

Computational Algebra

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Transcript

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1 Integer Arithmetic

Topics:

- Addition and Multiplication
- GCD computation
- Primality testing
- Factorization

1.1 Addition and Multiplication

Agreement:

- $a, x \in \mathbb{N}$ represented as $x = \sum_{i=0}^{n-1} a_i \cdot B^i$ $B \in \mathbb{N}_{>1}$ fixed Base ($a_i \in \{0, \dots, B-1\}$)
- if $x \neq 0$, assume $a_{n-1} \neq 0$ then define:
length of $x := l(x) = n$ = number of digits = $\lfloor \log_B(x) \rfloor + 1$
(mnemonic: $\log_B(B) + 1 = 2$)
- $l(0) = 1$
(Amount of memory required to store $x = 0$)
- $l(x) := l(|x|)$
- for $x \in \mathbb{Z}$ represent if as $x = \text{sgn}(x) * |x|$

1.1.1 Algorithm 1 (Simple addition)

input : $x = \sum_{i=0}^{n-1} a_i \cdot B^i$, $y = \sum_{i=0}^{n-1} b_i \cdot B^i$, $x, y \in \mathbb{N}$

output: $x + y = \sum_{i=0}^n c_i \cdot B^i$

- (1) $\sigma = 0$
- (2) for $i = 0, \dots, (n-1)$:
- (3) set $c_i := a_i + b_i + \sigma_i$ and $\sigma := 0$
- (4) if $(c_i \geq B)$
- (5) set $c_i = c_i - B$
- (6) set $\sigma = 1$
- (7) set $c_n = \sigma$

If $B = 2$ then (3) - (6) can be realized by logic gates:

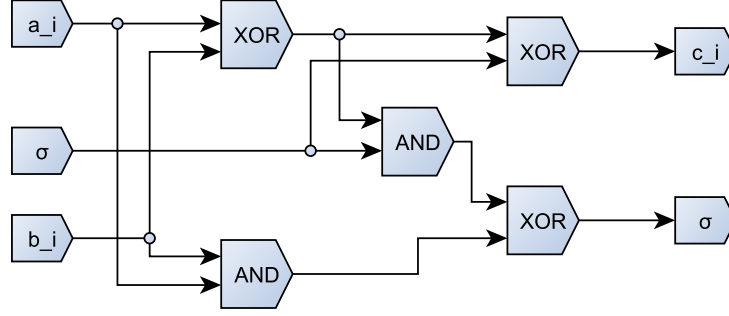


Figure 1: Logic circuit for addition

1.1.2 Definition 2 (Bit-Operation)

A bit operation is an operation that can be performed by a logic gate or by searching or writing a bit from / into memory.

1.1.3 Definition 3 (Big O)

Let M be a set (usually $M = \mathbb{N}$), $f, g : M \mapsto \mathbb{R}_{>0}$
we write $f \in O(g)$ if $\exists c \in \mathbb{R} : f(x) \leq c \cdot g(x) \forall x \in M$

1.1.4 Theorem 4 (Lower bound for addition)

Let $f : \mathbb{N} \mapsto \mathbb{R}$, $n \mapsto$ maximal number of bit operations required by Algorithm 1 to add $x, y \in \mathbb{N}$ with $l(x), l(y) \leq n$

Let $g = id_{\mathbb{N}}$ Then $f \in O(g)$

We say Algorithm 1 requires $O(n)$ bit operations for adding two numbers of length $\leq n$.
 \Rightarrow "linear complexity"

Set $M := \{\text{Set of all algorithms for addition in } \mathbb{N}\}$

For $A \in M$ define $f_A : \mathbb{N} \mapsto \mathbb{R}$ as above.

We would like to find $f_{\text{odd}} : \mathbb{N} \mapsto \mathbb{R}$, $n \mapsto \inf\{f_A(n) | A \in M\}$

Since one needs to read x, y (and write the result) we can not do any better than linear complexity for addition.

Subtraction

let x, y as Algorithm 1, $x \geq y$

For $\bar{y} := \sum_{i=0}^{n-1} (B-1-b_i)B^i$ (digitwise / bitwise complement)

$\Rightarrow x + \bar{y} = x - y + B^n - 1$

$\Rightarrow x - y = x + \bar{y} + 1 - B^n$ (initially set $\sigma = 1$)

Conclusion: Addition and Subtraction have cost $O(n)$

1.1.5 Algorithm 5 (Multiplication by "grid method")

input : $x = \sum_{i=0}^{n-1} a_i \cdot 2^i, \quad y = \sum_{i=0}^{m-1} b_i \cdot 2^i$

output: $z = x \cdot y$

- (1) $z := 0$
- (2) for $i = 0, \dots, (n-1)$
- (3) if $(a_i \neq 0)$ set $z := z + \sum_{j=0}^{m-1} b_j 2^{i+j}$

1.1.6 Theorem 6 (Runtime of Algorithm 5)

Algorithm 5 requires $O(n * m)$ bit operations.

As of the total input length $n + m$:

$$n \cdot m \leq \frac{1}{2}(n + m)^2 \rightarrow O((n + m)^2)$$

\Rightarrow Quadratic complexity

Karatsuba-multiplication:

Observation for polynomials:

$$a + bx, c + dx \text{ have } (a + bx)(c + dx) = ac + (ac + db - (a - b)(c - d))x + bdx^2$$

The point: only used 3 multiplications instead of 4.

Specialize $x = B$ "large" such that $x = a + bB$ partition into two blocks. Then multiply the blocks by a recursive call.

1.1.7 Algorithm 7 (Karatsuba)

input : $x, y \in \mathbb{N}$

output: $z = x \cdot y$

- (1) Choose $k \in \mathbb{N}$ minimal such that $l(x), l(y) \leq 2^k$.
Set $B = 2^{2^{k-1}}$
- (2) if $(k = 0)$ return $x \cdot y$ (by bit-operation AND)
- (3) write $x = x_0 + x_1 B, \quad y = y_0 + y_1 B$ with $l(x_i), l(y_i) \leq 2^{k-1}$
- (4) compute $x_0 \cdot y_0, \quad x_1 \cdot y_1, \quad (x_0 - x_1) \cdot (y_0 - y_1)$ by a recursive call
- (5) return $z = x_0 y_0 + (x_0 y_0 + x_1 y_1 - (x_0 - x_1)(y_0 - y_1))B + x_1 y_1 B^2$

1.1.8 Theorem 8 (Runtime of Algorithm 7)

For multiplying two numbers of length $\leq n$ Algorithm 7 requires $O(n^{\log_2 3}) \approx O(n^{1.59})$ bit operations.

Proof:

Set $\Theta(k) :=$ maximal numbers of bit operations for $l(x), l(y) \leq 2^k$

We have for $k > 0$: $\Theta(k) \leq 3 \underbrace{\Theta(k-1)}_{\text{recursive calls}} + c \underbrace{2^k}_{\text{additions}}$ with $(c \text{ some constant})$

Claim: $\Theta(k) \leq 3^k + 2c(3^k - 2^k)$

Proof by Induction on k :

$k = 0$: $\Theta(k) = 1$

$$\begin{aligned} k-1 \rightarrow k : \Theta(k) &= 3\Theta(k-1) + c2^{k-1} \\ &\leq 3(3^{k-1} + 2c(3^{k-1} - 2^{k-1})) + c2^k \\ &= 3^k + 2c(3^k - 2^k) \end{aligned}$$

So $\Theta(k) \leq (2c+1)3^k$

Now $l(x) \leq n$ hence $2^{k-1} < n$ by minimality of k

So $k-1 < \log_2 n$

$$\begin{aligned} \Rightarrow \Theta(k) &\leq 3(2c+1)3^{\log_2(n)} \\ &= 3(2c+1)2^{\log_2(3) \log_2(n)} \\ &= 3(2c+1)n^{\log_2(3)} \quad \square \end{aligned}$$

One can modify the terminal condition of Karatsuba to switch to Grid-Multiplication, which is faster for small numbers.

Fast-Fourier Transform

Reminder: For a function $f : \mathbb{R} \mapsto \mathbb{C}$ define:

$\hat{f} : \mathbb{R} \mapsto \mathbb{C}$ by

$$\hat{f}(\omega) = \int_{\mathbb{R}} f(t) e^{-i\omega t} dt \quad (\text{if it exists})$$

Think of ω as frequency.

Definition (Convolution)

Let $f, g : \mathbb{R} \mapsto \mathbb{C}$

$$(f * g)(x) = \int_{\mathbb{R}} f(t) g(x-t) dt$$

Convolution is analogous to polynomial multiplication **Formula:** $\underbrace{(f * g)}_{\text{(Cauchy formula)}} = \hat{f} \cdot \hat{g}$

For a function $M \mapsto C$ with $|M| < \infty$ we need the discrete Fourier transform (DFT)

1.1.9 Definition 9 (Root of unity)

Let R be a commutative ring with 1. An element $\mu \in R$ is called an n -th root of unity (= root of 1) if $\mu^n = 1$.

It is called primitive if $\mu^i \neq 1$ for $(0 < i < n)$ i.e. $\text{ord}(\mu) = n$

let μ be a primitive n -th root of 1 (e.g. $e^{2\pi \frac{i}{n}} \in \mathbb{C}$)

Then the map $DFT_\mu : R^n \mapsto R^n$

$$(\hat{a}_0, \dots, \hat{a}_n) \mapsto (\hat{a}_0, \dots, \hat{a}_n) \quad \text{with } \hat{a}_i = \sum_{j=0}^{n-1} \mu^{ij} a_j$$

is called discrete Fourier transformation

For polynomials:

$$DFT_\mu : R[x] \mapsto R^n$$

$$f \mapsto (f(\mu^0), \dots, f(\mu^{n-1}))$$

Convolution rule: (from $f(\mu^i)g(\mu^i) = (f * g)(\mu^i)$)

$$DFT_\mu(f * g) = DFT_\mu(f) \cdot DFT_\mu(g) \quad (\text{component wise product})$$

Addition of two polynomials in $R[x]$ of $\deg(n)$ require $O(n)$ ring operations.

Multiplication require $O(n^l)$.

With Karatsuba have $O(n^{\log_2(3)})$ ring operations.

Cost $DFT_\mu(f) \cdot DFT_\mu(g) : O(n)$ ring operations (with μ as $2n$ -th root of 1)

Want: Cheap way of doing DFT and back-transformation.

1.1.10 Algorithm 10 (Fast Fourier transformation FFT)

input : $f \in R[x]$, $\mu \in R$ primitive 2^k -th root of 1, such that $\mu^{2^{k-1}} = -1$

output: $DFT_\mu(f)$

(1) Write $f(x) = g(x^2) + xh(x^2)$ with $f, g, h \in R[x]$

(2) if $(k = 1)$ $// (\Rightarrow \mu = 1)$
return $DFT_\mu(f) = (g(1) + h(1), g(1) - h(1))$

(3) Recursive call: compute $DFT_{\mu^2}(g) = \hat{g}, DFT_{\mu^2}(h) = \hat{h} \in R^{2^{k-1}}$

(4) return $DFT_\mu(f) = (\hat{f}_0, \dots, \hat{f}_{2^k-1})$ with $\hat{f}_i = \hat{g}_i + \mu \hat{h}_i$
where $\hat{g}_i = \hat{g}_{i-2^{k-1}}$ for $i \geq 2^{k-1}$

Note: Components of \hat{g} and \hat{h} are:

$$\hat{g} = g(\mu^{2^i}), \quad \hat{h}_i = h(\mu^{2^i}) \quad \text{so}$$

$$\hat{f}_i := f(\mu^i) = \hat{g}_i(\mu^{2^i}) + \mu \hat{h}_i(\mu^{2^i}) = \hat{g}_i + \mu \hat{h}_i$$

Convention: $\lg(x) = \log_2(x)$

1.1.11 Theorem 11 (Runtime of Algorithm 10)

Let $n = 2^k$, $f \in R[x]$ with $\deg(\psi) < n$

Then Algorithm 10 requires $O(n \cdot \lg(n))$ ring operations.

Better than $O(n^{1+\epsilon})$, $\forall \epsilon > 0$!

Proof:

Set $\Theta(k) = \max$ number of ring operations required. By counting obtain for $k > 1$:

$$\Theta(k) \leq 2\Theta(k-1) + \underbrace{(\text{compute } \mu^i (i \leq 2^{k-1}))}_{2^{k-1}} + \underbrace{(\mu^i \hat{k}_i)}_{2^{k-1}} + \underbrace{(\text{sums and differences})}_{2^k}$$

$$= 2\Theta(k-1) + 2^{k+1}$$

Claim: $\Theta(k) \leq (2k-1)2^k$

$$k=1 : f = a_0 + a_1 \cdot x \quad DFT_\mu(f) = (a_0 + a_1 \cdot a_0 - a_1) \Rightarrow \Theta(a) = 2$$

$$k-1 \rightarrow k : \Theta(k) \leq 2 \cdot \Theta(k-1) + 2^{k+1} \leq 2 \cdot (2k-3) \cdot 2^{k-1} + 2^{k+1} = (2k-1) \cdot 2^k$$

since $k = \lg(n)$ obtain $O(k) \leq (2 \cdot \lg(n) - 1) \cdot n \in O(n \cdot \lg(n)) \quad \square$

Back-transformation?

