

1 The Diffie-Hellman Key Exchange System

1.1 Mechanism

Goal: Alice and Bob want to exchange a shared “secret key” via an insecure channel. (Incidentally, the shared secret key is usually used for encryption and decryption with a symmetric cipher.)

- Alice and Bob choose and publish a large prime p and an integer g with large prime order in \mathbb{F}_p^* .
- Alice chooses a secret $a \in \mathbb{Z}$, computes $A := g^a \bmod p$, and sends $A \in \mathbb{F}_p^*$ to Bob via the insecure channel.
- Bob chooses a secret $b \in \mathbb{Z}$, computes $B := g^b \bmod p$, and sends $B \in \mathbb{F}_p^*$ to Alice via the insecure channel.
- Bob computes $A^b = (g^a)^b = g^{ab} \in \mathbb{F}_p^*$.
- Alice computes $B^a = (g^b)^a = g^{ab} \in \mathbb{F}_p^*$.
- Alice and Bob have both arrived at $g^{ab} \in \mathbb{F}_p^*$, which is to be their shared secret.

1.2 Security of the Diffie-Hellman key exchange system

- The presumed security of the Diffie-Hellman key exchange system is based on the presumed difficulty in solving the

The Diffie-Hellman Problem:

Let p be a prime, and $g \in \mathbb{F}_p^*$. Given $A, B \in \langle g \rangle \subset \mathbb{F}_p^*$, find $g^{ab} \in \mathbb{F}_p^*$, where $a, b \in \mathbb{Z}$ are determined by $A = g^a$ and $B = g^b$.

Note that the solution of the Diffie-Hellman problem does NOT require the knowledge of a and b , but only that of $g^{ab} \in \langle g \rangle$.

- It is clear that an efficient algorithm for the Discrete Logarithm Problem will lead to an efficient algorithm for the Diffie-Hellman Problem. An efficient algorithm of the Discrete Logarithm Problem will break the Diffie-Hellman key exchange system.

The Discrete Logarithm Problem:

Let p be a prime, and $g \in \mathbb{F}_p^*$ be a primitive element, i.e. $\mathbb{F}_p^* = \langle g \rangle$. Given any $h \in \mathbb{F}_p^*$, find $x \in \mathbb{Z}$ such that $h = g^x$.

2 The ElGamal Public Key Cryptosystem

2.1 Mechanism

Goal: Bob wants to send Alice an encrypted message via an insecure channel.

- Alice chooses and publishes a large prime p and an integer g with large prime order in \mathbb{F}_p^* .
- Alice chooses a secret $a \in \mathbb{Z}$, with $1 \leq a \leq p - 1$, computes $A := g^a \bmod p$, and publishes $A \in \mathbb{F}_p^*$ as her public key.
- Bob chooses a plaintext $m \in \mathbb{F}_p^*$, and a secret random ephemeral key $k \in \mathbb{Z}$, with $1 \leq k \leq p - 1$. Bob computes $B_1 := g^k \bmod p$, and $B_2 := m \cdot A^k \in \mathbb{F}_p^*$. Bob's ciphertext is (B_1, B_2) and Bob sends it to Alice via the insecure channel.
- Alice decrypts Bob's ciphertext (B_1, B_2) by computing $B_2 \cdot (B_1^a)^{-1} \in \mathbb{F}_p^*$. Note that

$$B_2 \cdot (B_1^a)^{-1} \equiv (m \cdot A^k) \cdot (g^{ak})^{-1} \equiv m \cdot (g^a)^k \cdot (g^{ak})^{-1} \equiv m \cdot (g^{ak}) \cdot (g^{ak})^{-1} \equiv m \bmod p$$

2.2 Security of the ElGamal public key cryptosystem

- The ElGamal public key encryption system is at least as secure as the difficulty of Diffie-Hellman problem, in the sense that an ElGamal oracle (efficient solver of the ElGamal cryptosystem problem) can be used to efficiently solve the Diffie-Hellman problem.

3 The RSA Public Key Cryptosystem

3.1 Mechanism

Goal: Alice wants to send Bob an encrypted message through an insecure channel.

