Practice

Time Complexity

$\boldsymbol{f(n)}$	$\boldsymbol{g(n)}$	$\boldsymbol{O/\Omega/\Theta}$
$n-100$	$n-200$	Θ
$n^{1/2}$	$n^{2/3}$	O
$100n+\log n$	$n+(\log n)^2$	$\Theta\ (a)$
$\log 2n$	$\log 3n$	$\Theta\ (b)$
$10\log n$	$\log n^2$	$\Theta\ (c)$
$n^{1/2}$	$5^{log_2 n}$	$O\ (d)$
2^n	2^{n+1}	$\Theta\ (e)$

(a): $n\ dominates\ (\log n)^c \to n+(\log n)^2 = \Theta(n)$

(b): $\log ab = \log a + \log b$

(c): $\log a^b = b\log a$

(d): $5 = 2^{2x}\ where\ x > 0 \to 5^{log_2 n} = (2^{2x})^{log_2 n} = (2^{log_2 n})^{2x} = \Omega(n^{1/2})$

(e): $2^{n+1} = 2\times 2^n$

Fibonacci 1

$$F_n = \begin{cases} 0, if\ n=0 \\ 1, if\ n=1 \\ F_{n-1}+F_{n-2}, otherwise \end{cases}$$

Prove: $F_n = \Omega(\sqrt{2}^n)$.

By trial and error: It appears that $F(n) \geq 2^{n/2}$ for all $n \geq 7$

To prove: For all positive integers $n \geq 7 \to F_n \geq 2^{n/2}$

By induction on n. Base case: $n = 7$

Step, assume: Indeed, true that for all $i = 7, 8, \dots, k \to F_i \geq 2^{i/2}$

To prove: $F_{k+1} \geq 2^{k+1/2}$

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Practice

LHS: $F_{k+1} = F_k + F_{k-1} \geq 2^{k/2} + 2^{(k-1)/2}$

Suffices to prove: $2^{k/2} + 2^{(k-1)/2} \geq 2^{(k+1)/2}$

$2^{1/2} + 1 \geq 2^{2/2}$ (by dividing the above by $2^{(k-1)/2}$)

It is indeed true that $2^{1/2} + 1 \geq 2^{2/2} = 1$

Multiplication

Figure 1.1 Multiplication à la Français.

```
function multiply(x, y)
Input: Two n−bit integers x and y, where y ≥ 0
Output: Their product

if y = 0: return 0
z = multiply(x, ⌊y/2⌋)
if y is even:
   return 2z
else:
   return x + 2z
```

Suppose instead of both x and y being n-bit, x is n-bit and y is m-bit. What is the worst-case time efficiency of $multiply$?

Proposed: $O(nm)$

Time Efficiency:

- $\#\, recursive\ calls\ x\ time/call$
- $\#\, worst\ case\ recursive\ calls = O(m)$
- $Worst\ case\ time/call =$
    - $2z\ is\ at\ worst\ O(n+m) \rightarrow because\ very\ last\ addition\ is\ 2z = xy - x$
    - $x\ is\ n\ bits$
    - $So, addition's\ time: O(max\ \{n, n+m\}\ ) = O(max\ \{n, m\})$

So, final answer: $O(m \times \max\{n, m\})$

Fibonacci 2

Practice

Let F_n be the n[th] Fibonacci number, Prove $F_n = O(2^n)$.

- Somewhere, we have shown: $F_n = \Omega(\sqrt{2}^n)$
- But here, seek to show: There exists positive real $F_n \leq c \cdot 2^n$, for all n in N
- Natural proof strategy for "there exists" – construction (i.e., propose some concrete c, and show that it works)
- Try some small values for n, and see what c would work
  - $n = 0, F_0 = 0, 2^0 = 1 \rightarrow c = 1\ works$
  - $n = 1, F_1 = 1, 2^1 = 2 \rightarrow c = 1\ works$
  - $n = 2, F_2 = 1, 2^2 = 4 \rightarrow c = 1\ works$
  - $n = 3, F_3 = 2, 2^3 = 8 \rightarrow c = 1\ works$
  - $n = 4, F_4 = 3, 2^4 = 16 \rightarrow c = 1\ works$
- Appears that $c = 1$works. Adopt it and check if proof goes through. Now, proof by induction with $c = 1$
- Base case, $n = 1, F_1 = 1, 2^1, 1 \leq 2 \rightarrow True$
- Step: Seek to show $F_n \leq 2^n$ given that $F_k \leq 2^k$ for all $k = 1, 2, \ldots, n-1$
- $F_n = F_{n-1} + F_{n-2} \leq 2^{n-1} + 2^{n-2}$ by induction assumption
- $F_n = 2^{n-2}\ (2+1) = 3 \times 2^{n-2} \leq 2^n = 2^2 \times 2^{n-2} = 4 \times 2^{n-2} \rightarrow Done$