1.1.12 Definition 12 (Good root of unity)

A primitive n -th root of unity is called good (caveat: this is ad-hoc terminology) if:

$$\sum_{j=0}^{n-1} \mu^{ij} = 0 \quad \text{for } (0 < i < n)$$

example:

(1) $\mu = e^{2\pi \frac{i}{n}}$ is a good primitive root of unity

(2) $R = \mathbb{Z}/(8)$, $\mu = \bar{3} \Rightarrow \mu \cdot B$ is primitive 2^{nd} root of unity
But $\bar{B}^0 + \bar{3}^1 = \bar{u} \neq \bar{0}$ so μ is not good.

1.1.13 Proposition 13 ($DFT_{\mu^{-1}}$)

let $\mu \in R$ be a good root of 1

$$(a) = (a_0, \dots, a_{n-1}) \in R^n \Rightarrow DFT_\mu^{-1}(DFT_\mu(a)) = n \cdot (a) \quad \text{where } n = 1 + \dots + 1 \in R$$

Proof:

$$DFT_\mu(a) = (\hat{a}) = (\hat{a}_0, \dots, \hat{a}_{n-1})$$

$$\text{with } \hat{a}_j = \sum_{k=0}^{n-1} \mu^{jk} a_k$$

$$DFT_{\mu^{-1}}(\hat{a}) = (\hat{\hat{a}}_0, \dots, \hat{\hat{a}}_1)$$

$$\text{with } \hat{\hat{a}}_i = \sum_{j=0}^{n-1} \mu^{-ij} \sum_{k=0}^{n-1} \mu^{jk} a_k = \sum_{k=0}^{n-1} \left(a_k \cdot \underbrace{\sum_{j=0}^{n-1} \mu^{j(k-i)}}_{=0 \text{ if } n \neq k-i \text{ (i.e. } k=i)} \right) = a_i \cdot n \quad \square$$

1.1.14 Proposition 14 (Finding good roots of unity)

let $\mu \in R, n \in \mathbb{N}$

Assume:

- a) R is an integral Domain and μ is a primitive or n -th root of 1
(Integral Domain: nonzero commutative ring in which the product of two nonzero elements is nonzero)
 \Rightarrow Granted by FFT
- b) $n = 2^b, \mu^{\frac{n}{2}} = -1$, then $h > 0 \wedge \text{char}(R) \neq 2$
 $\rightarrow \mu$ is a good primitive n -th root of 1 ("root of unity")

Proof:

- a) for $0 < i < n$

$$\underbrace{(\mu^i - 1)}_{\neq 0} \underbrace{\left(\sum_{j=0}^{n-1} \mu^{ij}\right)}_{=0} = \mu^{in} - 1 = 0$$

$\Rightarrow \mu$ is a good root of unity

- * Let $0 < i < n$, write $i = 2^{k-s} \cdot r$ with r odd $\wedge s > 0$

$$\sum_{j=0}^{2^k-1} \mu^{ij} = \sum_{l=0}^{2^{k-s}-1} \sum_{j=0}^{2^s-1} \mu^{i(l \cdot 2^s + j)}$$

$$\mu^{i \cdot 2^s} = 1$$

$$i \cdot 2^s = 2^{k-s} \sum_{j=0}^{2^s-1} \mu^{ij} = 2^{k-s} \sum_{j=0}^{2^{s-1}-1} (\mu^{ij} + \mu^{i(2^{s-1}+j)})$$

$$\text{But } \mu^{i \cdot 2^{s-1}} = \mu^{2^{k-s} \cdot r \cdot 2^{s-1}} = \mu^{2^{k-1} \cdot r} = (-1)^r = -1$$

$$\text{So } \sum_{j=0}^{n-1} \mu^{ij} = 0 \quad \square$$

- b) $\mu^n = 1, n = 2^k \Rightarrow \text{ord}(\mu) | n \Rightarrow \text{ord}(\mu)$ is power of 2

1.1.15 Algorithm 15 (Polynomial multiplication using DFT)

input : $f, g \in R[x]$ with $\deg(f) + \deg(g) < 2^k =: n$
 $\mu \in R$ as a good root of unity; Assume $2 \in R$ is invertible

output: $h = f \cdot g$

- (1) compute $\hat{f} = DFT_{\mu}(f), \hat{g} = DFT_{\mu}(g)$ with $f, g \in R^n$
- (2) compute $\hat{h} = \hat{f} \cdot \hat{g}$
- (3) compute $(h_0, \dots, h_{n-1}) = DFT_{\mu^{-1}} \hat{h}$ (same as $DFT_{\mu}(\hat{h})$ but with different order)
= Back-transformation $\cdot 2^k$
set $h = \frac{1}{2^k} \sum_{i=0}^{n-1} h_i x^i$

1.1.16 Theorem 16 (Runtime of Algorithm 15)

Algorithm 15 uses $O(n \cdot \log(n))$ ring operations for polynomials of $\deg < n$

Proof:

- Choose k minimal so that $\deg(f) \cdot \deg(g) < 2^k$
 $\Rightarrow 2^{k-1} \leq 2n \Rightarrow k \leq \log(n) + 2$
- $\underbrace{O(2k \cdot 2^k)}_{\text{Step 1}} + \underbrace{2^k}_{\text{Step 2}} + \underbrace{O(k \cdot 2^k) + 2^k}_{\text{Step 3}} \in O(2k \cdot 2^k) = O(n(g(n))) \quad \square$

Goal: Multiplication in \mathbb{N} using DFT

Idea: find roots of 1 in a suitable $\mathbb{Z}/(m)$

Choose $m = 2^l + 1, \mu = \bar{2} \in R$

1.1.17 Proposition 17 (Add and mul in $O(l)$)

Let $m = 2^l + 1, R = \mathbb{Z}/(m)$

Addition in R and multiplication by $\bar{2}^i \in R$ ($0 \leq i < 2l$) can be done in $O(l)$ bit operations

Proof:

- Let $\bar{x} \in R$ with $0 \leq x \leq 2^l$
- Addition: $x + \bar{y}$
 - (1) compute $x + y \in \mathbb{N}$: $O(l)$
 - (2) if $x + y > 2^l + 1$ subtract $2^l + 1$: $O(l)$
 - Multiplication by $\bar{2}^i$ ($0 \leq i < l$)
 - (1) Bit-shift i Bits to the left by relocating in memory:

$$\underbrace{O(\text{length}(i))}_{\text{compute addr. of new first bit}} + \underbrace{l}_{\text{copying}} = O(\log(l)) + l \in O(l)$$
 - Multiplication by $\bar{2}^i$ ($l \leq i < 2l - 1$)
 - (1) Multiplication by $\bar{2}^{i-l}$: $O(l)$
 - (2) take negative $\bar{2}^i \cdot \bar{x} = -\bar{2}^{i-l} \cdot \bar{x}$: $O(l)$

1.1.18 Proposition 18 (Sort of summary)

Let $k, r \in \mathbb{N}, r > 0, m = 2^{2^k \cdot r} + 1, R = \mathbb{Z}/(m), \mu = \bar{2}^r \in R$

$\Rightarrow 2 \in R$ is invertible, μ is a good primitive 2^{k+1} -th root of 1

$\Rightarrow \mu^{2^k} = 1$

Proof: \rightarrow from above

1.1.19 Algorithm 19 (Multiplication using FFT)

input : $x, y \in \mathbb{N}$

output: $Z = x \cdot y$

- (1) Choose $k \in \mathbb{N}$ minimal such that $l(x), l(y) \leq 2^{2^k}$
- (2) if $k \leq 3$, compute $z = x \cdot y$ by Algorithm 5
- (3) set $B = 2^{2^k}$, $m = 2^{2^k \cdot 4} + 1$, $R = \mathbb{Z}/(m)$, $\mu = \bar{2}^4 \in R$
 $(\Rightarrow \text{so } \mu \text{ is a good primitive } 2^{k+1}\text{-th root of 1})$
- (4) write $x = \sum_{i=0}^{2^k-1} x_i \cdot B^i$, same for y with $(0 \leq x_i, y_i < B)$
 possible since $x, y < 2^{2^{2^k}} = 2^{2^k \cdot 2^k} = B^{2^k}$
- (5) compute: $\hat{x} = DFT_\mu(\bar{x}_0, \dots, \bar{x}_{2^k-1}, \underbrace{0, \dots, 0}_{2^k \text{ zeros}}) \in R^{2^{k+1}}$
 same for y
 \rightarrow use FFT
- (6) compute: $\hat{z} = \hat{x} \cdot \hat{b} \in R^{2^{k+1}}$ (component wise multiplication)
 Perform multiplication in R as follows:
 Multiply representatives (non negative and $< m$) by recursive call.
 Then reduce modulo m by "negative bit shift" (see proof of Proposition 17)
- (7) compute: $(\bar{z}_0, \dots, \bar{z}_{2^{k+1}-1}) = \frac{1}{2^{k+1}} DFT_{\mu^{-1}}(\hat{z}) \in R$ with $0 \leq z < m$
- (8) set $z := \sum_{j=0}^{2^{k+1}-1} z_j \cdot B^j$

1.1.20 Theorem 20 (Runtime of Algorithm 19)

Algorithm 19 correctly computes $t = x \cdot y$ and requires $O(n \cdot (\log n)^4)$ bit operations for $l(x), l(y) \leq n$

Proof: Correctness

write $x(t) \sum_{i=0}^{2^k-i} x_i t^i \in \mathbb{Z}[t]$, $y(t)$, $\bar{x}(t) \in R[t], \bar{y}(t), \bar{z}(t)$

by Proposition 18 and Proposition 13 we have $\bar{z}(t) = \bar{x}(t) \cdot \bar{y}(t)$

The l -th coefficient of $x(t) \cdot y(t)$ is $0 \leq \sum_{i+j=l} x_i \cdot y_j < 2^k \cdot B^2 = 2^{k+2 \cdot 2^k} \leq 2^{2^{k+2}} < m$

So $z(t) = x(t) \cdot y(t) \Rightarrow z = z(B) = x(B) \cdot y(B) = x \cdot y$ Cost:

Write $\Theta(k) := \max$ number of bit operations

Analyze Steps:

- (1) compute $\max \{l(x), l(y)\} : O(l(n)) = O(k)$
- (2) $O(1)$
- (3) no bit operations
- (4) compute starting addresses of x_i, y_i in memory: $2 \cdot 2^k$ increments of the address:
 $2 \cdot 2 \cdot 2^k = 2^{k+2}$ bit ops
 $\Rightarrow O(2^k)$
- (5) By Theorem 11 need $O(2 \cdot 2^{k+1} \cdot (k+1))$ operations in R which are additions and multiplications by powers of \bar{z} costing $O(2^{k+2})$ bit operations.
Total for (5): $O(k \cdot 2^{2 \cdot k})$
- (6) 2^{k+1} multiplications of numbers $< m$, i.e. of length $\leq 2^{k+2}$.
So $k' \leq \frac{k+3}{2}$ for k' : the "new" k used in the next recursion level.
For $\alpha \in R_{>0}$ define $\Theta(\alpha) := \Theta(\lfloor \alpha \rfloor)$
Total for (6): $2^{k+1}(\Theta(\frac{k+3}{2}) + \underbrace{O(2^{k+2})}_{\text{reduction (mod } m)})$
- (7) For $DFT_{\mu^{-1}}(\hat{z}) : O(k \cdot 2^{2 \cdot k})$ as (5) Since \bar{z} is a n root of 1, multiplication by $\bar{2}^{-k-1}$ is multiplication by a positive power of $\bar{2}$, which costs $O(2^{k+2})$
Total for (7): $O(k \cdot 2^{2 \cdot k})$
- (8) For $j \leq 2^{k+1}$ have $\sum_{i=0}^{j-1} z_i \cdot B^i \leq (m-1) \sum_{i=0}^{j-1} B^i = (m-1) \frac{B^j-1}{B-1} < 2(m-1) \frac{B^j}{B} = 2^{1+2^{k+2}+(j-1)2^k}$ so the sum has length $(j+3) \cdot 2 + 1$
Adding $z_j \cdot B^j$ to this sum happens at $(j \cdot 2^k)$ -th bit and higher \Rightarrow cost is $O(2^k)$
Total for (8): $O(2^{2 \cdot k})$

