = 1 \mid \pi_1 = \pi^*] - \Pr_{\pi^{(k)}, \rho}[c \in \mathcal{G}_0 \mid \pi_1 = \pi^*] - \frac{\varepsilon}{6k} \\ & \stackrel{(4.9)}{=} \Pr_{\pi^{(k)}, \rho}[g(c) = 1 \mid \pi_1 = \pi^*] - \Pr_{u \leftarrow \mu_\delta^{(k)}}[u \in \mathcal{G}_0] - S_{\pi^*, 0} - \frac{\varepsilon}{6k}. \end{aligned} \quad (4.19)$$

$x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}$

We insert (4.19) to (4.14) and calculate the expected value of over π^* which yields

$$\begin{aligned} & \Pr_{\pi, r}[\Gamma_V(D_2(x, r)) = 1] \geq \mathbb{E}_{\pi^*} \left[\frac{\Pr_{\pi^{(k)}, \rho}[g(c) = 1 \mid \pi_1 = \pi^*] - \Pr_{u \leftarrow \mu_\delta^{(k)}}[u \in \mathcal{G}_0] - \frac{\varepsilon}{6k}}{\Pr_{u \leftarrow \mu_\delta^{(k)}}[u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \right] \\ & \quad - \mathbb{E}_{\pi^*} \left[\left(\Pr_r[\Gamma_V(D_2(x^*, r)) = 1] (S_{\pi^*, 1} - S_{\pi^*, 0}) + S_{\pi^*, 0} \right) \frac{1}{\Pr_{u \leftarrow \mu_\delta^{(k)}}[u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \right]. \end{aligned}$$

$x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{trans}$
 $(\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P(1)}$

(4.20)

We show that if Gen does not recurse, then the majority of estimates is low almost surely. Let us assume that

$$\Pr_\pi \left[\left(S_{\pi, 0} \leq (1 - \frac{1}{2k})\varepsilon \right) \wedge \left(S_{\pi, 1} \leq (1 - \frac{1}{2k})\varepsilon \right) \right] < 1 - \frac{\varepsilon}{6k}, \quad (4.21)$$

then Gen recurses almost surely, because the probability that Gen does not find $\tilde{S}_{\pi, b} \geq (1 - \frac{3}{4k})\varepsilon$ in all of the $\frac{6k}{\varepsilon}n$ iterations is at most

$$\left(1 - \frac{\varepsilon}{6k} \right)^{\frac{6k}{\varepsilon}n} \leq e^{-n}$$

almost surely, where we used Lemma 4.7. Therefore, under the assumption that Gen does not recurse with high probability it holds

$$\Pr_{\pi, \rho} \left[\left(S_{\pi, 0} \leq (1 - \frac{1}{2k})\varepsilon \right) \wedge \left(S_{\pi, 1} \leq (1 - \frac{1}{2k})\varepsilon \right) \right] \geq 1 - \frac{\varepsilon}{6k}. \quad (4.22)$$

Let us define a set

$$\mathcal{W} = \left\{ \pi : \left(S_{\pi,0} \leq \left(1 - \frac{1}{2k}\right)\varepsilon \right) \wedge \left(S_{\pi,1} \leq \left(1 - \frac{1}{2k}\right)\varepsilon \right) \right\}, \quad (4.23)$$

and use $\overline{\mathcal{W}}$ to denote the complement of \mathcal{W} . We bound the numerator of the second summand in (4.20)

$$\begin{aligned} & \mathbb{E}_{\pi^*} [S_{\pi^*,0} + \Pr_{\substack{x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P^{(1)}}}} [\Gamma_V(D_2(x^*, r)) = 1] (S_{\pi^*,1} - S_{\pi^*,0})] \\ &= \mathbb{E}_{\pi^*} \left[S_{\pi^*,0} + \Pr_{\substack{x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P^{(1)}}}} [\Gamma_V(D_2(x^*, r)) = 1] (S_{\pi^*,1} - S_{\pi^*,0}) \mid \pi^* \in \overline{\mathcal{W}} \right] \Pr_{\pi^*} [\pi^* \in \overline{\mathcal{W}}] \\ &+ \mathbb{E}_{\pi^*} \left[S_{\pi^*,0} + \Pr_{\substack{x^* := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P^{(1)}}}} [\Gamma_V(D_2(x^*, r)) = 1] (S_{\pi^*,1} - S_{\pi^*,0}) \mid \pi^* \in \mathcal{W} \right] \Pr_{\pi^*} [\pi^* \in \mathcal{W}] \\ &\leq \frac{\varepsilon}{6k} + \mathbb{E}_{\pi^*} \left[S_{\pi^*,0} + \Pr_{\substack{x := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(r) \rangle_{P^{(1)}}}} [\Gamma_V(D_2^{\tilde{C}}(x^*, r)) = 1] \left(\left(1 - \frac{1}{2k}\right)\varepsilon - S_{\pi^*,0} \right) \mid \pi^* \in \mathcal{W} \right] \Pr_{\pi^*} [\pi^* \in \mathcal{W}] \\ &\leq \frac{\varepsilon}{6k} + \left(1 - \frac{1}{2k}\right)\varepsilon = \left(1 - \frac{1}{3k}\right)\varepsilon. \end{aligned} \quad (4.24)$$