Fibonacci 3

Let F_n be the n[th] Fibonacci number, Prove $F_n \neq O(n^2)$.

- Recall from logic: not (there exists an egg-laying mammal) = for all mammals m, m is not egg-laying
- Here, $f = O(g)$: There exists positive real c, for all natural n, $f(n) \leq c \cdot g(n)$
- So here, need to prove: Given any positive real c, it is true that there exists n such that $F_n > c \cdot n^2$
- By contradiction: Suppose that there exists positive real c, such that, for all natural n, $F_n \leq c \cdot n^2$
- Then: $F_n = F_{n-1} + F_{n-2} \leq c(n-1)^2 + c(n-2)^2 = c(n^2 - 2n + 1 + n^2 - 4n + 4) = c(2n^2 - 6n + 5) \leq cn^2$
- $2n^2 - 6n + 5 \leq n^2$

Practice

- $2 - \frac{1}{n^2}(6n - 5) \leq 1$
- This is true only if $\frac{1}{n^2}(6n - 5)$ is "large" compared to $2n^2$
- What is large? We need $\frac{1}{n^2}(6n - 5) \geq 1 \rightarrow true\ for\ n = 1$
- Try $n = 2$: $\frac{1}{4}(12 - 5) = \frac{7}{4} \geq 1$
- Try $n = 3$: $\frac{1}{8}(18 - 5) = \frac{13}{8} \geq 1$
- Try $n = 4$: $\frac{1}{16}(24 - 5) = \frac{19}{16} \geq 1$
- Try $n = 5$: $\frac{1}{25}(30 - 5) = 1$
- Try $n = 6$: $\frac{1}{36}(36 - 5) < 1$
- Try $n = 7$: $\frac{1}{49}(42 - 5) < 1$
- Prove by induction: $6n - 5 < n^2$ for all natural $n > 5$
- Base case $n = 6$: See above
- Step: $6(n - 1) - 5 \leq (n - 1)^2 \rightarrow from\ induction\ assumption$
- $6n - 5 - 6 < n^2 - 2n + 1$
- $6n - 5 \leq n^2 - (2n - 7) \leq n^2\ whenever\ 2n - 7 \geq 0 \rightarrow which\ it\ is\ for\ n \geq 6$
- So far: We have shown that indeed, for $n \geq 6$, $F_n < cn^2 \rightarrow Done$

Selection Sort

$SELECTIONSORT(A[1, \dots, n])$

$foreach\ i\ from\ 1\ to\ n\ do$

$m \leftarrow i - 1 + INDEXOFMIN(A[i, \dots, n])$

$if\ \ i \neq m\ then\ swap\ A[i], A[m]$

$INDEXOFMIN(B[1, \dots, m])$

$min \leftarrow B[1], idx \leftarrow 1$

$foreach\ j\ from\ 2\ to\ m\ do$

$if\ \ B[j] < \min then$

$\min \leftarrow B[j], idx \leftarrow j$

$return\ idx$

What is a meaningful characterization of the time efficiency of $SELECTIONSORT$?