- Bob chooses his public key $(n, e) \in \mathbb{N}^2$ and private key $d \in \mathbb{N}$. Bob publishes his public key.
 - $n \in \mathbb{N}$ is called the modulus, with $n = pq$, where p and q are large distinct prime numbers. Note that Bob publishes n but keeps p and q secret.
 - $e \in \mathbb{N}$ is the called encryption exponent, and satisfies $\gcd(e, (p-1)(q-1)) = 1$.
 - $d \in \mathbb{N}$ is the called decryption exponent, and is determined by e and $n = pq$ via $d = e^{-1} \in \mathbb{Z}_{(p-1)(q-1)}$. Note that $e^{-1} \in \mathbb{Z}_{(p-1)(q-1)}$ exists since $\gcd(e, (p-1)(q-1)) = 1$.
- Alice
 - chooses plaintext $m \in \mathbb{Z}_n = \mathbb{Z}/n\mathbb{Z}$.
 - encrypts her plaintext m using Bob's public key (n, e) by **raising $m \in \mathbb{Z}_n$ to the e^{th} power**. In other words, Alice computes her ciphertext $c = m^e \in \mathbb{Z}_n$.
 - sends to Bob through the insecure channel the ciphertext $c \in \mathbb{Z}_n$.
- Bob decrypts the ciphertext $c \in \mathbb{Z}_n$ from Alice by **taking the e^{th} root** of c in \mathbb{Z}_n using his private key $d \in \mathbb{N}$ as follows:

$$c^d = (m^e)^d = m^{ed} = m^{1+k(p-1)(q-1)} = m \cdot (m^{(p-1)(q-1)})^k = m \cdot (1)^k = m \in \mathbb{Z}_n$$

- The second last equality follows from $m^{(p-1)(q-1)} \equiv 1 \pmod{n}$, which follows immediately from Euler's Theorem. It can also be justified with Fermat's Little Theorem as follows:

$$\text{Fermat's Little Theorem} \implies \begin{cases} m^{(p-1)(q-1)} = (m^{p-1})^{q-1} \equiv 1 \pmod{p}, \text{ and} \\ m^{(p-1)(q-1)} = (m^{q-1})^{p-1} \equiv 1 \pmod{q}. \end{cases}$$

Hence, $m^{(p-1)(q-1)} - 1$ is divisible by both p and q , and hence also by $pq = n$ (since p and q are distinct primes). Thus, $m^{(p-1)(q-1)} \equiv 1 \pmod{n = pq}$.

3.2 Comments

- One-way function (easy): Exponentiation in \mathbb{Z}_n .
 - Repeating Squaring Algorithm
- (Difficult) inverse function: Taking roots in \mathbb{Z}_n , for $n = pq$, where p and q are large distinct prime numbers.
- Trapdoor: If the factorization of $n = pq$ is known, then we can convert the inverse function (taking roots in \mathbb{Z}_n , which is slow) to an exponentiation in \mathbb{Z}_n , which is fast.

3.3 How to find large prime numbers?

- Generate a large N -bit (say $N = 1024$) random number x , i.e. $2^{N-1} < x < 2^N$. Use an efficient primality test to check whether x is prime. If so, we are done. If not, repeat until we succeed.
- The Prime Number Theorem (from Analytic Number Theory) gives an estimate of how many times we need to try before succeeding. The Prime Number Theorem states that

$$\lim_{x \rightarrow \infty} \frac{\pi(x)/x}{1/\ln(x)} = 1,$$

where $\pi(x)$ is the number of prime numbers less than or equal to x . Hence, it implies that, for large values of N , the probability that a randomly selected integer $x \in (2^{N-1}, 2^N)$ is prime is approximately

$$\frac{1}{\ln(2^N)}$$

Conversely, this implies that, on average, out of every $\ln(2^N) = N \cdot \ln(2) \approx 0.693 \cdot N$ randomly and independently selected integers from $(2^{N-1}, 2^N)$, one of them will be a prime number. For example, if $N = 1024$, then $0.693 \cdot N \approx 709.78$; in other words, if we are selecting random integers from $(2^{1023}, 2^{1024})$, then on average, we expect repeating approximately 710 times before we succeed in selecting a prime number. Note that $2^{1023} = 10^{1023 \times \log_{10}(2)} \approx 10^{1023 \times 0.301} \approx 10^{307.95}$.