We observe that

$$\begin{aligned} \Pr_{u \leftarrow \mu_{\delta}^k} [g(u) = 1] &= \Pr[u \in \mathcal{G}_0 \vee (u \in \mathcal{G}_1 \setminus \mathcal{G}_0 \wedge u_1 = 1)] \\ &= \Pr[u \in \mathcal{G}_0] + \delta \Pr[u \in \mathcal{G}_1 \setminus \mathcal{G}_0]. \end{aligned} \quad (4.25)$$

Finally, we insert (4.24) into (4.20) which yields

$$\Pr_{\substack{x := \langle P^{(1)}(\pi), D_1(\rho) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi), D_1(\rho) \rangle_{P^{(1)}}}} [\Gamma_V(D_2(x, \rho)) = 1] \geq \mathbb{E}_{\pi^*} \left[\frac{\Pr_{\pi^{(k)}, \rho} [g(c) = 1 \mid \pi_1 = \pi^*] - \Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_0] - \left(1 - \frac{1}{6k}\right)\varepsilon}{\Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \right].$$

From the assumptions of Lemma 4.5 it follows that

$$\Pr_{\pi^{(k)}, \rho} [g(c) = 1] \geq \Pr_{u \leftarrow \mu_{\delta}^{(k)}} [g(u) = 1] + \varepsilon.$$

thus we get

$$\begin{aligned} \Pr_{\substack{x := \langle P^{(1)}(\pi^*), D_1(\rho) \rangle_{\text{trans}} \\ (\Gamma_V, \Gamma_H) := \langle P^{(1)}(\pi^*), D_1(\rho) \rangle_{P^{(1)}}}} [\Gamma_V(D_2(x, \rho)) = 1] &\geq \frac{\Pr_{u \leftarrow \mu_{\delta}^k} [g(u) = 1] + \varepsilon - \Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_0] - \left(1 - \frac{1}{6k}\right)\varepsilon}{\Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \\ &\stackrel{(4.25)}{\geq} \frac{\varepsilon + \delta \Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0] - \left(1 - \frac{1}{6k}\right)\varepsilon}{\Pr_{u \leftarrow \mu_{\delta}^k} [u \in \mathcal{G}_1 \setminus \mathcal{G}_0]} \geq \delta + \frac{\varepsilon}{6k} \end{aligned} \quad (4.26)$$

Clearly, the running time of Gen is $\text{poly}(k, \frac{1}{\varepsilon}, n)$. \square

4.1.5 Putting it together

In the previous sections we have shown that it is possible to partition the domain Q such that if the number of hint and verification queries is small, then success probability of a solver for the k -wise direct product is still substantial. As shown in Lemma 4.5 we can build a circuit that returns a solution that is likely to be good. It remains to use these build blocks and prove Theorem 4.2.

Proof (of Theorem 4.2). Let $\widetilde{\text{Gen}}$ be the algorithm from Theorem 4.2, and $\widetilde{D} = (D_1, D_2)$ the corresponding solver circuit for P that is output by $\widetilde{\text{Gen}}$ as in Theorem 4.2. They are defined on the following code excerpts.

Circuit $\widetilde{D}_2^{D, P^{(1)}, \text{hash}, g, \Gamma_V, \Gamma_H}(x, \rho)$

Oracle: A circuit $D := (D_1, \widetilde{D}_2)$ from Lemma 4.5, a problem poser $P^{(1)}$, functions $\text{hash} : Q \rightarrow \{0, 1, \dots, 2(h+v)-1\}$, $g : \{0, 1\}^k \rightarrow \{0, 1\}$ a verification oracle Γ_V , a hint oracle Γ_H .

Input: Bitstrings $x \in \{0, 1\}^*$, $\rho \in \{0, 1\}^*$.

$(q, y) := D_2^{P^{(1)}, \widetilde{C}, \text{hash}, g, \Gamma_H}(x, \rho)$
Ask verification query (q, y) to Γ_V .

Algorithm $\widetilde{\text{Gen}}^{P^{(1)}, g, C}(n, \varepsilon, \delta, k, h, v)$

Oracle: A problem poser $P^{(1)}$, a function $g : \{0, 1\}^k \rightarrow \{0, 1\}$, a solver circuit C for $P^{(g)}$.

Input: Parameters $n, \varepsilon, \delta, k, h, v$.

Return: A circuit $\widetilde{D} = (D_1, \widetilde{D}_2)$.

$\text{hash} := \text{FindHash}((h+v)\varepsilon, n, h, v)$

Let $\widetilde{C} := (C_1, \widetilde{C}_2)$ be as in Lemma 4.4 with oracle access to C , hash .