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Practice

- Suppose we invoke $INDEXOFMIN(A[5, \dots, 13])$. In $INDEXOFMIN$: $B[1, \dots, 9]$. Suppose now, min is at index 3 in $B[1, \dots, 9]$. This → index of a min in $A[5, \dots, 13]$ is at index $(5-1)+3=7$
- Suppose on input: $A[1, \dots, 5] = [13, -23, 45, -23, 1]$. Then A evolves in $SELECTIONSORT$ as follows:
  - $i = 1, m = 2, [-23, 13, 45, -23, 1]$
  - $i = 2, m = 4, [-23, -23, 45, 13, 1]$
  - $i = 3, m = 4, [-23, -23, 13, 45, 1]$
- For time efficiency: Need to make meaningful assumption(s)
- Customary Assumptions: (1) n is unbounded, (ii) each $A[i]$ is bounded
- What should we count? Suppose we all agree that counting # swaps is a meaningful measure for time efficiency
- Then: $Worst\ case\ \#\ swaps = n - 1 = \Theta(n)$
- Now, let's say we want to get a bit more fine-grained. Incorporate (worst case) time for each swap x # swaps
- So now, time efficiency: $(n-1)+(n-2)+\cdots+1 = \Theta(n^2)$

Modular Simplification

1. Is $6^6 \equiv 5^3\ (mod\ 31)$?

$$6 \times 6 = 36 \equiv 5\ (mod\ 31)$$

So: $(6^2)^3 \equiv (5)^3\ (mod\ 31)$

2. $2^{125} \equiv ?\,(mod\ 127)$

$$2^7 = 128 = 127 + 1$$

So: $128\ mod\ 127 = 1$

Now: $125/7\ = 17 + 6/7$

So: $2^{125} = 2^{17\times7+6} = 2^{17\times7} \times 2^6$

So: $2^{125} \equiv 2^{17\times7} \times 2^6 \equiv (2^7)^{17} \times 2^6 \equiv 1^{17} \times 2^6 \equiv 64\ (mod\ 127)$

Practice

3. Is $4^{1536} - 9^{4824}$ divisible by 35?

$$4^{1536} \equiv 9^{4824}\ (mod\ 35)$$

Trick: Keep exponentiating until numbers start to repeat.

Suppose we repeatedly exponentiate 4:

$$4$$

$$\rightarrow 16$$

$$\rightarrow 64 \equiv 29\ (mod\ 35)$$

$$\rightarrow 116 = 35 \times 3 + 11 = 11\ (mod\ 35)$$

$$\rightarrow 9\ (mod\ 35)$$

$$\rightarrow 36 \equiv 1\ (mod\ 35)$$

So: $4^6 \equiv 1\ (mod\ 35)$. And $1536 = 6 \times 256$. So $4^{1536} \equiv 1\ (mod\ 35)$

Now check whether 1536 is divisible by 4. Indeed: $1536 = 4 \times 384$

Repeat with 9. Repeated exponentiation of 9:

$$9$$
$$\rightarrow 81 \equiv 11\ (mod\ 35)$$
$$\rightarrow 99 \equiv 29\ (mod\ 35)$$
$$\rightarrow 261 = 7 \times 35 + 16 \equiv 16\ (mod\ 35)$$
$$\rightarrow 144 \equiv 4 \times 35 + 4 \equiv 4\ (mod\ 35)$$
$$\rightarrow 36 \equiv 1\ (mod\ 35)$$

So: $9^6 \equiv 1\ (mod\ 35)$

Now: $9^{4824} = 9^{804 \times 6} \equiv 1\ (mod\ 35)$.

$\therefore\ It\ is\ divisible\ by\ 35.$

4. $2^{2^{2006}}\ (mod\ 3) = ?$

Practice

$$2^{2^{2006}} = (2^2)^{2^{2005}} = 4^{2^{2005}} \equiv 1\ (mod\ 3)$$

5. Is $5^{30000} - 6^{123456}$ a multiple of 31?

   31 is prime. And $5^{30000} = (5^{30})^{1000} \equiv 1\ (mod\ 31)$.

   Compare with $6^{123456} = 6^{123450} \times 6^6$:

   $$1 \times 6^6 \equiv 5^3 \equiv 125 \equiv 31 \times 4 + 1\ (mod\ 31) \equiv 1\ (mod\ 31)$$

   $\therefore It\ is\ a\ multiple\ of\ 31.$

Show that if a has a multiplicative inverse modulo N, then this inverse is unique (modulo N).

Let's assume $a \epsilon \{1, \ldots, N-1\}$.