- The Miller-Rabin Primality test
 - **Proposition** Let p be an odd prime and write $p - 1 = 2^k t$, where t is odd. Then, for each $a \in \mathbb{Z}$ with $p \nmid a$, one of the following is true:
 - $a^t \equiv 1 \pmod{p}$, or
 - One of $a^t, a^{2t}, a^{4t}, \dots, a^{2^{k-1}t}$ is congruent to $-1 \pmod{p}$.
 - **Corollary** Let $n \in \mathbb{Z}$ be an odd number, with $n - 1 = 2^k t$, t being odd. Then, n is composite, if any of the following is true:
 - There exists $a \in \mathbb{Z}$ such that $\gcd(a, n) > 1$.
 - There exists $a \in \mathbb{Z}$ such that $\gcd(a, n) = 1$, and $a^t \not\equiv 1 \pmod{n}$, and $a^{2^i t} \not\equiv -1 \pmod{n}$, for each $i = 0, 1, 2, \dots, k - 1$.
 - **Proposition** Let n be an odd composite number. Then, at least 75% of integers between 1 and $n - 1$ are Miller-Rabin witnesses for n .

3.4 Factorization algorithms

- Pollard's $p - 1$ factorization algorithm

This method “probably” works for producing a non-trivial factor for composite $n \in \mathbb{N}$ admitting a prime factor p such that $p - 1$ is a product of small primes.

- **Proposition:** Let $n = pq$, where p and q are distinct prime numbers. Then the following two statements hold:

- For each $L \in \mathbb{N}$ and a with $p \nmid a$, we have the following implication:

$$(p-1) \mid L \implies L = (p-1)K \implies a^L = \left(a^{(p-1)}\right)^K \equiv 1 \pmod{p} \implies p \mid (a^L - 1)$$

- For any $a \in \mathbb{N}$ with $p \nmid a$ and $q \nmid a$, and any $L \in \mathbb{N}$,

$$\left. \begin{array}{l} p \mid (a^L - 1) \\ q \nmid (a^L - 1) \end{array} \right\} \implies p = \gcd(a^L - 1, n)$$

– **Key observations:**

- If $p-1$ is a product of small primes, then $(p-1) \mid N!$, for some not-too-large N .
- If $(q-1) \nmid N!$, then $q \nmid (a^{N!} - 1)$ is “probably” true.
- If $p-1$ is a product of small primes, and $q-1$ is NOT so, then computing $\gcd(a^{k!} - 1, n)$, for $k = 2, 3, \dots$, will “probably” yield p as a non-trivial factor of n .

- Factorization via difference of squares

– **Key observations:**

Suppose we know that $n \in \mathbb{N}$ is odd and composite. We want to find a non-trivial factor of n .

- If we can find $a, b \in \mathbb{N}$ such that n is the difference of their squares, i.e. $n = a^2 - b^2 = (a-b)(a+b)$, then computing $\gcd(a-b, n)$ will yield a non-trivial factor of n .
- Conversely, suppose $n = cd$. Since n is odd, both c and d must also be odd. Hence, $a := \frac{1}{2}(c+d) \in \mathbb{Z}$ and $b := \frac{1}{2}(c-d) \in \mathbb{Z}$. And, $a^2 - b^2 = \dots = cd = n$. In other words, every composite odd integer can be written as the difference of two squares.
- If some multiple kn is a difference of squares, i.e. $kn = a^2 - b^2 = (a-b)(a+b)$, then computing $\gcd(a-b, n)$ will “probably” yield a non-trivial factor of n , since it should be unlikely that n divides $a-b$.
- **In summary, if we could find $a, b \in \mathbb{Z}$ such that $a^2 \equiv b^2 \pmod{n}$, then computing $\gcd(a-b, n)$ will probably yield a non-trivial factor of n .**

– **Outline of general procedure:**

- (1) **Find B -smooth perfect squares in \mathbb{Z}_n .** Find many $a_1, a_2, \dots, a_r \in \mathbb{Z}$ such that every prime factor of $c_i \equiv a_i^2 \pmod{n}$ is less than or equal to B .
- (2) Find sub-collections $c_{i_1}, c_{i_2}, \dots, c_{i_s}$ such that $c_{i_1}c_{i_2} \dots c_{i_s} \equiv b^2 \pmod{n}$ are perfect squares in \mathbb{Z}_n .
- (3) Let $a := a_{i_1}a_{i_2} \dots a_{i_s} \pmod{n}$. Then, computing $\gcd(a-b, n)$ will probably yield a non-trivial factor of n .