Grad total: For $k \geq 4$:

$\Theta(k) \leq 2^{k+1} \cdot \Theta(\frac{k+3}{2}) + c \cdot k \cdot 2^{2 \cdot k}$ with c constant

Also for $k \in \mathbb{R}_{\geq 4}$

Define $\Lambda(k) := \frac{\Theta(k)}{2^{2 \cdot k}} \Rightarrow \Lambda(k) \leq \frac{2^{k+1} \Theta(\frac{k+3}{2})}{2^{2 \cdot k}} + c \cdot k = 16 \cdot \Lambda(\frac{k+3}{2}) + c \cdot k$

Define $\Omega(k) := \Lambda(k+3)$ So for $k \in \mathbb{R}_{>1}$

$$\Omega(k) \leq 16 \cdot \Lambda(\frac{k}{2} + 3) + c \cdot (k+3) = \underbrace{16\Omega(\frac{k}{2})}_{*} + c \cdot (k+3)$$

Claim: For $i \in \mathbb{N}$ with $2^{i-1} \leq k-3$ have:

$$\Lambda(k) \leq 16^i \Omega(\frac{k-3}{2^i}) + c \cdot (k+3)(1+8+\dots+8^{i-1}) + 3 \cdot c \cdot (1+16+\dots+16^{i-1})$$

Proof by induction:

$$i = 0: \Lambda(k) = \Omega(k-3)$$

$$i \rightarrow i+1: \Lambda(k) \leq 16^i \Omega(\frac{k-3}{2^i}) + c \cdot (k-3)(1+\dots+8^{i-1}) + 3 \cdot c \cdot (1+\dots+16^{i-1}) \leq 2^i \leq k-3 \quad *$$

$$\leq 16^i (16\Omega(\frac{k-3}{2^{i+1}})) + c(\frac{k-1}{2^i} + 3) + c(k-3)\dots = \text{claimed result}$$

Take $u \in \mathbb{N}$ minimal with $2^u > k-3 \Rightarrow \Omega(\frac{k-3}{2^u}) \leq \Omega(\lfloor \frac{k-3}{2^u} \rfloor) = \Omega(0) =: D$ (constant)

Note: u roughly is recursion depth

$$\text{Have } 2^{u-1} \leq k-3 \xRightarrow{\text{claim}} \Lambda(k) \leq 16^u \cdot D + c \cdot \underbrace{(k-3)}_{< 2^u} \cdot \frac{8^u-1}{7} + 3c \cdot \frac{16^u-1}{15} \in O(16^u)$$

$$\text{Have } 2^{u-1} \leq k-3 \Rightarrow u \leq \lg(k-3) + 1$$

$$\Rightarrow \Lambda(k) \in O(16^{\lg(k-3)}) = O((k-3)^4)$$

$$\Rightarrow \Theta(k) = 2^{2 \cdot k} \cdot \Lambda(k) \in O(2^{2k} \cdot (k-3)^4)$$

$$\text{Have } 2^{2(k-1)} < \underbrace{n}_{\max\{l(x) \cdot l(y)\}} \Rightarrow k \leq \frac{\lg(n)}{2} + 1$$

$$\text{So } \Theta(k) \in O(n \cdot (\lg(n))^4) \quad \square$$

1.1.21 Theorem 21 (Schönhage-Strassen 1971)

Multiplication of integers of length $\leq n$ can be done in $O(n \cdot \lg(n) \cdot \lg(\lg(n)))$ bit operations. Schönhage-Strassen is used for integers of length ≥ 100.000 .

Asymptotically faster: Fürer's algorithm.

Comments on Bit complexity

1. Memory requirement may explode!
 \Rightarrow No Problem as bit complexity is upper bound for memory requirements, since memory access is included in bit operations
 $(\rightarrow$ only store what is calculated)
2. Computation of addresses in memory take time
 \Rightarrow length of addresses $\approx \lg(\text{memory space})$ computations of addresses $\approx \lg(\text{memory space})^2$
3. As memory requirement gets larger access times will get longer.
 \Rightarrow transportation time for data $\geq \frac{\text{diameter of physical storage}}{2 \cdot \text{speed of light}}$

1.2 Division with remainder, Euclidean algorithm

1.2.1 Algorithm 1 (Division with remainder)

input : $b = \sum_{i=0}^{n-1} b_i 2^i$ $a = \sum_{i=0}^{n+m-1} a_i 2^i$ with $a_i, b_i \in 0, 1$, $b_{n-1} = 1$

output: $r, q \in \mathbb{N}$ such that $a = q \cdot b + r$, $0 \leq r < b$

(1) $r = a$ $q = 0$

(2) for $i = m, m-1, \dots, 0$ do

(3) if $r \leq 2^i \cdot b$ then set $r := r - 2^i \cdot b$, $q = q + 2^i$

1.2.2 Proposition 2 (Runtime of Algorithm 1)

Algorithm 1 is correct and requires $O(n \cdot (m+1))$ bit operations.

Proof:

Always have $a = q \cdot b + r$

Claim:

before setp (3), have $0 \leq 2^{i+1} \cdot b$

$i = m$; $0 \leq r = a < 2^{m+n} = 2^{m+1} \cdot 2^{n-1} \leq 2^{m-1} \cdot b$ $i < m$ By step (3)

So after last passage through the loop $0 \leq r < b$

Running Time: In step(3), have comparison and (possiby) subtraction. Only n bits involved $\Rightarrow O(n)$

Total: $O(b \cdot (m+1))$

Remarks:

- (1) Division with remainder can be reduced to multiplication.
Precisely: given an algorithm for multiplication that requires $M(n)$ bit operations, there exists an algorithm for division with remainder that requires $O(M(n))$ bit operations.
- (2) Practically relevant:
Jebelean's algorithm (1997): $O(n^{\lg 3})$
- (3) Alternatively, may choose $r \in \mathbb{Z}$ such that $\lfloor \frac{-b}{2} \rfloor < r \leq \lfloor \frac{b}{2} \rfloor$
- (4) Algorithm 1 extends to \mathbb{Z} .
- (5) All Euclidean rings have division with remainder (by definition).
(e.g., $R = K[x] \rightarrow$ polynomial ring over field,
 $R = \mathbb{Z}[i] = \{a + bi | a, b \in \mathbb{Z}\} \subseteq \mathbb{C}$, $i^2 = -1$)

1.2.3 Algorithm 3 (Euclidean algorithm)

input : $a, b \in \mathbb{N}$

output: $\gcd(a, b)$ "greatest common divisor"

- (1) set $r_0 := a, \quad r_i := b$
- (2) for $i = 1, 2, 3, \dots$ perform steps (3) and (4)
- (3) if $r_i = 0$ then $\gcd(a, b) = |r_{i-1}|$
- (4) Division with remainder: $r_{i-1} = q \cdot r_i + r_{i+1} \quad r_{i+1} \in \mathbb{Z}$
 $|r_{i+1}| \leq \frac{1}{2}|r_i|$

Example:

$$a = 287, \quad b = 126$$

$$287 = 2 \cdot 126 + 35 \tag{1}$$

$$126 = 4 \cdot 35 - 14 \tag{2}$$

$$35 = (-2) \cdot (-14) + 7 \tag{3}$$

$$-14 = (-2) \cdot 7 + 0 \tag{4}$$

$$\begin{aligned} \text{So: } 7|(-14) &\xRightarrow{(3)} 7|35 \\ &\xRightarrow{(2)} 7|126 \\ &\xRightarrow{(1)} 7|287 \end{aligned}$$

On the other hand take a common divisor d ; $d|287$; $d|126$

$$\xRightarrow{(1)} d|d \xRightarrow{(2)} d|14 \xRightarrow{(3)} d|7$$

1.2.4 Theorem 4 (Correctness of Algorithm 3)

Algorithm 3 is correct.

Proof:

Since $r_{i-1} = q \cdot r_i + r_{i+1}$ every integer $x \in \mathbb{Z}$ satisfies the equivalence $x|r_{i-1}$ and $x|r_i \Leftrightarrow x|r_{i+1}$ and $x|r_i$ so $\gcd(r_{i-1}, r_i) = \gcd(r_i, r_{i+1}) = \gcd(a, b)$ when terminating have $\gcd(a, b) = \gcd(r_{i-1}, 0) = |r_{i-1}| \quad \square$

1.2.5 Theorem 5 (Runtime of Algorithm 3)

Algorithm 3 requires $O(m \cdot n)$ bit operations for $n = l(a), m = l(b)$

Proof:

If $a < b$ then the first passage yields $r_2 = a, r_1 = b$. Cost: $O(n)$

May assume: $a \geq b$. Write $n_i = l(r_i)$

By Proposition 2 $\exists c$ constant such that the total time is $\leq c \cdot \underbrace{\sum_{i=1}^k n_i \cdot (n_{i-1} - n_i + 1)}_{=: \sigma(n_0, \dots, n_k)}$

For $i > 2$: $n_i = n_{i-1} - 1$

Special Case: $n_i = n_{i-1} - 1$ for $i \geq 2$

$\Rightarrow n_i = n_i - i + 1, n_i = m, k = m + 1$

Obtain $\sigma(n_0, \dots, n_k) = m \cdot (n - m + 1) + \sum_{i=2}^{m+1} (m - i + 1) \cdot 2 = m \cdot n - m^2 + m + m(m - 1) = m \cdot n$.

Claim: The special case is the worst (most expensive)!

From any sequence $n_1 > n_2 > \dots > n_k$ get to the special case by iteratively inserting numbers in the gaps. Insert s with $n_{j-1} > s > n_j$.

$\sigma(n_0, \dots, n_{j-1}, s, n_j, \dots, n_k) - \sigma(n_0, \dots, n_k) = \dots = s + (n_{j-1} - s) \cdot (s - n_j)$

$sp\sigma(n_0, \dots, n_k) \leq \sigma(n, m, m - 1, \dots, 2, 1, 0) = n \cdot m \quad \square$

Complexity is quadratic \rightarrow cheap

1.2.6 Algorithm 6 (Extended Euclidean Algorithm)

input : $a, b \in \mathbb{N}$

output: $d = \gcd(a, b)$ and $s, t \in \mathbb{Z}$ such that $d = s \cdot a + t \cdot b$

(1) $r_0 := a, r_1 := b, s_0 := 1, t_0 := 0, s_1 := 0, t_1 := 1$

(2) for $i = 1, 2, \dots$ perform steps (3) - (5)

(3) if $r_i = 0$ set $d = |r_{i-1}|$
 $s := \text{sgn}(r_{i-1}) \cdot s_{i-1},$
 $t := \text{sgn}(r_{i-1}) \cdot t_{i-1}$

(4) division with remainder:
 $r_{i+1} = r_{i-1} - q_i \cdot r_i, \quad \text{with } |r_{i+1}| \leq \frac{1}{2}|r_i|$

(5) set $s_{i+1} := s_{i-1} - q_i \cdot s_i,$
 $t_{i+1} := t_{i-1} - q_i \cdot t_i$

Justification : $r_i = s_i \cdot a + t_i \cdot b$ throughout

Application: $m, x \in \mathbb{N}$ such that m, x co-prime (i.e. $\gcd(x, m) = 1$)

Algorithm 6 yields: $1 = s \cdot x + t \cdot m \Rightarrow s \cdot x \equiv 1 \pmod{m}$

So obtain inverse of $\bar{x} \in \mathbb{Z}/(m)$

1.3 Primality testing

Let $\mathbb{P} \subseteq \mathbb{N}$ be the set of prime numbers.

Challenge: Given $n \in \mathbb{N}$ decide if $n \in \mathbb{P}$

Naive Method: Trivial division by $m \leq \lfloor \sqrt{n} \rfloor$.

Running time is exponential in $l(n)$. Even when restricted to division by prime numbers, need approximately $\frac{\sqrt{n}}{|n|^{1/\sqrt{n}}}$ trivial divisions (prime number theorem)
 \rightarrow hardly any better!

Reminder: (arithmetic modulo m)

G finite group $\Rightarrow \forall a \in G \quad a^{|G|} = 1$ Fermat's little theorem

For $G = (\mathbb{Z}/(p))^\times \quad a^{p-1} \equiv 1 \pmod{p} \in \mathbb{P} \quad \forall a \in \mathbb{Z} \quad \text{with } p \nmid a$

In fact $(\mathbb{Z}/(p))^\times \cong Z_{p-1}$ is cyclic

For $m = p_1^{e_1} \dots p_r^{e_r}$ with $p_i \in \mathbb{P}, e_i \in \mathbb{N}_{>0}$:

$\mathbb{Z}_{(m)} \cong \mathbb{Z}_{(p_1^{e_1})} \oplus \dots \oplus \mathbb{Z}_{(p_r^{e_r})} \Rightarrow \mathbb{Z}_{(m)}^\times \cong \mathbb{Z}_{(p_1^{e_1})}^\times \times \dots \times \mathbb{Z}_{(p_r^{e_r})}^\times$

what is $\mathbb{Z}_{(p^e)}$ for $p \in \mathbb{P}, e \in \mathbb{N}_{>0}$?

