

Crypto - HW 5

Hagai Ben Yehuda, ID num: 305237000
Jonathan Bauch, ID num: 204761233

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1.a

Indeed, suppose we have the signature σ_m of $m \in \{1, \dots, n\}$, then we can obtain the signature of any $k \in \{1, \dots, n\}$ that satisfies $k < m$ by simply setting $\sigma_k = f^{m-k}(\sigma_m)$, then we have that

$$f^k(\sigma_k) = f^k(f^{m-k}(\sigma_m)) = f^m(\sigma_m) = y$$

With the last equality is due to the assumption that σ_m is the correct signature for m . Thus this is not a one time secure signature scheme.

1.b

First we show that f^k is also a permutation, we do so using induction:

For the base case where $k = 1$ this is true from the definition of f . Now assume that f^{k-1} is a permutation, let $x \in \{0, 1\}^n$ then x has a source under f (as f is a permutation), namely there exists $y \in \{0, 1\}^n$ such that $f(y) = x$, from the induction assumption we know that f^{k-1} is a permutation and thus y has a source $z \in \{0, 1\}^n$ such that $f^{k-1}(z) = y$, thus $f(f^{k-1}(z)) = f(y) = x$, thus x has a source under f^k . Since $\{0, 1\}^n$ is finite and since we have showed that f^k is on-to, we have that f is also one to one, Hence f is a permutation. Now assume that f^k is not one-way then there is a polynomial time algorithm A that satisfies:

$$\Pr_{x \leftarrow \{0,1\}^n} (A(f^k(x)) = x) > \epsilon$$

We define an algorithm A' that on input $y = f(x)$ calculates $f^{k-1}(y)$ and feeds it to A . A' is polynomial since A and k are polynomial, also note that:

$$\Pr_{x \leftarrow \{0,1\}^n} (A'(f(x)) = x) = \Pr_{x \leftarrow \{0,1\}^n} (A(f^k(x)) = x) > \epsilon$$

Leading to a contradiction to the assumption that f is a OWP, hence no such A exists and f^k is also a OWP.

1.c

Indeed, assuming that there is some polynomial algorithm A such that for some $m \in \{1, \dots, n\}$, $\sigma_m = f^{n-m}(m)$ and $m' > m$, A satisfies: Indeed, assuming that there is some polynomial algorithm A such that for some $m \in \{1, \dots, n\}$, $\sigma_m = f^{n-m}(m)$ and $m' > m$, A satisfies:

$$\Pr_{x \leftarrow \{0,1\}^n} (A(\sigma_m) = \sigma(m') = f^{n-m'}(x)) > \epsilon$$

We construct a polynomial algorithm A' that inverts f with the same probability, on input $f(w)$ A' will do the following:

- Set $k = m' - m$.
- Set $\sigma_m = f^{k-1}(f(w)) = f^k(w)$.
- Execute $A(\sigma_m)$ and return its result.

Then

$$\begin{aligned}
\Pr_{w \leftarrow \{0,1\}^n} (A'(f(w)) = w) &= \Pr_{x \leftarrow \{0,1\}^n} (A'(f^{n-m}(x)) = w) \\
&= \Pr_{x \leftarrow \{0,1\}^n} (A(\sigma_m) = w) \\
&= \Pr_{x \leftarrow \{0,1\}^n} (A(\sigma_m) = f^{-k}(\sigma_m)) \\
&= \Pr_{x \leftarrow \{0,1\}^n} (A(\sigma_m) = \sigma_{m'}) > \epsilon
\end{aligned}$$

The first equality is due to the fact that given a random w , the probability for any $x \leftarrow \{0,1\}^n$ to be its k 'th source is equal for every x as we assume that f was chosen uniformly from the random permutation functions. This is obviously a contradiction to the assumption that f is a OWF, showing no such algorithm A exists.

1.d

////////// need to do this!!!!!!!!!!!! //////////

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Let \tilde{f} be some one way function, and define f such that $f(xl) = f(x0)$ (with l being a single bit). Suppose that f is not a one way function, then there is an algorithm A such that

$$\Pr_{x \leftarrow \{0,1\}^n} (A(f(x)) = x) > \epsilon$$

Construct an algorithm A' that on input $\tilde{f}(x)$ executes A to receive yl and returns $y0$. Then we have:

$$\begin{aligned}
\Pr_{x \leftarrow \{0,1\}^n} (A'(\tilde{f}(x)) = x) &\geq \frac{1}{2} \Pr_{x \leftarrow \{0,1\}^{n-1}} (A'(\tilde{f}(x0)) = x0) \\
&= \frac{1}{2} \Pr_{x \leftarrow \{0,1\}^n} (A(f(x)) = x) \\
&> \frac{\epsilon}{2}
\end{aligned}$$

The equality (in the second line) is correct because A' will produce a valid result whenever A is able to find the source of a message in $\{0,1\}^n$ because if it finds a source then we know that zeroing the last bit of the result will be the source of the original function. Now that we have constructed f and have shown that it is a one way function, then if we use Lamports scheme with f and have a valid signature $(x_{m_1,1}, \dots, x_{m_n,n})$ of the message (m_1, \dots, m_n) then we know that if $x_{m_1,1} = xl$ then $(x\bar{l}, \dots, x_{m_n,n})$ is also a signature for the same message from the construction of f , since $f(xl) = f(x\bar{l})$, showing that using f Lamports scheme is not a strong one time signature scheme.