– Comments on the general procedure:

- Step (3) can be performed efficiently using the Euclidean Algorithm.
- Step (2) is equivalent to solving a homogeneous (sparse) system of linear equations over \mathbb{F}_2 .
- The main challenge in difference-of-squares factorization is Step (1), namely, given $n \in \mathbb{Z}$, finding enough B -smooth perfect squares in \mathbb{Z}_n .

4 Methods for finding smooth numbers

4.1 The Quadratic Sieve

Problem

- Let $N, B \in \mathbb{N}$ be given.
- Find a large number of integers of the form $a^2 \bmod N$ that are B -smooth.

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5 Algorithms for solving the Discrete Logarithm Problem

The Discrete Logarithm Problem

- Let G be a finite cyclic group of order $N \geq 2$, and let $g \in G$ be a generator of G .
- Let $h \in G \setminus \{e\}$ be given.
- Find $x \in \mathbb{N}$, $1 \leq x < N$, such that $g^x = h$.

5.1 Shank's Baby-Step-Giant-Step Algorithm

Theory:

- Let $n := 1 + \lfloor \sqrt{N} \rfloor$, where $\lfloor \sqrt{N} \rfloor$ is the round-down of \sqrt{N} . Note, trivially, that $n^2 > N$.
- Define:

$$\begin{aligned} \mathcal{S}_1 &:= \{ g^i \mid i = 0, 1, 2, \dots, n \} = \{ e, g, g^2, g^3, \dots, g^n \} \\ \mathcal{S}_2 &:= \{ h \cdot (g^{-n})^i \mid i = 0, 1, 2, \dots, n \} = \{ h, h \cdot (g^{-n}), h \cdot (g^{-n})^2, h \cdot (g^{-n})^3, \dots, h \cdot (g^{-n})^n \} \end{aligned}$$

- Then, it turns out that $\mathcal{S}_1 \cap \mathcal{S}_2 \neq \emptyset$.
- For any $i, j \in \{0, 1, 2, \dots, n\}$ such that $g^i = h \cdot (g^{-n})^j \in \mathcal{S}_1 \cap \mathcal{S}_2$, the positive integer $x := i + jn \in \mathbb{Z}$ is a solution to the Discrete Logarithm Problem.

Outline of proof:

- $h \in \langle g \rangle \setminus \{e\} \implies$ there exists $x \in \mathbb{N}$, $1 \leq x < N$, such that $g^x = h$.
- Then, $x = nq + r$, with $0 \leq r < n$ and $q \geq 0$. We claim that the quotient q in fact satisfies: $q < n$. Indeed, $x < N$ implies

$$q = \frac{x - r}{n} \leq \frac{N}{n} < \frac{n^2}{n} = n.$$

- Hence, $g^x = h$ can be rewritten as follows:

$$g^x = h \iff g^{nq+r} = h \iff g^r = h \cdot (g^{-n})^q \in \mathcal{S}_1 \cap \mathcal{S}_2, \quad \text{for some } 0 \leq r, q < n$$

Pseudocode:

- (1) Let $n := 1 + \lfloor \sqrt{N} \rfloor$, where $\lfloor \sqrt{N} \rfloor$ is the round-down of \sqrt{N} .
- (2) Generate $\mathcal{S}_1 = \{ g^i \mid i = 0, 1, 2, \dots, n \} = \{ e, g, g^2, g^3, \dots, g^n \}$.
- (3) Form the hash table \mathcal{T} :

key g^i	e	g	g^2	g^3	\dots	g^n
value i	0	1	2	3	\dots	n

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(4) gn := g^{-n};
    temp := h;
    for j = 0, 1, 2, ...
        if (T{temp} exists) {
            i := T{temp};
            return(i + n*j);
        } else {
            temp := temp * gn;
        }
    end

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Complexity: $\mathcal{O}(\sqrt{N})$

- The generation of \mathcal{S}_1 takes approximately n multiplications.
- The generation of the hash table takes $\mathcal{O}(1)$ computation steps.
- The loop takes $\mathcal{O}(n)$ computational steps.
- Hence, the whole algorithm takes $\mathcal{O}(n + n + 1) = \mathcal{O}(n) = \mathcal{O}(\sqrt{N})$ computational steps.