1.3.1 Theorem 1 (Cyclic group)

Let $p \in \mathbb{P}$ odd $e \in \mathbb{N}_{>0} \Rightarrow (\mathbb{Z}_{(p^e)})^\times = Z_{(p-1) \cdot p^{e-1}}$ cyclic

Proof:

$(\mathbb{Z}_{(p^e)})^\times \cong Z_{p-1} \Rightarrow \exists z \in \mathbb{Z} : \text{order}(z + p\mathbb{Z}) = p - 1$

Set $a = \bar{z}^{p^{e-1}} \in (\mathbb{Z}_{(p^e)})^\times =: G$

$$a^{p-1} = \bar{z}^{(p-1) \cdot p^{e-1}} = \bar{z}^{|a|} = 1$$

On the other hand, take $i \in \mathbb{Z}$ such that

$$a^i = 1 \Rightarrow \bar{z}^{i \cdot p^{e-1}} \equiv 1 \pmod{p} \Rightarrow (p-1) \mid (i \cdot p^{e-1}) \Rightarrow (p-1) \mid i.$$

So $\text{ord}(a) = p - 1$.

Now consider $b = (p + 1) \in G$

Claim: $\text{ord}(b) = p^{e-1}$

Proof by induction on $k \in \mathbb{N}_{>0}$ that $(p + 1)^{p^{k-1}} \equiv p^k + 1 \pmod{p^{k+1}}$

$k = 1 \quad \checkmark$

$k \rightarrow k + 1$: By induction have $(p + 1)^{p^{k-1}} = 1 + p^k + x \cdot p^{k+1}, \quad x \in \mathbb{Z}$

$$\text{Compute: } (p + 1)^{p^k} = ((1 + p^k) + x \cdot p^{k+1})^p = \sum_{i=0}^p \binom{p}{i} (i + p^k)^{p-i} \cdot x^i \cdot p^{i \cdot (k+1)}$$

$$\stackrel{\text{Only 0-th summand}}{\equiv} (i + p^k) = \sum_{i=0}^p \binom{p}{i} p^{i \cdot k} \stackrel{p \text{ odd}}{\equiv} 1 + p^{k+1} \pmod{p^{k+2}} \quad \checkmark$$

For $k = e$: $(p + 1)^{p^{e-1}} \equiv 1 \pmod{p^e} \Rightarrow b^{p^e} = 1 \Rightarrow \text{ord}(b) \mid p^{e-1}$

But $(p + 1)^{p^{e-2}} \equiv p^{e-1} + 1 \pmod{p^e} \Rightarrow b^{p^{e-2}} \neq 1 \in G$

So $\text{ord}(b) = p^{e-1}$

Claim: $\text{ord}(a \cdot b) = (p - 1)p^{e-1} \quad (\Rightarrow \text{Theorem})$

Let $(a \cdot b)^i = 1 \in G$ with $i \in \mathbb{Z}$

$$\text{Then } 1 = (a \cdot b)^{i \cdot (p-1)} = (a^{p-1})^i \cdot b^{i \cdot (p-1)} = b^{i \cdot (p-1)} \Rightarrow p^{e-1} \mid i \cdot i(p-1) \Rightarrow p^{e-1} \mid i$$

$$\text{Also } 1 = (a \cdot b)^{p^{e-1} \cdot i} = a^{p^{e-1}} \Rightarrow (p-1) \mid p^{e-1} \cdot i \Rightarrow (p-1) \mid i \rightarrow (p-1) \cdot p^{e-1} \mid i \quad \square$$

Reminder: $(\mathbb{Z}/(2^e))^\times \cong Z_2 \times Z_2^{e-2} \quad (e \geq 2)$

1.3.2 Algorithm 2 (Fermat Test)

input : $n \in \mathbb{N}_{>0 \text{ odd}}$

output: " $n \notin \mathbb{P}$ " or "probably $n \in \mathbb{P}$ "

- (1) Choose $a \in 2, \dots, n-1$ randomly
- (2) Compute $a^{n-1} \bmod n$
- (3) If $a^{n-1} \not\equiv 1 \pmod{n}$ return " $n \notin \mathbb{P}$ "
else return "probably $n \in \mathbb{P}$ "

Not very satisfying. Is this fast?

1.3.3 Algorithm 3 (Fast exponentiation)

input : $a \in G$ G is a monoid, $e \in \mathbb{N}$, $e = \sum_{i_0}^{n-1} e_i 2^i$, $e_i \in \{0, 1\}$

output: $a^e \in G$

- (1) Set $b := a$, $y := 1$
- (2) For $i = 0, \dots, n-1$ perform (3) - (4)
- (3) if $e_i = 1$ set $y := y \cdot b$
- (4) set $b := b^2$
- (5) return y

this requires $O(l(e))$ operations in G

For $G = (\mathbb{Z}/(n)_i)$, each multiplication requires $O(l(n)^2)$ bit operations

\Rightarrow Fermat test requires $O(l(n)^3)$ bit operations \rightarrow cubic complexity \rightarrow "fast"!

Example:

$n = 561 = 3 \cdot 11 \cdot 17$ For $a \in \mathbb{Z}$ with $\gcd(a, n) = 1 \Rightarrow$ have $a^{n-1} = (a^2)^{280} \equiv 1 \pmod{3}$

$a^{n-1} \equiv 1 \pmod{n}$ Fermat's test says "probably $n \in \mathbb{P}$ " in 57% of cases.

$n = 2207 \cdot 6619 \cdot 15443$: output "probably $n \in \mathbb{P}$ " in 99,93% of cases.

1.3.4 Definition 4 (Pseudo-prime, witness, Carmichael numbers)

Let $n \in N_{>1} \text{ odd}$, $a \in 1, \dots, n-1$

- (a) n is pseudo-prime to base a if $a^{n-1} \equiv 1 \pmod{n}$
- (b) otherwise a is called a witness of composition of n
- (c) If $n \notin \mathbb{P}$ but $a^{n-1} \equiv 1 \pmod{n} \quad \forall a$ with $\gcd(n, a) = 1$
then n is called a Carmichael number.
There are ∞ Carmichael numbers

1.3.5 Proposition 5 (Number of witnesses)

Let $n \in N_{>1}$, $\text{odd} \wedge \notin \mathbb{P} \wedge \text{not Carmichael}$

$\Rightarrow |\{a \in \mathbb{Z} \mid 0 < a < n, a \text{ is witness of composite of } n\}| > \frac{n-1}{2}$

Proof: Consider

$\phi : (\mathbb{Z}/(n))^{\times} =: G \rightarrow G, \quad \bar{a} \mapsto \bar{a}^{n-1}$

group homomorphism. By assumption,

$|\text{im}(\phi)| > 1 \Rightarrow |\text{Ker}(\phi)| \leq \frac{|G|}{2} < \frac{n-1}{2}$

$\Rightarrow |\{a \in \mathbb{Z} \mid 0 < a < n \text{ a witness of composite of } n\}| > \frac{n-1}{2} \quad \square$

Miller-Rabin Test

1.3.6 Proposition 6 (Inference from Fermat)

Let $p \in \mathbb{P} \text{ odd}$, $a \in \{1, \dots, (p-1)\}$ write $p-1 = 2^k \cdot m$ with $m \text{ odd}$ Then:
 $a^m \equiv 1 \pmod{p}$ or $\exists i \in \{0, \dots, k-1\} : a^{2^i \cdot m} \equiv -1 \pmod{p}$

Proof:

Little Fermat: $\bar{a}^{2^k \cdot m} = 1 \in \mathbb{F}_p$

Assume $\bar{a}^m \neq 1$ take i maximal such that:

$\bar{b} = \bar{a}^{2^i \cdot m} \neq 1 \Rightarrow \bar{b}^2 = 1 \Rightarrow \bar{b} \in \mathbb{F}_p$ is a zero of $x^2 - 1 \in \mathbb{F}_p[x] \Rightarrow \bar{b} = -1$

1.3.7 Algorithm 7 (Miller-Rabin-test)

input : $n \in \mathbb{N}_{>1}, \text{odd}$

output: either " $n \notin \mathbb{P}$ " or "probably $n \in \mathbb{P}$ " \rightarrow Monte Carlo Algorithm.

- (1) write $n - 1 = 2^k \cdot m$ with m odd
- (2) Choose $a \in \{2, \dots, n - 1\}$ randomly
- (3) Compute $b := a^m \pmod n$
- (4) if $(b \equiv \pm 1 \pmod n)$
return "probably $n \in \mathbb{P}$ "
- (5) for $(i = 0, \dots, k - 1)$ do steps (6) - (7)
- (6) set $b := b^2 \pmod n$
- (7) if $(b \equiv -1 \pmod n)$
return "probably $n \in \mathbb{P}$ "
- (8) return $n \notin \mathbb{P}$

1.3.8 Definition 8 (strong pseudo-prime / witness)

Let $n \in \mathbb{N}_{>1}, \text{odd}$ $a \in \{1, \dots, n - 1\}$

- (a) n is called a strongly pseudo-prime to base a if Proposition 6 holds for a and p replaced by n .
- (b) Otherwise a is called a strong witness of composition of n .

Example

Let $n \in \mathbb{N}_{>1}, \mathbb{P} \text{ odd}$

$a = 2$ strong witness if $n < 2047$ (including 561)

2 or 3 strong witness if $n < 1373653$

2,3 or 5 strong witness if $n < 25326001$

1.3.9 Theorem 9 (Bit-complexity of Algorithm 7)

- (a) Algorithm 7 requires $O(l(n)^3)$ bit operations. \rightarrow "qubic complexity" \rightarrow fast!
- (b) if $b \in \mathbb{P}$ then Algorithm 7 returns "probably $b \in \mathbb{P}$ " \rightarrow no false positives.
- (c) if $n \notin \mathbb{P}$ then more than half of the numbers in $\{1, \dots, n - 1\}$ are strong witnesses.

Proof:

- (a) Step 1 takes $O(l(n))$ bit operations:
Using Algorithm 3, we need $O(l(n-1))$ multiplications in $\mathbb{Z}/(n)$ each requiring $O(l(n)^2)$ bit operations.
- (b) Proposition 6
- (c) split in three cases:

Case 1: n is not a Carmichael number. $\xRightarrow{\text{Prop 5}}$ more than half of all numbers are.

Fermat witness thus also strong witness.

Case 2: $n = p^r \cdot l$ with $p \in \mathbb{P}$ $r > 1$ $l \in \mathbb{N}_{>0}$ $p \nmid l$

Theorem 1 $\exists x \in \mathbb{Z}$ such that $x^p \equiv 1 \pmod{p^r}$ $x \not\equiv 1 \pmod{p^r}$

Chinese remainder theorem: $\exists a \in \mathbb{Z}$ such that $a \equiv x \pmod{p^r}$ $a \equiv 1 \pmod{l}$

So $\bar{a}^p = 1 \in \mathbb{Z}/(n) \Rightarrow \bar{a}^n = 1 \Rightarrow \bar{a} \in (\mathbb{Z}/(n))^\times$

i.e. $\gcd(n, a) = 1$ if $\bar{a}^{n-1} = 1$ then $\bar{a} = 1$

But $a \equiv x \not\equiv 1 \pmod{p^r}$ so $\bar{a}^{n-1} \neq 1$ hence n is not Carmichael \rightarrow Case 1.

Case 3: n is a Carmichael number. By Case 2 have $n = p \cdot l$ with $p \in \mathbb{P}$ $p \nmid l$ $l \geq 3$
 n Carmichael: $\forall a \in \mathbb{Z}$ with $\gcd(a, n) = 1$

have $a^{2^k \cdot m} \equiv 1 \pmod{n}$ (where $n-1 = 2^k \cdot m$)

$a^{2^k \cdot m} \equiv 1 \pmod{p}$ Take j minimal such that

$a^{2^j \cdot m} \equiv 1 \pmod{p}$ $\forall a \in \mathbb{Z}$ such that $\gcd(a, n) = 1$

so $0 \leq j \leq l$ in fact, $j > 0$ since $(-1)^{2^0 \cdot m} = -1$ with m odd.

Consider the subgroup $H := \{\bar{a} \in \mathbb{Z}/(n) \mid \bar{a}^{2^{j-1} \cdot m} \in \{1, -1\} \subseteq (\mathbb{Z}/(n))^\times\}$

Let $a \in \{1, \dots, n-1\}$ $\gcd(n, a) = 1$ a not a strong witness.