3

Assume we are given such an algorithm A that breaks (ϵ, t) -existential-unforgeability of the scheme, we construct A' that breaks $(\frac{\epsilon}{t+1}, 1)$ -existential-unforgeability of the underlying one-time scheme as follows:

- Draw $r \in \{1, \dots, t\}$ uniformly.
- If $r > 1$
 - Execute A while simulating the oracle for the first $r - 1$ messages (generating our own keys).
 - For the $r - 1$ 'th message A asks for, m_{r-1} , sign (m_{r-1}, pk) with pk being the oracle's public key (which we are trying to attack) and send the result to A .
 - For the next message A asks for, m_r , generate the keys (pk_r, sk_r) , ask the oracle to sign (pk_r, m_r)
- If $r = 1$
 - Publish that our public key is pk (the public key that belongs to the oracle we are trying to break).
 - For the first message A asks for, m_1 generate keys (pk_2, sk_2) and ask the oracle to sign (pk_2, m_1)
- Continue the process of simulating the oracle normally.
- A returns a trust chain of messages that is different from the one we have provided to it. (We can assume that A 's output is of the form $((m_1, pk_2, \sigma_1), \dots, (m_n, pk_{n+1}, \sigma_n))$) If the chain A returned differs from the chain we have created in the $r + 1$ 'th index (even if $r = t$ that mean there is an extra message in the chain returned by A), return $(m_{r+1}, pk_{r+2}, \sigma_{r+1})$.

First we note that A' runs roughly in the same time as A since it only runs A , generates keys and queries the oracle once. As stated, note that A' only queries the oracle once so we only need to prove that A' can forge a new message with probability $> \frac{\epsilon}{t+1}$. Note that A is able to forge a signature with probability $> \epsilon$ which mean that with probability $> \epsilon$ at the end of A we will have a valid trust chain which is different from the one we have supplied A with. note that the chain must differ at atleast one place (from the first to one after the last) thus the probability that the chain remains unchanged up to the r 'th index and then there is a difference in the $r + 1$ 'th place is $\frac{1}{t+1}$ (because the difference must start somewhere and we have chosen r uniformly) in which case if A indeed forged a valid chain, σ_{r+1} is a signature of (m_{r+1}, pk_{r+2}) signed using sk_r (because that is the private key matching the public key signed by the previous message, which we assume was unchanged) which is the oracles private key and (m_{r+1}, pk_{r+2}) is a previously unsigned message. Thus in this case (which happens with probability $> \frac{\epsilon}{t+1}$) A' breaks the underling signature scheme, showing it is not $(\frac{\epsilon}{t+1}, 1)$ -existential- unforgeability as requested.

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4.a

Indeed, assume that $y \notin QR$ then there are two options:

- If in the first step P sent z such that $z = r^2$ (i.e. z is a quadratic residue) then with probability $\frac{1}{2}$ - V will choose $b = 1$ in step 2 in which case there is no a_1 that satisfies $a_1^2 = zy$ because if there was such an a_1 then $y = (a_1 z^{-1})^2$ which contradicts the assumption that $y \notin QR$. Hence with probability at least $\frac{1}{2}$ - V rejects.
- If in the first step P sends z such that z is not a quadratic residue then if P chooses $b = 0$ in the second step then P cannot send r in the third step (because no such r exists by definition, regardless of P 's computational limitations). Hence with probability at least $\frac{1}{2}$ - V rejects.

Thus regardless of what P sends in the first step there is a probability of at least $\frac{1}{2}$ that V will reject if $y \notin QR$ (meaning it will accept with probability $\leq \frac{1}{2}$). As requested.

4.b

Indeed, we show a simulator that operates on input (y, b) :

- If $b = 0$:
 - Draw $r \in \mathbb{Z}_N^*$ and set $\tilde{z} = r^2$
 - Set $\tilde{a}_0 = r$.
- If $b = 1$
 - Draw $\tilde{a}_1 \in \mathbb{Z}_N^*$ uniformly.
 - Set $\tilde{z} = \tilde{a}_1^2 y^{-1} \pmod{N}$

Then the output of this simulator is Indeed indistinguishable from a true execution of the protocol since if $b = 0$ then we send values that would have made the verifier accept (we send a quadratic residue and then its root) and if $b = 1$ we send $\tilde{z} = \tilde{a}_1^2 y^{-1} \pmod{N}$ and thus $\tilde{z}y = \tilde{a}_1^2 \pmod{N}$, hence the verifier would accept. Since by assumption quadratic residues are polynomially indistinguishable from non quadratic residues \tilde{z} is indistinguishable from a uniformly drawn quadratic residue and thus the view is indistinguishable from a successful transaction between the prover and verifier.

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6.a

6.b