5.2 The Pohlig-Hellman Algorithm

Summary:

- The Pohlig-Hellman Algorithm renders efficiently solvable the Discrete Logarithm Problem (DLP) in any finite cyclic group G whose order $|G|$ is a product of powers of small primes.
- This is achieved with two results of Pohlig-Hellman:
 - The DLP in a finite cyclic group G with $|G| = N = p_1^{\alpha_1} p_2^{\alpha_2} \cdots p_t^{\alpha_t}$ can be “converted” to t distinct associated DLP’s in groups of orders $p_1^{\alpha_1}, \dots, p_t^{\alpha_t}$, respectively. Here, “converted” means that the solutions of the associated DLP’s can be “combined”, via the Chinese Remainder Theorem, to give the solution of the original DLP.
 - The DLP in a finite cyclic group of prime power order p^α can be “converted” to a series of DLP’s, each in a group of power p .
 - This last set of DLP’s in groups of prime power must then be solved by methods other than Pohlig-Hellman, for example, either by brute force, or by Shank’s Baby-Step-Giant-Step method. Thus, if each of the primer orders involved is relatively small (*e.g.*, efficiently solvable by brute force or Baby-Step-Giant-Step), then the original DLP can be efficiently solved.

Theory:

- The assumption that $h \in G = \langle g \rangle$ implies the solvability of $g^x = h$ for $x \in \{0, 1, 2, \dots, N-1\}$.
- Let $N = p_1^{\alpha_1} p_2^{\alpha_2} p_3^{\alpha_3} \cdots p_t^{\alpha_t}$ be the prime factorization of N . Then, it turns out that the following associated discrete logarithm problems (in finite cyclic groups with prime power order):

$$g_i^{y_i} = h_i, \quad \text{where } g_i := g^{N/p_i^{\alpha_i}}, \quad h_i := h^{N/p_i^{\alpha_i}}, \quad \text{and } i = 1, 2, \dots, t,$$

are also solvable for

$$y_i \in \{0, 1, 2, \dots, p_i^{\alpha_i} - 1\}.$$

If we also have their solutions y_i , then we can compute x using the Chinese Remainder Theorem, namely, by solving the following simultaneous congruences for $x \in \{0, 1, 2, \dots, N-1\}$:

$$x \equiv y_i \pmod{p_i^{\alpha_i}}, \quad \text{for } i = 1, 2, \dots, t.$$

Outline of proof:

The solvability of $g^x = h$ implies the solvability of the auxiliary discrete logarithm problems $g_i^{y_i} = h_i$ for $y_i \in \{0, 1, 2, \dots, p_i^{\alpha_i} - 1\}$, for each $i = 1, 2, \dots, t$, since

$$g^x = h \implies (g^x)^{N/p_i^{\alpha_i}} = (h)^{N/p_i^{\alpha_i}} \implies \left(g^{N/p_i^{\alpha_i}}\right)^x = (h)^{N/p_i^{\alpha_i}} \implies (g_i)^x = h_i \implies (g_i)^{y_i} = h_i,$$

where $y_i := x \pmod{p_i^{\alpha_i}}$. Here, note that $\text{ord}(g_i) = p_i^{\alpha_i}$, for each $i = 1, 2, \dots, t$. Next, observe that, for each $i = 1, 2, \dots, t$,

$$\begin{aligned} (g^{-x} \cdot h)^{N/p_i^{\alpha_i}} &= (g^{-x})^{N/p_i^{\alpha_i}} \cdot (h)^{N/p_i^{\alpha_i}} = \left(g^{N/p_i^{\alpha_i}}\right)^{-x} \cdot (h)^{N/p_i^{\alpha_i}} \\ &= (g_i)^{-x} \cdot h_i \\ &= (g_i)^{-y_i} \cdot h_i, \quad \text{since } x \equiv y_i \pmod{p_i^{\alpha_i}}, \text{ and } \text{ord}(g_i) = p_i^{\alpha_i} \\ &= 1_G, \quad \text{since } g_i^{y_i} = h_i \end{aligned}$$