Claim 1: $\bar{a} \in H$

Case 3.1: $\bar{a}^{2^{j-1} \cdot m} = 1 \Rightarrow \bar{a} \in H$

Case 3.2: $a^{2^{j-1} \cdot m} \not\equiv 1 \pmod{n}$ $a^m \not\equiv 1 \pmod{n}$

$\xRightarrow{\text{a nonwitness}} \exists i$ such that $\underbrace{a^{2^i \cdot m} \equiv -1 \pmod{n}}_*$

$\Rightarrow a^{2^i \cdot m} \equiv -1 \pmod{p} \xRightarrow{\text{def of } j} i < j$

if $i < j-1$ then $a^{2^{j-1} \cdot m} = (a^{2^i \cdot m})^{2^{j-1-i}} \equiv (-1)^{2^{j-1-i}} = 1 \pmod{n}$

$\xRightarrow{\text{with } *} \text{not in case 3.2}$

Claim 2: $H \subseteq (\mathbb{Z}/(n))^\times$ proper subgroup.

By definition of j $\exists x \in \mathbb{Z}$ such that $x^{2^{j-1} \cdot m} \not\equiv 1 \pmod{p}$

Chinese remainder: $\exists a \in \mathbb{Z}$ such that

$a \equiv x \pmod{p}$ $a \equiv 1 \pmod{l}$

$\Rightarrow a^{2^{j-1} \cdot m} \not\equiv 1 \pmod{p} \equiv 1 \pmod{l} \Rightarrow \bar{a} \notin H$

Claim 2 \checkmark

It follows that $|H| \leq \frac{|(\mathbb{Z}/(n))^\times|}{2} < \frac{n-1}{2}$

so the number of witnesses is $\geq n-1-|H| > \frac{n-1}{2}$ \square

Remarks:

- (a) A more careful analysis shows that $2\frac{3}{4}$ of all candidates are strong witnesses
- (b) Calling Algorithm 7 repeatedly decreases the probability of false positives. Running time for prescribed error probability p is $O(\lg(p^{-1} \cdot l(n)^3))$
(Independence assumptions!)

Connection with Riemann hypothesis

Let $n \in \mathbb{N}_{>0}$ $\bar{X} : (\mathbb{Z}/(n))^\times \rightarrow \mathbb{C}^\times$ group homomorphism

$$X : \mathbb{Z} \rightarrow \mathbb{C}, a \mapsto \begin{cases} \bar{X}(\bar{a}) & \text{if } \gcd(a, n) = 1 \text{ for } (\bar{a} = a + n\mathbb{Z}) \\ 0 & \text{otherwise} \end{cases}$$

"residue class character (mod n)

$$Ex : n = 1 \Rightarrow X(a) = 1 \forall a \in \mathbb{Z}$$

Dirichlet L-series:

$$L_X(s) = \sum_{n=1}^{\infty} \frac{X(n)}{n^s} \text{ converges for } s \in \mathbb{C} \text{ until } Re(s) > 1$$

$L_X(s)$ extends to a meromorphic function on $\mathbb{C} \mapsto$ "Dirichlet L-function".

For $n = 1 : L_X(s) = \zeta(s)$ Riemann Zeta-function.

Euler Product:

$$\text{From } (1 - X(p) \cdot p^{-s})^{-1} = \sum_{i=0}^{\infty} (X(p) \cdot p^{-s})^i = \sum_{i=0}^{\infty} \frac{X(p^i)}{p^{is}} \quad \text{derive } L_X(s) = \prod_{p \in \mathbb{P}} \frac{1}{1 - X(p) \cdot p^{-s}}$$

Generalized Riemann hypothesis (GRH):

For X residue class character, $s \in \mathbb{C}$

with $L_X(s) \neq 0$, $0 < Re(s) < 1$ ("critical strip")

then $Re(s) = \frac{1}{2}$

For $X = 1 \rightarrow$ ordinary Riemann hypothesis.

1.3.10 Theorem (Arkeny & Bach)

GRH $\Rightarrow \forall X \neq 1$ residue class character

$$\exists p \in \mathbb{P} : X(p) \neq 1, p < 2 \ln(n)^2$$

Let $H \subsetneq (\mathbb{Z}/(n))^\times =: G$ proper subgroup.

Choose $N \subsetneq G$ maximal proper subgroup such that $H \subseteq N \Rightarrow G/N$ cyclic.

$$\bar{X} : G \mapsto \mathbb{C}^\times \text{ with } N = \text{Ker}(\bar{X}) \Rightarrow H \subseteq \text{Ker}(\bar{X})$$

$$\xrightarrow{\text{GRH, Thm1}} \exists p \in \mathbb{P} : p + n\mathbb{Z} \not\subseteq H, p < 2 \cdot \ln(n)^2$$

Corollary: Assume GRH.

Let $n \in \mathbb{N}_{>1}$ \mathbb{P} odd Then there is a strong witness a of compositeness of n with $a < 2 \cdot \ln(n)^2$.

\rightarrow Obtain deterministic primality test with time $O(\ln(n)^5)$ bit operations.

AKS-test

A deterministic polynomial time primality test \rightarrow "holy grail"

Agrawal, Kayal, Saxena: PRIMES is in P, Annals of Mathematics, 2004.

1.3.11 Proposition 10 (Modulo over ideals)

Let $n \in \mathbb{P}$ $a \in \mathbb{Z} \Rightarrow (x + a)^n \equiv x^n + a \pmod{n}$

where x is a indeterminate and for $r \in \mathbb{N}$:

$$(x + a)^n \equiv (x^n + a) \pmod{n, x^r - 1} \quad (1)$$

(i.e. $(x + a)^n - (x^n + a) = n \cdot f + (x^r - 1) \cdot g$ with $f, g \in \mathbb{Z}[x]$)

Proof:

$$(x + a)^n = \sum_{i=0}^n \binom{n}{i} \cdot a^{n-i} \cdot x^i \quad (\text{where } \binom{n}{i} \text{ is a multiple of } n \text{ for } 0 < i < n)$$

$$\equiv x^n + a^n \quad (\leftarrow \text{little Fermat})$$

$$\equiv x^n + a \quad (1) \text{ follows by weakening this.}$$

Cost analysis for checking (1) with $l = \text{length}(n)$.

Using Algorithm 3, need $O(l)$ multiplications in $\mathbb{Z}[x]/(n, x^r - 1) =: R$

Elements of R are represented as polynomials of degree $< r$,
coefficients between 0 and n .

Multiply polynomials: $O(r^2)$ operation in $\mathbb{Z}/(n) : O(r^2 \cdot l^2)$

since $x^{r+k} \equiv x^k \pmod{x^r - 1}$,

add coefficients of x^{r+k} of product polynomial to coefficients $x^k : O(r \cdot l)$

Total for checking (1): $O(r^2 \cdot l^3)$ bit operations.

Reduction $\pmod{x^r - 1}$ is just for keeping the cost under control.

The following is part of AKS-test:

1.3.12 Algorithm 11 (Test for perfect power)

input : $n \in \mathbb{N}_{>1}$

output: $m, e \in \mathbb{N}$ $e > 1$ such that $n = m^e$ or "n is not a perfect power"

(1) for $(e = 2, \dots, \lfloor \lg(n) \rfloor)$ perform (2) - (7) //possible exponents

(2) set $m_1 = 2, m_2 = n$ //initialize interval $[m_1, m_2]$ for searching $\sqrt[e]{n}$

(3) while($m_1 \leq m_2$) do (4) - (7)

(4) set $m = \lfloor \frac{m_1 + m_2}{2} \rfloor$ // bisect interval

(5) if $m^e = n$ return m, e

(6) if $m^e > n$ set $m_2 = m - 1$

(7) if $m^e < n$ set $m_1 = m + 1$

(8) return "not a perfect power"

Cost: (for $l = \text{length}(n)$)

Compute $m^e : O(\lg(l) \cdot l^2)$ (abort computation once the result exceeds n)

Number of passages through inner loops $\leq \lg(n)$

Number of passages through outer loops $\leq \lg(n)$

Total cost of Algorithm 11: $O(l^4 \cdot \lg(l))$

1.3.13 Algorithm 12 (AKS-test)

input : $n \in \mathbb{N}_{>1}$ of length $l = \text{length}(n) = \lfloor \lg(n) \rfloor + 1$

output: " $n \in \mathbb{P}$ " or " $n \notin \mathbb{P}$ "

- (1) if (n is a perfect power)
return " $n \notin \mathbb{P}$ "
- (2) find $r \in \mathbb{N}_{>1}$ minimal such that $r|n \vee n^i \not\equiv 1 \pmod{r} \quad \forall i = 1, \dots, l^2$
//exhaustive search (we will show that $r \leq l^5$)
- (3) if $r|n$
if ($r = n$) return " $n \in \mathbb{P}$ "
if ($r < n$) return " $n \notin \mathbb{P}$ "
- (4) for $a = 1, 2, \dots, \lfloor \sqrt{r} \cdot l \rfloor$ do (5)
- (5) if $((x+a)^n \not\equiv (x^n + a) \pmod{(n, x^r - 1)})$
return " $n \notin \mathbb{P}$ "
- (6) return " $n \in \mathbb{P}$ "

1.3.14 Lemma 13 (Least common multiple)

For $n \in \mathbb{N}_{>0}$ have $\lambda(n) := \text{lcm}(1, 2, \dots, n) \geq 2^{n-2}$

Proof: For $f = \sum_{i=0}^m a \cdot x^i \in \mathbb{Z}(x) \quad a_i \in \mathbb{Z}$

$$\Rightarrow \int_0^1 f(x) dx = \sum_{i=0}^m \frac{a_i}{i+1} = \frac{k}{\lambda(m+1)}$$

with $k \in \mathbb{Z}$. Consider $f_m = x^m \cdot (1-x)^m$

For $0 < x < 1$:

$$0 < f_m(x) \leq 4^{-m}$$

$$\Rightarrow 0 < \int_0^1 \underbrace{f_m(x)}_{\frac{k_m}{\lambda(2m+1)}} dx \leq 4^{-1}$$

$$\lambda(2 \cdot m + 1) \geq k_m \cdot 4^m \geq 4^m$$

$$\text{For } n \in \mathbb{N}_{>0} \lambda(n) \geq \lambda(2 \cdot \lfloor \frac{n-1}{2} \rfloor + 1) \geq 4^{\lfloor \frac{n-1}{2} \rfloor} \geq 4^{\frac{n-1}{2}} = 2^{n-2} \quad \square$$

Corollary: (not related to AKS)

For $n \in \mathbb{N}$

$$\pi(n) := |\{p \in \mathbb{P} | p \leq n\}| \geq \frac{n-2}{\lg(n)}$$

Proof:

$$2^{n-2} \leq \lambda(n) = \prod_{p \in \mathbb{P}, p \leq n} p^{\lfloor \log_p(n) \rfloor} \leq \prod_{p \leq n} p^{\log_p(n)} = n^{\pi(n)} = 2^{\lg(n)\pi(n)} \quad \square$$

Prime number theorem:

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

Interpretation:

The average distance of two primes around some value $x \in \mathbb{R}_{>1}$ is $\ln(x)$

1.3.15 Lemma 14 (Property of r in Algorithm 12)

In Algorithm 12, have $r \leq l^5$

Proof:

if $r < l^5 \Rightarrow \forall k \in \{2, \dots, l^5\} : \exists i \in \{1, \dots, l^2\}$

$$n^i \equiv 1 \pmod{k}$$

$$\Rightarrow k \mid \prod_{i=1}^{l^2} (n^i - 1)$$

$$\Rightarrow \lambda(l^5) \mid \prod_{i=1}^{l^2} (n^i - 1)$$

$$\xrightarrow[\text{Lemma 13}]{=} 2^{l^5-2} < \prod_{i=1}^{l^2} n^i = n^{\frac{l^2(l^2+1)}{2}}$$

$$\Rightarrow l^5 - l^3 < 4 \quad \text{not true since } l \geq 2 \quad \square$$

1.3.16 Theorem 15 (Bit-Complexity of Algorithm 12)

Algorithm 12 requires $O(l^{16.5})$ bit operations ("polynomial complexity")

Proof:

Step(1): $O(l^4 \cdot \lg(l)) \checkmark$

Step(2): For each r need:

- test $r \mid n : O(l^2)$
- compute all $n^i \pmod{r} : O(l^2 \cdot \lg(r)^2) \xrightarrow[\text{Lemma 14}]{} O(l^2 \cdot \lg(l)^2)$

Step(3): $O(1)$

$$\text{Step(4): } O(\sqrt{r} \cdot l \cdot r^2 \cdot l^3) \xrightarrow[\text{Lemma 14}]{} O(l^{16.5}) \quad \square$$

Reminder: There is a variant of Algorithm 12 with running time $\tilde{O}(l^6)$, i.e., $O(l^6 \cdot \lg(l)^m)$ with $m \in \mathbb{N}$.