$D := \text{Gen}^{P^{(1)}, \widetilde{C}, g, \text{hash}}(\varepsilon, \delta, n, k)$

return $\widetilde{D} := (D_1, \widetilde{D}_2)$

We show that Theorem 4.2 follows from Lemma 4.3 and Lemma 4.5. We fix $P^{(1)}$, g , $P^{(g)}$ in the whole proof and consider a solver circuit $C = (C_1, C_2)$, asking at most h hint queries and v verification queries, such that

$$\Pr_{\pi^{(k)}, \rho} \left[\text{Success}^{P^{(g)}, C}(\pi^{(k)}, \rho) = 1 \right] \geq 16(h+v) \left(\Pr_{u \leftarrow \mu_\delta^k} [g(u) = 1] + \varepsilon \right).$$

First, we note that C meets the requirements of Lemma 4.3. Thus, the algorithm $\widetilde{\text{Gen}}$ can call FindHash to obtain $hash$ such that with high probability it holds

$$\Pr_{\pi^{(k)}, \rho} \left[\text{CanonicalSuccess}^{P^{(g)}, C, hash}(\pi^{(k)}, \rho) = 1 \right] \geq \Pr_{u \leftarrow \mu_\delta^k} [g(u) = 1] + \varepsilon.$$

Applying Lemma 4.4 for C we obtain a circuit $\tilde{C} = (C_1, \tilde{C}_2)$ that satisfies

$$\Pr_{\pi^{(k)}, \rho} [\Gamma_V^{(g)}(\tilde{C}_2^{\Gamma_H^{(k)}, C_2, hash}(x, \rho)) = 1] \geq \Pr_{u \leftarrow \mu_\delta^k} [g(u) = 1] + \varepsilon.$$

$x := \langle P^{(g)}(\pi^{(k)}), C_1(\rho) \rangle_{trans}$
 $(\Gamma_V^{(g)}, \Gamma_H^{(k)}) := \langle P^{(g)}(\pi^{(k)}), C_1(\rho) \rangle_{P^{(g)}}$

Now, we use the algorithm Gen as in Lemma 4.5 that yields a circuit $D = (D_1, D_2)$ which with high probability satisfies

$$\Pr_{\pi, \rho} [\Gamma_V(D_2^{P^{(1)}, \tilde{C}, hash, g, \Gamma_H}(x, \rho)) = 1] \geq (\delta + \frac{\varepsilon}{6k}). \quad (4.27)$$

$x := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{trans}$
 $(\Gamma_H, \Gamma_V) := \langle P^{(1)}(\pi), D_1^{\tilde{C}}(\rho) \rangle_{P^{(1)}}$

Finally, $\widetilde{\text{Gen}}$ outputs $\tilde{D} = (D_1, \tilde{D}_2)$ with oracle access to D , $P^{(1)}$, $hash$, g such that with high probability it holds

$$\Pr_{\pi, \rho} [\text{Success}^{P^{(1)}, \tilde{D}}(\pi, \rho) = 1] \geq (\delta + \frac{\varepsilon}{6k}).$$

The running time of FindHash is $\text{poly}(h, v, \frac{1}{\varepsilon}, n)$ with oracle calls and of Gen $\text{poly}(k, \frac{1}{\varepsilon}, n)$ with oracle access. Thus, the overall running time of $\widetilde{\text{Gen}}$ is $\text{poly}(k, \frac{1}{\varepsilon}, h, v, n, t)$ with oracle access. Furthermore, the circuit \tilde{D} asks at most $\frac{6k}{\varepsilon} \log(\frac{6k}{\varepsilon})h$ hint queries and one verification query. Finally, we have $\text{Size}(\tilde{D}) \leq \text{Size}(C) \cdot \frac{6k}{\varepsilon}$. This finishes the proof of Theorem 4.2. \square

4.2 Discussion

Appendix A

Appendix

A.1 Basic Inequalities

Lemma A.1 (Chernoff Bounds) *For independent, identically distributed Bernoulli random variables X_1, \dots, X_n with $X := \sum_{i=1}^n X_i$ with $\Pr[X_i = 1] = p_i$ and $\Pr[X_i = 0] = 1 - p_i$ for all $1 \leq i \leq n$. we have the following inequalities for $0 \leq \delta \leq 1$ and $\mathbb{E}[X] = \sum_{i=1}^n p_i$:*

$$\Pr[X \geq (1 + \delta)\mathbb{E}[X]] \leq e^{-\mathbb{E}[X]\delta^2/3} \quad (\text{A.1})$$

$$\Pr[X \leq (1 - \delta)\mathbb{E}[X]] \leq e^{-\mathbb{E}[X]\delta^2/2} \quad (\text{A.2})$$

$$\Pr[|X - \mathbb{E}[X]| \geq \delta\mathbb{E}[X]] \leq 2e^{-\mathbb{E}[X]\delta^2/3}. \quad (\text{A.3})$$

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