Hence, $\text{ord}(g^{-x} \cdot h)$ divides $\frac{N}{p_i^{\alpha_i}}$, for each $i = 1, 2, \dots, t$, which in turn implies that $\text{ord}(g^{-x} \cdot h)$ divides $\gcd\left(\frac{N}{p_1^{\alpha_1}}, \dots, \frac{N}{p_t^{\alpha_t}}\right)$. However, $\frac{N}{p_1^{\alpha_1}}, \dots, \frac{N}{p_t^{\alpha_t}}$ have no non-trivial common divisors, and hence $\gcd\left(\frac{N}{p_1^{\alpha_1}}, \dots, \frac{N}{p_t^{\alpha_t}}\right) = 1$. Thus, $\text{ord}(g^x \cdot h) = 1$; hence, $g^x = h$, as desired.

- The discrete logarithm problem in a finite cyclic group of prime power order p^α can be “converted” to a series of discrete logarithm problems in finite cyclic groups, each of order p .

Outline of proof:

Let $G = \langle g \rangle$ be a finite cyclic group, with $|G| = p^\alpha$, where $p \in \mathbb{N}$ is a prime number, $\alpha \in \mathbb{N}$, and $g \in G$ is a generator of G . Given $h \in G$, we wish to find $x \in \{0, 1, 2, \dots, p^\alpha - 1\}$ such that $g^x = h$.

We make two key observations:

- $\text{ord}(g^{p^{\alpha-1}}) = p$, and
- we may express $x \in \{0, 1, 2, \dots, p^\alpha - 1\}$ as a linear combination of powers of p :

$$x = x_0 + x_1 p + x_2 p^2 + x_3 p^3 + \dots + x_{\alpha-1} p^{\alpha-1}, \quad \text{where } x_i \in \{0, 1, 2, \dots, p-1\}.$$

Assuming we have an “oracle” that can solve the discrete logarithm problem in a finite cyclic group of order p , we can then sequentially solve for $x_0, x_1, x_2, \dots, x_{\alpha-1}$ as solutions to discrete logarithm problems in groups of order p , starting with x_0 .

Raising both sides of $g^x = h$ to the power of $p^{\alpha-1}$ yields:

$$\begin{aligned} \left(g^{x_0 + \sum_{i=1}^{t-1} x_i p^i}\right)^{p^{\alpha-1}} &= (h)^{p^{\alpha-1}} \implies (g^{x_0})^{p^{\alpha-1}} \cdot \left(g^{\sum_{i=1}^{t-1} x_i p^i}\right)^{p^{\alpha-1}} = (h)^{p^{\alpha-1}} \\ &\implies \left(g^{p^{\alpha-1}}\right)^{x_0} \cdot (g^{p^\alpha})^{\sum_{i=1}^{t-1} x_i p^{i-1}} = (h)^{p^{\alpha-1}} \\ &\implies \left(g^{p^{\alpha-1}}\right)^{x_0} = (h)^{p^{\alpha-1}}, \quad \text{since } g^{p^\alpha} = 1_G. \end{aligned}$$

Since $\text{ord}(g^{p^{\alpha-1}}) = p$, we can use our oracle to obtain $x_0 \in \{0, 1, 2, \dots, p-1\}$.

Assuming we have successfully solved for $x_0 \in \{0, 1, 2, \dots, p-1\}$, we raise both sides of $g^x = h$ to the power $p^{\alpha-2}$ and obtain:

$$\begin{aligned} \left(g^{x_0 + x_1 p + \sum_{i=2}^{t-1} x_i p^i}\right)^{p^{\alpha-2}} &= (h)^{p^{\alpha-2}} \implies (g^{x_0})^{p^{\alpha-2}} \cdot (g^{x_1 p})^{p^{\alpha-2}} \cdot \left(g^{\sum_{i=2}^{t-1} x_i p^i}\right)^{p^{\alpha-2}} = (h)^{p^{\alpha-2}} \\ &\implies (g^{x_0})^{p^{\alpha-2}} \cdot \left(g^{p^{\alpha-1}}\right)^{x_1} \cdot (g^{p^\alpha})^{\sum_{i=2}^{t-1} x_i p^{i-2}} = (h)^{p^{\alpha-2}} \\ &\implies \left(g^{p^{\alpha-1}}\right)^{x_1} = (h \cdot g^{-x_0})^{p^{\alpha-2}}, \quad \text{since } g^{p^\alpha} = 1_G. \end{aligned}$$

We can now use our oracle to obtain $x_1 \in \{0, 1, 2, \dots, p-1\}$.

