CS255: Cryptography and Computer Security **Winter 2006**

Assignment #3

Due: Monday, Mar. 6th, 2006.

Problem 1 Conference key setup.

Parties $A_1, \ldots, A_n$ and B wish to generate a secret conference key. All parties should know the conference key, but an eavesdropper should not be able to obtain any information about the key. They decide to use the following variant of Diffie-Hellman: there is a public prime p and a public element $g \in \mathbb{Z}_p^*$ of order q for some large prime q dividing $p-1$. User B picks a secret random $b \in [1, q-1]$ and computes $y = g^b \bmod p$. Each party A_i picks a secret random $a_i \in [1, q-1]$ and computes $x_i = g^{a_i} \bmod p$. User A_i sends x_i to B. User B responds to party i by sending $z_i = x_i^b \bmod p$.

a. Show that party i given z_i (and a_i) can determine y.

b. Explain why (a hash of) y can be securely used as the conference key. Namely, explain why at the end of the protocol all parties $A_1, \ldots, A_n$ and B know y and give a brief informal explanation why an eavesdropper cannot determine y.

c. Prove part (b). Namely, show that if there exists an efficient algorithm $\mathcal{A}$ that given the public values in the above protocol, outputs y, then there also exists an efficient algorithm $\mathcal{B}$ that breaks the Computational Diffie-Hellman assumption (using p and g as the public values). Use algorithm $\mathcal{A}$ as a subroutine in your algorithm $\mathcal{B}$. Note that algorithm $\mathcal{B}$ takes $g^a \bmod p$ and $g^b \bmod p$ as input and should output $g^{ab} \bmod p$.

Problem 2 Commitment schemes. A commitment scheme enables Alice to commit a value x to Bob. The scheme is *secure* if the commitment does not reveal to Bob any information about the committed value x. At a later time Alice may *open* the commitment and convince Bob that the committed value is x. The commitment is *binding* if Alice cannot convince Bob that the committed value is some $x' \neq x$. Here is an example commitment scheme:

Public values: (1) a 1024 bit prime p, and (2) two elements g and h of $\mathbb{Z}_p^*$ of prime order q.

Commitment: To commit to an integer $x \in [0, q-1]$ Alice does the following: (1) she picks a random $r \in [0, q-1]$, (2) she computes $b = g^x \cdot h^r \bmod p$, and (3) she sends b to Bob as her commitment to x.

Open: To open the commitment Alice sends (x, r) to Bob. Bob verifies that $b = g^x \cdot h^r \bmod p$.

Show that this scheme is secure and binding.

a. To prove security show that b does not reveal any information to Bob about x. In other words, show that given b, the committed value can be any integer x' in $[0, q-1]$.
Hint: show that for any x' there exists a unique $r' \in [0, q-1]$ so that $b = g^{x'}h^{r'}$.

b. To prove the binding property show that if Alice can open the commitment as (x', r') where $x \neq x'$ then Alice can compute the discrete log of h base g. In other words, show that if Alice can find an (x', r') such that $b = g^{x'}h^{r'} \bmod p$ then she can find the discrete log of h base g. Recall that Alice also knows the (x, r) used to create b.

Problem 3. Incremental hashing. Let p be a prime and let $g \in \mathbb{Z}_p^*$ be an element of prime order q. We let G denote the group generated by g and we let I denote the set of integers $\{1, \ldots, q\}$. Fix n values $g_1, \ldots, g_n \in G$ and define the hash function $H : I^n \to G$ by

$$H(x_1, \ldots, x_n) = g_1^{x_1} g_2^{x_2} \cdots g_n^{x_n}$$

a. Show that H is collision resistant assuming discrete-log in G is intractable. That is, show that an attacker capable of finding a collision for H for a *random* $g_1, \ldots, g_n \in G$ can be used to compute discrete-log in G.
Hint: given a pair $g, h \in G$ your goal is to find an $\alpha \in \mathbb{Z}$ such that $g^\alpha = h$. Choose $g_1, \ldots, g_n \in G$ so that a collision on the resulting H will reveal α.

b. Let M be a message in I^n. Suppose user Alice already computed the hash of M, namely $H(m)$. Now, Alice changes only one coordinate of M to obtain a new message M'. Show that Alice can quickly compute $H(M')$ from $H(M)$ in time that is independent of the length of M.
You have just shown that after making a small change to a message there is no need to rehash the entire message. Collision resistant hash functions of this type are said to support *incremental hashing*.