Correctness:

For $r \in \mathbb{N}_{>0}$ and $p \in \mathbb{P}$ write $I(r, p) := \{m, f\} \in \mathbb{N} \times \mathbb{F}_p[x] \mid f(x)^m \equiv f(x^m) \pmod{x^r - 1}\}$
 "m is introspective for f and r ".

Example: Proposition 10 says that:

$$(p, x + \bar{a}) \in I(r, p) \text{ for } a \in \mathbb{Z} \quad r \in \mathbb{N}_{>0} \quad p \in \mathbb{P} \quad (1)$$

1.3.17 Lemma 16 (Rules for ideals)

- (a) $(m, f), (m', f) \in I(r, p) \Rightarrow (m \cdot m', f) \in I(r, p)$
- (b) $(m, f), (m, g) \in I(r, p) \Rightarrow (m, f \cdot g) \in I(r, p)$
- (c) $(m \cdot p, f) \in I(r, p), p \nmid r \Rightarrow (m, f) \in I(r, p)$

Proof:

- (a) $f(x)^{m \cdot m'} \equiv f(x^m)^{m'} \pmod{(x^r - 1)}$
 $f(x^m)^{m'} \equiv f(x^{m \cdot r}) \pmod{(x^{m \cdot r} - 1)}$
 But $(x^r - 1) \mid (x^{m \cdot r} - 1)$
- (b) $(f \cdot g)(x)^m = f(x)^m \cdot g(x)^m \equiv f(x^m) \cdot g(x^m) = (f \cdot g)(x^m) \pmod{(x^r - 1)}$
- (c) $(f(x)^m)^p \equiv f((x^m)^p) \stackrel{\text{Frobenius homomorphism}}{\equiv} (f(x^m))^p \pmod{(x^r - 1)}$
 $\Rightarrow (x^r - 1) \mid ((f(x)^m)^p - f(x^m)^p) \stackrel{\text{Frobenius homomorphism}}{\equiv} (f(x)^m - f(x^m))^p$
 $p \nmid r \Rightarrow x^r - 1$ is square free. So
 $(x^r - 1) \mid (f(x)^m - f(x^m)) \Rightarrow (m, f) \in I(r, p) \quad \square$

1.3.18 Theorem 17 (Correctness of Algorithm 12)

Algorithm 12 is correct.

Proof:

If the algorithm terminates in step(1),(3) or (5), it is correct. To show: If it terminates in step(6) it is correct, i. e. $n \in \mathbb{P}$

Claim 1: $\exists p \in \mathbb{P} : p \mid n \quad p \not\equiv 1 \pmod{r} \quad p > r$

Indeed if all prime divisors of n were $\equiv 1 \pmod{r}$ then $n \equiv 1 \pmod{r}$

Contradiction to step(2). All prime divisors of n are $> r$ by step (2) and (3) ✓

Steps(2) and (3) imply that $\gcd(n, r) = 1 \Rightarrow G := \langle \bar{n}, \underbrace{\bar{p}}_{p \bmod r} \rangle \subseteq (\mathbb{Z}/(r))^\times$

Step(2): $\text{ord}(\bar{n}) > l^2 \Rightarrow l^2 < |G| < r$ (2)

Set $s := \text{ord}(\bar{p} \in G) \Rightarrow r \mid (p^s - 1)$ with $q := p^s \Rightarrow r \mid |\mathbb{F}_q^\times| \Rightarrow \exists \zeta \in \mathbb{F}_q$ r -th root of 1

Set $k := \lfloor \sqrt{r} \cdot l \rfloor \quad m := \left(\frac{n}{p}\right)$

By (1) $(p, x + \bar{a}) \in I(r, p)$ with $\bar{a} \in \mathbb{F}_p$

By step(4), have $(n, x + \bar{a}) \in I(r, p)$

For $\underline{e} = e_0, \dots, e_k \in \mathbb{N}_0$ set $f_{\underline{e}} := \prod_{a=0}^k (x + \bar{a})^{e_a}$

Lemma 16 (b): $(p, f_{\underline{e}}) \in I(r, p)$

$(n, f_{\underline{e}}) \in I(r, p)$

$\xRightarrow{\text{Lemma 16(c)}} (m, f_{\underline{e}}) \in I(r, p)$

$\xRightarrow{\text{Lemma 16(a)}} \forall s, t \in \mathbb{N}_0 : (p^s \cdot m^t, f_{\underline{e}}) \in I(r, p)$

$\Rightarrow f_{\underline{e}}(\zeta^{p^s \cdot m^t}) = f_{\underline{e}}(\zeta)^{p^s \cdot m^t}$ (3)

Set $H := \langle \zeta + \bar{a} | a \in \{0, \dots, k\} \rangle \subseteq \mathbb{F}_q^\times$
 $(\zeta \notin \mathbb{F}_p \text{ since } r \nmid (p-1) \text{ by Claim 1})$

Consider: $T := \{(e_0, \dots, e_k) \in \mathbb{N}_0^{k+1} \mid \sum_{a=0}^k e_a < |G|\}$

$\Phi : T \mapsto H, (e_0, \dots, e_k) \mapsto f_{\underline{e}}(\zeta) = \prod_a (\zeta + \bar{a})^{e_a} \in H$

Claim 2: Φ is injective.

Indeed, take $(\underline{e}), (\underline{\hat{e}}) \in T$ such that $\Phi(\underline{e}) = \Phi(\underline{\hat{e}})$

$$\Rightarrow \forall s, t \in \mathbb{N}_0 : f_{\underline{e}}(\zeta^{p^s \cdot m^t}) \stackrel{(3)}{=} f_{\underline{e}}(\zeta)^{p^s \cdot m^t} = f_{\underline{\hat{e}}}(\zeta)^{p^s \cdot m^t} \stackrel{(3)}{=} f_{\underline{\hat{e}}}(\zeta^{p^s \cdot m^t})$$

$f_{\underline{e}} - f_{\underline{\hat{e}}}$ has roots ζ^e with $e \in G$ since $G = \langle \bar{p}, \bar{m} \rangle$

These are all distinct (since ζ is primitive)

But $\deg(f_{\underline{e}} - f_{\underline{\hat{e}}}) < |G|$ So $f_{\underline{e}} - f_{\underline{\hat{e}}} = 0$

Since $k \leq \sqrt{r} \cdot l < r < p$ the $(x + \bar{a})$ with $a \in \{0 \dots k\}$ are primitive distinct.

So $(\underline{e}) = (\underline{\hat{e}})$ ✓

So is $|H| \geq |T|$?

Let M be the set of all $\{x_0, \dots, x_k\} \subseteq \{1, \dots, |G| + k\}$

with $x_0 < x_1 < \dots < x_k$

For $\{x_0, \dots, x_k\} \in M$ define $(e_0, \dots, e_k) \in \mathbb{N}_0^{k+1}$ by $e_a = x_a - x_{a-1}$ with $x_{-1} := 0$

$$\Rightarrow \sum_{a=0}^k e_a = \sum_{a=0}^k (x_a - x_{a-1} - 1) = x_k - (k+1) < |G|$$

Obtain injection $M \Leftrightarrow T$

$$\text{So } |H| \geq |T| \geq |M| = \binom{|G|+k}{k+1} \stackrel{(2)}{\geq} \binom{\lfloor l\sqrt{|a|} \rfloor + 1 + k}{k+1} = \binom{\lfloor l\sqrt{|a|} \rfloor + 1 + k}{\lfloor l\sqrt{|a|} \rfloor} \stackrel{(2)}{\geq} \binom{2 \cdot \lfloor l\sqrt{|a|} \rfloor + 1}{\lfloor l\sqrt{|a|} \rfloor}$$

1.3.19 Lemma 18 (Property of binomial coefficients)

$$\forall n \in \mathbb{N}_{>1} : \binom{2 \cdot n + 1}{n} > 2^{n+1}$$

Proof:

$n = 2 :$

$$\binom{5}{2} = 10 > 2^3$$

$n - 1 \rightarrow n :$

$$\binom{2 \cdot n + 1}{n} = \binom{2 \cdot n}{n-1} + \binom{2 \cdot n}{n} = \binom{2 \cdot n - 1}{n-2} + \binom{2 \cdot n - 1}{n-1} + \binom{2 \cdot n - 1}{n-1} + \binom{2 \cdot n - 1}{n} \geq 2 \cdot \binom{2 \cdot n - 1}{n-1} \stackrel{ind.}{>} 2 \cdot 2^n = 2^{n+1}$$

Continuation of Proof of Theorem 17

$$|H| > 2^{\lfloor l \cdot \sqrt{|a|} \rfloor + 1} \geq 2^{l \cdot \sqrt{|a|}} \geq 2^{\lg(n) \cdot \sqrt{|a|}} = n \sqrt{|a|} \quad (4)$$

Assume $n \notin \mathbb{P}$ By step (1) m is not a perfect power

\Rightarrow the map $\mathbb{N}_0 \times \mathbb{N}_0 \mapsto \mathbb{N} \quad (s, t) \mapsto p^s m^t$ is injective.

Set $A := \{p^s m^t \mid s, t \in \{0, \dots, \lfloor \sqrt{a} \rfloor\}\} \subseteq \mathbb{N}$

$$\Rightarrow |A| = (\lfloor \sqrt{|a|} \rfloor + 1)^2 > |G|$$

Since $G = \langle \bar{p}, \bar{m} \rangle \subseteq (\mathbb{Z}/(r))^\times$ this implies that $\exists n, \hat{n} \in A$

such that $n \neq \hat{n}$ but $b \equiv \hat{n} \pmod{r}$.

$$\text{Let } h \in H \Rightarrow h = f_{\underline{e}}(\zeta) \text{ with } (\underline{e}) \in \mathbb{N}_0^{k+1} \Rightarrow h^n \stackrel{(3)}{=} f_{\underline{e}}(\zeta^n) \stackrel{n \equiv \hat{n} \pmod{r}}{=} f_{\underline{e}}(\zeta^{\hat{n}}) \stackrel{(3)}{=} h^{\hat{n}}$$

So the polynomial $Y^n - Y^{\hat{n}} \in \mathbb{F}_q[Y]$ has all elements of H as zeros.
 But $\deg(Y^n - Y^{\hat{n}}) \leq \max\{n, \hat{n}\} \leq (p \cdot m)^{\lfloor \sqrt{|G|} \rfloor} \leq n\sqrt{|G|} < |H|$
 \Rightarrow contradiction since $Y^n - Y^{\hat{n}} \neq 0$ \square

1.4 Cryptology

A ("Alice") wants to send a message to B ("Bob") such that an eavesdropper E ("Eve") can not read the clear message. So A and B encrypt the message.

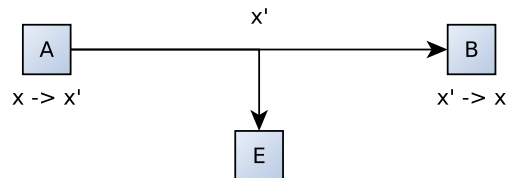


Figure 2: Scheme of eavesdropping

Symmetric-key cryptography

A and B share secret keys for encryption ($x \mapsto x'$) and decryption ($x' \mapsto x$). Only A and B know the keys.

Example: AES approved by the US government in 2002

Application:

- sending messages
- encrypt files (A=B)

Problem: Key exchange between A and B

Public-key cryptography

Encryption-map $\phi : x \mapsto x'$ is made public by B, but decryption $\phi : x' \mapsto x$ is kept secret.

Advantage: No confidential key exchange.

Disadvantages:

- more costly than symmetric key cryptography
- doubt whether E can reconstruct ϕ^{-1} from ϕ with enough computing power

Applications:

- sending messages
- exchange of symmetric keys
- authentication: Together with x , B sends $\phi^{-1}(x)$ (or $\phi^{-1}|$ Part of x together with date). A verifies by applying ϕ .
 Better: challenge-response-protocol.