Assuming we have successfully solved for $x_0, x_1 \in \{0, 1, 2, \dots, p-1\}$, we raise both sides of $g^x = h$ to the power $p^{\alpha-3}$ and obtain:

$$\begin{aligned} & \left(g^{\sum_{i=0}^1 x_i p^i + x_2 p^2 + \sum_{i=3}^{t-1} x_i p^i} \right)^{p^{\alpha-3}} = (h)^{p^{\alpha-3}} \\ \implies & \left(g^{\sum_{i=0}^1 x_i p^i} \right)^{p^{\alpha-3}} \left(g^{x_2 p^2} \right)^{p^{\alpha-3}} \left(g^{\sum_{i=3}^{t-1} x_i p^i} \right)^{p^{\alpha-3}} = (h)^{p^{\alpha-3}} \\ \implies & \left(g^{\sum_{i=0}^1 x_i p^i} \right)^{p^{\alpha-3}} \left(g^{p^{\alpha-1}} \right)^{x_2} \left(g^{p^\alpha} \right)^{\sum_{i=3}^{t-1} x_i p^{i-3}} = (h)^{p^{\alpha-3}} \\ \implies & \left(g^{p^{\alpha-1}} \right)^{x_2} = \left(h \cdot g^{-\sum_{i=0}^{2-1} x_i p^i} \right)^{p^{\alpha-3}}, \quad \text{since } g^{p^\alpha} = 1_G. \end{aligned}$$

We can now use our oracle to obtain $x_2 \in \{0, 1, 2, \dots, p-1\}$.

Assuming we have successfully solved for $x_0, x_1, \dots, x_{k-1} \in \{0, 1, 2, \dots, p-1\}$, we raise both sides of $g^x = h$ to the power $p^{\alpha-k-1}$ and obtain:

$$\begin{aligned} & \left(g^{\sum_{i=0}^{k-1} x_i p^i + x_k p^k + \sum_{i=k+1}^{t-1} x_i p^i} \right)^{p^{\alpha-k-1}} = (h)^{p^{\alpha-k-1}} \\ \implies & \left(g^{\sum_{i=0}^{k-1} x_i p^i} \right)^{p^{\alpha-k-1}} \left(g^{x_k p^k} \right)^{p^{\alpha-k-1}} \left(g^{\sum_{i=k+1}^{t-1} x_i p^i} \right)^{p^{\alpha-k-1}} = (h)^{p^{\alpha-k-1}} \\ \implies & \left(g^{\sum_{i=0}^{k-1} x_i p^i} \right)^{p^{\alpha-k-1}} \left(g^{p^{\alpha-1}} \right)^{x_k} \left(g^{p^\alpha} \right)^{\sum_{i=k+1}^{t-1} x_i p^{i-k-1}} = (h)^{p^{\alpha-k-1}} \\ \implies & \left(g^{p^{\alpha-1}} \right)^{x_k} = \left(h \cdot g^{-\sum_{i=0}^{k-1} x_i p^i} \right)^{p^{\alpha-k-1}}, \quad \text{since } g^{p^\alpha} = 1_G. \end{aligned}$$

We can now use our oracle to obtain $x_k \in \{0, 1, 2, \dots, p-1\}$.

- The preceding two observations together imply that the original discrete logarithm problem in the group G , with $|G| = N = p_1^{\alpha_1} p_2^{\alpha_2} \cdots p_t^{\alpha_t}$, can be reduced to a number of associated DLP's in groups of order p_i , $i = 1, 2, \dots, t$. If each of these associated DLP's can be solved sufficiently quickly in practice, then so can the original DLP in G . These associated DLP's in groups of order p_i will be solved in practice by methods other than Pohlig-Hellman, such as by brute force, or by Shank's Baby-Step-Giant-Step method (*i.e.* the "oracle" mentioned above is either brute-force, or Baby-Step-Giant-Step, *etc.*). In particular, if each p_i is "small" (relative to the capacity of the brute-force or the Baby-Step-Giant-Step methods), then the original DLP in G can be solved.

5.3 The Index Calculus Method