Problem 4 Let's explore why in the RSA public key system each person has to be assigned a different modulus $N = pq$. Suppose we try to use the same modulus $N = pq$ for everyone. Each person is assigned a public exponent e_i and a private exponent d_i such that $e_i \cdot d_i = 1 \bmod \varphi(N)$. At first this appears to work fine: to encrypt a message to Bob, Alice computes $C = M^{e_{bob}}$ and sends C to Bob. An eavesdropper Eve, not knowing d_{bob} appears to be unable to decrypt C. Let's show that using e_{eve} and d_{eve} Eve can very easily decrypt C.

a. Show that given e_{eve} and d_{eve} Eve can obtain a multiple of $\varphi(N)$.

b. Show that given an integer K which is a multiple of $\varphi(N)$ Eve can factor the modulus N. Deduce that Eve can decrypt any RSA ciphertext encrypted using the modulus N intended for Alice or Bob.
Hint: Consider the sequence $g^K, g^{K/2}, g^{K/4}, \ldots g^{K/\tau(N)} \bmod N$ where g is random in $\mathbb{Z}_N$ and $\tau(N)$ is the largest power of 2 dividing K. Use the the left most element in this sequence which is not equal to 1 mod N.

Problem 5 Recall that a simple RSA signature $S = H(M)^d \bmod N$ is computed by first computing $S_1 = H(M)^d \bmod p$ and $S_2 = H(M)^d \bmod q$. The signature S is then found by combining S_1 and S_2 using the Chinese Remainder Theorem (CRT). Now, suppose a Certificate Authority (CA) is about to sign a certain certificate C. While the CA is computing $S_1 = H(C)^d \bmod p$, a glitch on the CA's machine causes it to produce the wrong value $\tilde{S}_1$ which is not equal to S_1. The CA computes $S_2 = H(C)^d \bmod q$ correctly. Clearly the resulting signature $\tilde{S}$ is invalid. The CA then proceeds to publish the newly generated certificate with the invalid signature $\tilde{S}$.

a. Show that any person who obtains the certificate C along with the invalid signature $\tilde{S}$ is able to factor the CA's modulus.
Hint: Use the fact that $\tilde{S}^e = H(C) \bmod q$. Here e is the public verification exponent.

b. Suggest some method by which the CA can defend itself against this danger.

Problem 6. Offline signatures. One approach to speeding up signature generation is to perform the bulk of the work offline, before the message to sign is known. Then, once the message M is given, generating the signature on M should be very fast. Our goal is to design a signature system with this property.

a. Does the RSA Full-Domain-Hash (FDH) signature system enable this form of offline signatures? In other words, can we substantially speed-up RSA signature generation if we are allowed to perform offline computation before the message M is given? It might help to explicitly write out the signing algorithm in RSA FDH.

b. Our goal is to show that any signature system can be converted into a signature where the bulk of the signing work can be done offline. Let $(\mathsf{KeyGen}, \mathsf{Sign}, \mathsf{Verify})$ be a signature system (such as RSA FDH) and let G be a group of order q where discrete log is hard. Consider the following modified signature system $(\mathsf{KeyGen}', \mathsf{Sign}', \mathsf{Verify}')$:

- The KeyGen' algorithm runs algorithm KeyGen to obtain a signing key SK and verification VK. It also picks a random group element $g \in G$ and sets $h = g^\alpha$ for some random $\alpha \in \{1, \ldots, q\}$. It outputs the verification key $VK' = (VK, g, h)$ and the signing key $SK' = (VK', SK, \alpha)$.
- The $\mathsf{Sign}'(SK', M)$ algorithm first picks a random $r \in \{1, \ldots, q\}$, computes $m = g^M h^r \in G$, and then runs $\mathsf{Sign}(SK, m)$ to obtain a signature σ. It outputs the signature $\sigma' = (\sigma, r)$.
- The $\mathsf{Verify}'(VK', M, \sigma')$ algorithm, where $\sigma' = (\sigma, r)$, computes $m = g^M h^r \in G$ and outputs the result of $\mathsf{Verify}(VK, m, \sigma)$.

show that the bulk of the work in the Sign' algorithm can be done before the message is given.
Hint: First, observe that since α is part of the secret key SK', the signer can compute $m = g^M h^r$ as $m = g^{M+\alpha r}$. Now, offline, try running $\mathsf{Sign}(SK, m)$ on

a message $m = g^s$ for a randomly chosen $s \in \{1, \ldots, q\}$. Let σ be the resulting signature. Then, once the message M is given, show that the signer can easily convert σ into a valid signature σ' for M using only one addition and one multiplication modulo q.

c. (extra credit) Prove that the modified signature scheme is secure. In other words, show that an existential forger under a chosen message attack on the modified scheme gives an existential forger on the underlying scheme. You may use the fact (proved in problem 3) that $H(M, r) = g^M h^r$ is a collision resistant hash function and hence the adversary cannot find collisions for it.