Examples: RSA, elliptic curve

1.4.1 Algorithm (RSA)

- (1) B chooses $p, q \in \mathbb{P}$ large (> 100 digits)
with $p \neq q$ $n := p \cdot q$
- (2) B chooses $e, f \in \mathbb{N}$ large such that $e \cdot f \equiv 1 \pmod{\phi(n)}$
with $\phi(n) = (p-1)(q-1)$
- (3) B makes n, e public, keep f secret
- (4) The message is encoded as an element $x \in \mathbb{Z}/(n)$
- (5) A computes $\phi(x) = x^e = y \in \mathbb{Z}/(n)$ and sends y
- (6) B receives y and computes $y^f = x \in \mathbb{Z}/(n)$

Comments on steps of RSA:

- (6) Have $e \cdot f = a \cdot (p-1) \cdot (q-1) + 1$ with $a \in \mathbb{N}_{>0}$
 $y^f = x^{e \cdot f}$

Case 1: $q \nmid f, q \nmid x \Rightarrow x^{a(p-1)(q-1)} = (x^{\phi(n)})^a \stackrel{\text{LittleFermat}}{\equiv} 1^a = 1 \Rightarrow x^{e \cdot f} = x \quad \checkmark$

Case 2: $p|x, q \nmid x \Rightarrow x^{e \cdot f} \equiv 0 \equiv x \pmod{p}$
 $x^{e \cdot f} \equiv x \pmod{q}$ as above.

Case 3: $q|x$ As Case 2

\Rightarrow Correctness of decryption

Cost:

- (1) Finding p, q of length approximately l . Prime-number theorem: Gap between two primes of length $\approx l$ is $O(l)$
Using Miller Rabin with error probability 2^m . Expected cost of (1) is $O(m \cdot l^4)$ bit operations.
- (2) Choose e co-prime to $\phi(n)$ obtain $f = \text{inverse} \pmod{\phi(n)}$ by extended euclidean Algorithm: $O(l^2)$
- (5)(6) Fast exponentiation: $O(l^3)$

Security of RSA: p and q must be so large that factorization of n is "impossible". Assumption that factorization is expensive could not be shown! But could f be obtained without knowing p and q ? The following algorithm gives a negative answer. It shows that the problem of breaking RSA is always basically factorization.

Remember: $\phi(n) | (e \cdot f - 1) \Rightarrow m \leq n^2$

1.4.2 Algorithm 1 (Finding a divisor)

Input : $n \in \mathbb{N}_{>2}$ odd squarefree $\notin \mathbb{P}$ and $m \in \mathbb{N}_{>0}$ such that $\phi(n) \mid m$ $m \leq n^2$

Output: $d \in \mathbb{N}$ with $d \mid n$ $1 < d < n$

- (1) Choose $a \in \{2, \dots, (n-2)\}$ randomly
set $k := m$
- (2) If $d := \gcd(a, n) \neq 1$
return d
- (3) Repeat steps (4) - (8) //while(true)
- (4) compute $d := \gcd(n, a^k - 1)$
- (5) If $d = 1$ go to (1)
- (6) If $d < n$ return d
- (7) if k is odd go to (1)
- (8) set $k := \frac{k}{2}$

Correctness is clear. What about termination and running time?

1.4.3 Proposition 2 (Complexity of Algorithm 1)

Algorithm 1 terminates in expected time $O(l(n)^4)$ bit operations (Las Vegas Algorithm).

Proof:

Set $l := \text{length}(n)$

Have $n = \prod_{i=1}^r p_i$ with $p_i \in \mathbb{P}$ distinct.

$\phi(n) = \prod_{i=1}^r (p_i - 1) \mid m$ So initially all $(p_i - 1)$ divide k .

At some iteration it happens for the first time that $(p_i - 1) \nmid k$

Then $k \equiv \frac{p_1-1}{2} \pmod{(p_1-1)} \Rightarrow a^k \equiv \pm 1 \pmod{p_i}$ -1 occurs for some a

For those j with $(p_j - 1) \mid k$ have $a^k \equiv 1 \pmod{p_j}$

Consider the group homomorphism: $\phi_i(\mathbb{Z}/(n))^\times \mapsto (\mathbb{Z}/(p_1))^\times \times \dots \times (\mathbb{Z}/(p_r))^\times$
 $\bar{a} \mapsto (a^k \pmod{p_1}, \dots, a^k \pmod{p_r})$

The image of ϕ is a product of groups $\{\pm 1\}$ or $\{1\}$ depending whether $(p_i - 1) \nmid k$ or $(p_i - 1) \mid k$

Conclusion:

For at least half of all a 's, $\phi(\bar{a})$ is neither $(1, \dots, 1)$ nor $(-1, \dots, -1)$

If $a^k \equiv 1 \pmod{p_j}$ then $p_j \mid (a^k - 1) \Rightarrow p_j \mid d$

If $a^k \equiv -1 \pmod{p_j}$ then $p_j \nmid (a^k - 1) \Rightarrow p_j \nmid d$

So for these a the algorithm is successful.

This means that the expected number of a 's that need to be tested is ≤ 2

(Since $\sum_{i=1}^{\infty} i \cdot \left(\frac{1}{2}\right)^i = 2$ More generally for $0 < p < 1 : p \cdot \sum_{i=1}^{\infty} i \cdot (1-p)^{i-1} = \frac{1}{p}$)

Analysis of running time (in bit operations) for each a (using gcd is quadratic) leads to the claim. \square

Problems of RSA:

- How difficult is factorization of integers (lower bound?)
- decryption of some or all messages without having f ?

1.4.4 Diffie-Hellmann Key Exchange

Goal: A, B want to exchange a symmetric key via a public channel

- (1) A and B agree on a $p \in \mathbb{P}$ (should be large) and $g \in (\mathbb{Z}/(p))^{\times}$ public
- (2) A chooses $a \in \{2, \dots, (p-2)\}$ randomly and sends $u := g^a$ to B
- (3) B chooses $b \in \{2, \dots, (p-2)\}$ randomly and sends $v := g^b$ to A
- (4) A computes $v^a = (g^b)^a = g^{a \cdot b}$
B computes $u^b = (g^a)^b = g^{a \cdot b}$

\Rightarrow A and B share $g^{a \cdot b}$

Example:

A chooses $a = 7$

$$\bar{3}^7 = \bar{11} \in \mathbb{Z}/(17)$$

$$\bar{13}^7 = \bar{4}$$

B chooses $b = 4$

$$\bar{3}^4 = \bar{13} \in \mathbb{Z}/(17)$$

$$\bar{11}^4 = \bar{4}$$

If Eve reconstructs a, b from g^a and g^b she can compute $g^{a \cdot b}$

The Security of Diffie-Hellmann depends on the difficulty of the discrete logarithm problem (DLP):

Given $g \in G$ element of a group or monoid and given $g^a \in G$, determine a (or determine $a' \in \mathbb{Z}$ such that $g^a = g^{a'}$)

1.4.5 Elliptic curve cryptography (ECC)

ECC uses elliptic curves as groups.

$$y^2 = x^3 + a \cdot x + b \rightsquigarrow y^2 z = x^3 + axz^2 + bz^3$$

ECC uses suitable elliptic curves on \mathbb{F}_a

1.5 Factorization

Let $m \in \mathbb{N}_{>1}$ $n \notin \mathbb{P}$ Find a divisor d with $1 < d < n$. From this we obtain the factorization of n by recursion.

Naive method: Trial division. Cost essentially exponential in $l(n)$

1.5.1 Algorithm 1 (Sieve of Eratosthenes)

Input : $n \in \mathbb{N}_{>1}$

Output: All primes $\leq n$

- (1) Create a list of all numbers $\leq n$
- (2) $p := 2$
- (3) Mark all multiples of p in the List
- (4) if all numbers are marked
return
- (5) Let p be the smallest number that is not marked
- (6) $p \in \mathbb{P}$ Go to (3)

Running time of Algorithm 1 is exponential.

Pollard's rho (ρ) algorithm:

Idea: Choose a function $\mathbb{Z}/(m) \mapsto \mathbb{Z}/(n)$ e.g. $f(x) = x^2 + 1$

Choose $x_0 \in \mathbb{Z}/(n)$ set $x_i := f^i(x_0)$ iterative application.

Let $p \mid n$ be a prime. Since $|\mathbb{Z}/(p)| < \infty$ then $\exists i < j : x_i \equiv x_j \pmod{p}$

Starting at x_i the sequence of x_j will be periodic.

$p \mid x_i - x_j \quad p \mid n \Rightarrow p \mid \gcd(n, x_i - x_j) =: d$

If $x_i \not\equiv x_j \pmod{n}$ (which is not guaranteed) then d is a proper divisor of n .

- Recall that gcd computation is cheap
- Testing all pairs is a lot
- Proposition 2 helps with this

1.5.2 Proposition 2

Let M be a set of functions $f : M \mapsto M$ and $x_0 \in M \quad x_i := f^i(x_0)$

If $x_{t+l} = x_t$ for $l, t \in \mathbb{N} > 0 \quad (\rightarrow t \text{ "off-period", } l \text{ "length of period"})$

$\Rightarrow \exists j \in \mathbb{N}$ with $0 < j \leq t + l$ such that $x_j = x_{2j}$

Proof:

$f^l(x_t) = x_t \Rightarrow \forall a \in \mathbb{N} \quad f^{a \cdot l}(x_t) = x_t \quad \text{Assume } j = a \cdot l \geq t \quad a \in \mathbb{N}$

$x_{2j} = x_{t+(j-t)+a \cdot l} = f^{(j-t)}(x_{t+a \cdot l}) = f^{(j-t)}(f^{al}(x_t)) = f^{(j-t)}(x_t) = x_j$

Case 1 $t = 0 \quad j = l \quad \checkmark$

Case 2 $t > 0 \quad j = t + (-t \bmod l) \in 0, \dots, (l-1) \quad \checkmark$

1.5.3 Algorithm 3 (Pollard's ρ -Algorithm)

Input : $n \in \mathbb{N}_{>1}, n \notin \mathbb{P}$

Output: a proper divisor of n or "FAIL"

- (1) Choose $x \in \{0, \dots, (n-1)\}$ randomly
set $y := x$
- (2) repeat (3)-(6)
- (3) $x := x^2 + 1 \pmod n$ $y := (y^2 + 1)^2 + 1$ $// x := x_j y := x_{2j}$
- (4) $d := \gcd(n, x - y)$
- (5) if $(1 < d < n)$
return d
- (6) if $d = n$
return "FAIL"

One "FAIL" includes no conclusion so you might want to repeat the Algorithm with a different x .

Running time? Assume the $x_i := f^i(x_0)$ are randomly distributed.

When can we expect that a match ($x_i \equiv x_j \pmod p$) occurs? \rightarrow "Birthday Problem"

Lemma (Birthday Problem):

We iteratively choose numbers in $\{1, \dots, n\}$ at random. The expected numbers of choices (if we keep choosing until a number has been chosen twice) is $< \sqrt{\frac{\pi \cdot n}{2}} + 2$

Proof:

Let $s \geq 2$ be the numbers of choices until a match occurs. For $k \in \mathbb{N}$ with $P()$ as probability

$$P(s > k) = \prod_{i=1}^k \left(1 - \frac{i-1}{n}\right) \leq \prod_{i=1}^k e^{-\frac{i-1}{n}} = e^{\sum_{i=1}^k -\frac{i-1}{n}} = e^{\frac{k(1-k)}{2n}} \leq e^{-\frac{(k-1)^2}{2n}}$$

* since $f(x) = e^x - (1-x) \geq 0$ for $x \geq 0$

$$f(0) = 0$$

$$f'(x) \geq 0 \text{ if } x \geq 0$$

$$\sum_{k=0}^{\infty} P(s > k) = 2 + \sum_{k=2}^{\infty} P(s > k) \leq 2 + \sum_{k=2}^{\infty} e^{-\frac{(k-1)^2}{2n}} \leq 2 + \int_1^{\infty} e^{-\frac{(x-1)^2}{2n}} dx$$

$$\stackrel{x:=x-1}{=} 2 + \int_0^{\infty} e^{-\frac{x^2}{2n}} dx = 2 + \int_0^{\infty} e^{-\left(\frac{x}{\sqrt{2n}}\right)^2} dx =$$

$$\stackrel{x:=\frac{x}{\sqrt{2n}}}{=} 2 + \sqrt{2n} \int_0^{\infty} e^{-x^2} dx = 2 + \sqrt{2n} \cdot \frac{\pi}{2} = 2 + \sqrt{\frac{n \cdot \pi}{2}}$$

Example:

People arrive at a party. When can you expect to have two that share their birthday?

\rightarrow when 26 have arrived!

1.5.4 Theorem 4 (Bit-complexity of Algorithm 3)

under suitable assumptions on the distribution $f^i(x)$ for $f(x) = x^2 + 1$ Algorithm 3 has the expected running time of $O(\sqrt[n]{n} \lg(n)^2)$ bit operations

Proof:

By Proposition 2 and the Lemma the expected number of runs through the loop is

$$O(\sqrt[p]{p}) = O(\sqrt[n]{n}) \text{ as } p \leq \sqrt{n}$$

Each run through the loop takes $O(\lg(n)^2)$ bit operations. \square

Pollard's p-1 Algorithm

Motivation: Let $p \mid n$ prime divisor

$$\Rightarrow \forall a \in \mathbb{Z} : a^{p-1} \equiv 1 \pmod{p} \quad \text{whith } \gcd(a, p) = 1$$

$$\Rightarrow \forall m \in \mathbb{Z} \text{ with } (p-1) \mid m : a^m \equiv 1 \pmod{p}$$

$$p \mid \gcd(a^m - 1, n)$$

Let B be an upper-bound for the prime powers dividing $p-1$.

" $p-1$ is B -power-smooth".

$$\text{Then } (p-1) \mid \prod_{(q \leq B) \in \mathbb{P}} q^{\lfloor \log_q(B) \rfloor}$$

Neither p nor B are known! But guess and try B and hope for the best.

1.5.5 Algorithm 5 (Pollard's p-1 method)

Input : $n \in \mathbb{N}_{>1}$ \mathbb{P}

Output: $d \in \mathbb{N}$ with $d \mid n$ $1 < d < n$ or "FAIL"

- (1) Choose a "smoothness bound" B
- (2) Choose $a \in \{2, \dots, (n-2)\}$ randomly
- (3) Use Algorithm 1 to find all $q \in \mathbb{P}$ with $q \leq B$
For every q perform steps (4) - (5)
- (4) $k := q^{\lfloor \log_q(B) \rfloor}$
set $a := a^k \pmod{n}$
compute $d := \gcd(n, a - 1)$
- (5) if $1 < d < n$ return d
- (6) return "FAIL" //or increase B and go to (1)

Consequence: when setting up RSA p, q should be chosen such that $p-1$ and $q-1$ have large prime divisors.

The quadratic sieve (State of the art factorization algorithm)

Observation: if $n = x^2 - y^2$ then $n = (x - y) \cdot (x + y)$

Conversely if $n = a \cdot b$ then $n = \left(\frac{a+b}{2}\right)^2 - \left(\frac{a-b}{2}\right)^2$

1-st Idea: Find $x, y \in \mathbb{Z}$ such that $x^2 \equiv y^2 \pmod{n} \wedge x \not\equiv \pm y \pmod{n}$

Then $n \mid (x - y) \cdot (x + y)$

\Rightarrow for every $y : e \in \mathbb{P}$ with $p \mid n : p \mid (x - y) \vee p \mid (x + y)$

$\Rightarrow p \mid \gcd(x - y, n) \vee p \mid \gcd(x + y, n)$

Since both gcd are $< n$ receive a non-trivial divisor of n

If $x^2 \equiv y^2 \pmod{n}$ how probable is it that $x \equiv \pm y \pmod{n}$?

Let $n = \prod_{i=1}^r p_i^{k_i}$ odd with $p_i \in \mathbb{P}$ distinct.

Assume $p_i \nmid x \forall i = 1 \dots r$ Since $(\mathbb{Z}/(p_i^{k_i}))^\times$ is cyclic there are 2^r classes $y \pmod{n}$ such that $x^2 \equiv y^2 \pmod{n}$

[Reason: These classes are given by $y \equiv \pm x \pmod{p_i^{k_i}}$ These are the only solutions since $\mathbb{Z}/(p_i^{k_i})^\times$ is cyclic of even order.

$G = \langle \sigma \rangle$ cyclic of order $2m$

$x = \sigma^i$ Find $l \in \mathbb{Z}$ such that $x^2 = (\sigma^l)^2$

$\Leftrightarrow 2j \equiv 2l \pmod{2m} \Leftrightarrow j \equiv l \pmod{m}$

$\Leftrightarrow l \equiv j \pmod{2m}$ or $l \equiv j + m \pmod{2m}$]

But have $x \equiv \pm y$ only for $2y$'s.

Failure probability: 2^{1-r}

Handle case $r = 1$ by Algorithm 11 in 1.3

Example 1

$n = 91$ Search $x, y \in \mathbb{Z} \quad k \in \mathbb{Z}$ such that $x^2 = k \cdot n + y^2$

Good chance if x is slightly bigger than $\sqrt{k \cdot n}$

$k := 1 \Rightarrow \sqrt{91} \approx 9,54 \Rightarrow x := 10 \Rightarrow 10^2 = 100 \equiv 3^2 \pmod{91}$

$n = 10^2 - 3^2 = (10 - 3) \cdot (10 + 3) = 7 \cdot 13 \quad \checkmark$

Another try:

$k := 8 \Rightarrow \sqrt{8 \cdot 91} \approx 26,98 \Rightarrow 27^2 \equiv 1^2 \pmod{91} \quad \gcd(26, 91) = 13$

Example 2

$n = 4633 \quad k := 3$

$\sqrt{3 \cdot n} \approx 117,89 \Rightarrow x^2 = 118^2 \equiv 5^2 \pmod{n}$

$\gcd(118 - 5, n) = 113$

$\gcd(118 + 5, n) = 41 \quad \checkmark$

2-nd Idea: Choose $B \in \mathbb{N}$ "smoothness bound" suitable.

Let $p_2, \dots, p_r \in \mathbb{P}$ be all primes $\leq B$ (Algorithm 1) set $p_1 := -1$

The p_i form a "factor basis".

For $a \in \mathbb{Z}$ write $(a \pmod{n})$

for the $x \in \mathbb{Z}$ with $x \equiv a \pmod{n}$ and $-\frac{n}{2} < x \leq \frac{n}{2}$

Procedure:

Search numbers $a_1, \dots, a_m \in \mathbb{Z}$ such that $(a_i^2 \pmod{n}) = \prod_{j=1}^r p_j^{e_{ij}}$

with $e_{ij} \in \mathbb{Z}$ ("B numbers")

So for $\mu_i, \dots, \mu_m \in \mathbb{N}_0$ have $\left(\prod_{i=1}^m a_i^{\mu_i} \right)^2 \equiv \prod_{i=1}^m \prod_{j=1}^r p_j^{\mu_i \cdot e_{ij}} \pmod{n} = \prod_{j=1}^r p_j^{\sum_{i=1}^m \mu_i \cdot e_{ij}} \pmod{n}$

If the vectors (e_{i1}, \dots, e_{ir}) become linearly dependant mod 2 (guaranteed if $m > r$) then $\exists \mu_1, \dots, \mu_m \in \{0, 1\}$ not all 0 such that:

$$\sum_{i=1}^m \mu_i \cdot e_{ij} = 2 \cdot k_j \quad k_j \in \mathbb{N}_0$$

with $x := \prod_{i=1}^m a_i^{\mu_i}$ $y := \prod_{j=1}^r p_j^{k_j}$ obtain $x^2 \equiv y^2 \pmod{n}$

2 Notes

2.1 Notation

- $\mathbb{N} := \mathbb{N}_0$
- $\lg(x) := \log_2(x)$
- $a \mid b$ a is divisible by b \Leftrightarrow $b \bmod a = 0$
 $a \nmid b$ a is not divisible by b \Leftrightarrow $b \bmod a \neq 0$
- $ord(a)$ order of a group element
 $n > 0$ minimal such that $a^n = e$ with neutral element e
if no such n can be found, $ord(a) = \infty$
- $char(A)$ Characteristic: the smallest positive n such that
 $\underbrace{1 + \dots + 1}_n = 0$ with 1 as the multiplicative identity element
 n summands
- $\mathbb{Z}/(m)$ Ring modulo m
polynomial rings measure for " $<$ " relations not the absolute value but max power.
- $lcm(a_1, \dots, a_n)$ "least common multiple of all a_i "
- \underline{e} = vector of e 's
- $\phi(n) := |\{x \in \mathbb{N} : x < n \wedge \gcd(x, n) = 1\}| = |(\mathbb{Z}/(n))^x|$
Euler's totient function

2.2 Various stuff

- Lagrange's theorem
Every element in a finite group has finite order
- Average number of bit operations for an increment:
One operation for the last bit + 50% chance for one on the next bit + 25% on the following etc. \Rightarrow Geometrical row
 \Rightarrow on average two bit operations
- "Monte Carlo Algorithm"
Always terminates in reasonable time but might yield false result.
- "Las Vegas Algorithm"
If it terminates the result is correct. No deterministic running time.
- Chinese remainder theorem
Given a system of congruences $x \equiv a_i \pmod{m_i}$ with $i = 1, \dots, r$
 m_i pairwise co-prime. Then the unique solution is:
 $x \equiv a_1 \cdot b \cdot \frac{N}{m_1} + \dots + a_r \cdot b_r \cdot \frac{N}{m_r} \pmod{N}$ with $b_i \cdot \frac{N}{m_i} \equiv 1 \pmod{m_i}$
- distance between two square numbers:
 $(n+1)^2 - n^2 = 2n + 1$
 \Rightarrow Squares are much more scarce than primes!

2.3 Algebraic structures

- Group $(G, *)$
 - one inner operation $(*)$: $G \times G \mapsto G$
 - associativity: $(a * b) * c = a * (b * c) \quad \forall a, b, c \in G$
 - neutral element $e \in G$: $a * e = e * a = a \quad \forall a \in G$
 - inverse element $a^{-1} \in G$: $a * a^{-1} = a^{-1} * a = e \quad \forall a \in G$
- Abelian group $(G, *)$
 - $(G, *)$ is a group
 - commutativity: $a * b = b * a \quad \forall a, b \in G$
- Finite group $(G, *)$
 - associativity: $(a * b) * c = a * (b * c)$
 - unambiguity of reduction: $(a * x = a * x') \wedge (x * a = x' * a) \Rightarrow x = x'$
 $\Rightarrow x \mapsto x * a$ and $x \mapsto a * x$ is bijective
 $\Rightarrow \exists x : a * x = a \Rightarrow$ neutral element
 $\exists x : a * x = x \Rightarrow$ inverse element
- Cyclic group $(G, *)$
 - G is a group
 - G is generated by one Element: $G = \langle g \rangle = \{g^n | n \in \mathbb{Z}\}$
 - not necessarily finite.
- Semi group $(S, *)$
 - one inner operation $(*)$: $S \times S \mapsto S$
 - associativity: $(a * b) * c = a * (b * c) \quad \forall a, b, c \in S$
- Field $(K, +, \cdot)$
 - two inner operations $(+, \cdot)$ such that:
 - $(K, +)$ is an abelian group with neutral element 0
 - $(K \setminus \{0\}, \cdot)$ is an abelian group with neutral element 1
 - distributivity: $a \cdot (b + c) = a \cdot b + a \cdot c$
 $(a + b) \cdot c = a \cdot c + b \cdot c \quad \forall a, b, c \in K$
- Ring $(R, +, \cdot)$
 - $(R, +)$ is an abelian group
 - (R, \cdot) is a semi group
 - distributivity: $a \cdot (b + c) = a \cdot b + a \cdot c$
 $(a + b) \cdot c = a \cdot c + b \cdot c \quad \forall a, b, c \in R$
- Commutative ring $(R, +, \cdot)$
 - $(R, +, \cdot)$ is a ring
 - commutativity for (\cdot) $a \cdot b = b \cdot a \quad \forall a, b \in R$
- Unitary ring (ring with 1) $(R, +, \cdot)$
 - (R, \cdot) is a semi group
 - (R, \cdot) has a neutral element "1"
- Euclidean ring R
 - $\exists F : R \mapsto \mathbb{N}_0 \cup \{0\}$
such that if $\exists q, r \in R \quad a = b \cdot q + r$ and $r = 0$ or $a, b \in R \quad F(r) < F(b)$

- Polynomial ring $R[X]$
 - R is a commutative unitary ring
 - set of all polynomials with coefficients $\in R$

2.4 Invertible elements

- Let $(\mathbb{Z}/(n), +)$ be a group or $(\mathbb{Z}/(n))^\times$ be a group with multiplication.
- $|(\mathbb{Z}/(n))^\times| = \phi(n)$
- $n \in \mathbb{P}$
 - $\Rightarrow (\mathbb{Z}/(n))^\times = \{\bar{0}, \dots, \bar{p-1}\} \cong (\mathbb{Z}/(p-1), +) = Z_{p-1}$ (cyclic Group Z)
- n is a power of 2
 - $\Rightarrow (\mathbb{Z}/(2^e))^\times \cong \mathbb{Z}/(2) \times \mathbb{Z}/(2^{e-2})$
- n is a power of an odd Prime
 - $\Rightarrow (\mathbb{Z}/(p^k))^\times \cong \mathbb{Z}/(p^{k-1} \cdot (p-1)) \cong Z_{(p^{k-1} \cdot (p-1))}$
- $n = p_1^{k_1} \cdot \dots \cdot p_r^{k_r}$
 - $\Rightarrow (\mathbb{Z}/(n))^\times \cong (\mathbb{Z}/(p_1^{k_1}))^\times \times \dots \times (\mathbb{Z}/(p_r^{k_r}))^\times$