Suppose $b, c\ \epsilon \{1, \ldots, N-1\}$ are both multiplicative inverses of $a\ modulo\ N$. Then:

$$ab \equiv 1\ (mod\ N)$$
$$ac \equiv 1\ (mod\ N)$$

$$ab \equiv ac\ (mod\ N)$$
$$ab \cdot b \equiv ac \cdot b\ (mod\ N)\ (1)$$

(1): Substitution Rule:

$$x \equiv x', y \equiv y' (mod\ N)$$
$$xy \equiv x'y' (mod\ N)$$

Then:

$$(ab) \cdot b \equiv (ab) \cdot c\ (mod\ N)\ (2)$$

(2): Commutativity

$$1 \cdot b \equiv 1 \cdot c\ (mod\ N)$$
$$b \equiv c\ (mod\ N)$$
$$b = c$$

Practice

Suppose $p \equiv 3\ (mod\ 4)$. Show that $(p+1)/4$ is an integer.

$$p \equiv 3\ (mod\ 4)$$
$$p = 4k + 3\ for\ some\ k\ \epsilon\ \mathbb{Z}$$

So: $p + 1 = 4k + 4$, which is divisible by 4.

We say that x is a square root of y modulo a prime p if $y \equiv x^2\ (mod\ p)$. Show that if $(i)\ p \equiv 3\ (mod\ 4)$ and $(ii)\ y$ has a square root modulo p, then $y^{(p+1)/4}$ is such a square root.

Let x be the square root of $y\ modulo\ p$. Then: $y \equiv x^2\ (mod\ p)$.

Write $p = 4k + 3$. Then, $\left(y^{\frac{p+1}{4}}\right)^2 = y^{2(p+1)/4} = y^{2(4k+3+1)/4} = y^{2k+2}$

Keep in mind: $(p+1)/4 = k + 1$.

Try plugging in x in the last expression:

Is $y^{2k+2} = x^{4k+4} \equiv x^2$?

So, we're asking: Is $x^{4k+4} - x^2 \equiv 0\ (mod\ p)$?

$$x^{4k+4} - x^2 = (x^{2k+2} - x)(x^{2k+2} + x)$$

So at least one of: $x^{2k+2} - x$ or $x^{2k+2} + x$ must be $\equiv 0\ (mod\ p)$.

- $\frac{(p+1)}{4} = \frac{(4k+3+1)}{4} = k + 1$
- $2 \cdot \frac{(p+1)}{4} = 2k + 2$
- $p - 1 = 4k + 2$

We know: There exists $x\ \epsilon\ \{1, \dots, p-1\}$ such that $y \equiv x^2\ (mod\ p)$.

We seek to prove: $\left(y^{\frac{(p+1)}{4}}\right)^2 \equiv y\ (mod\ p)$. Sufficient condition for that to be true:

$\left(y^{\frac{(p+1)}{4}}\right)^2 \cdot y^{-1} \equiv 1\ (mod\ p) \rightarrow$ is okay, because y is invertible modulo p

$$\Rightarrow (y^{2k+2}) \cdot y^{-1} \equiv 1\ (mod\ p)$$
$$\Rightarrow y^{2k+1} \equiv 1\ (mod\ p)$$

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Practice

$$\Rightarrow (x^2)^{2k+1} \equiv 1\ (mod\ p)$$
$$\Rightarrow x^{4k+2} \equiv 1\ (mod\ p)$$
$$\Rightarrow x^{p-1} \equiv 1\ (mod\ p)$$
$$\Rightarrow True\ (Fermat's\ little\ theorem)$$

Proving Recurrence 1

Suppose $x \epsilon \mathbb{Z}^+, y \epsilon \mathbb{Z}_0^+$. Prove recurrence correctness.

$$x^y = \begin{cases} 1, if\ y = 0 \\ (x^2)^{\lfloor y/2 \rfloor}, if\ y\ is\ even \\ x \cdot (x^2)^{\lfloor y/2 \rfloor}, otherwise \end{cases}$$

Case Analysis:

1. If $y = 0$, then $x^y = x^0$. So, the recurrence is correct for the case where $y = 0$
2. If $y \neq 0, y\ is\ even$: then $\lfloor y/2 \rfloor = y/2$. So $x^y = x^{2\times y/2} = (x^2)^{y/2} = (x^2)^{\lfloor y/2 \rfloor}$
3. If $y \neq 0, y\ is\ odd$: then $\lfloor y/2 \rfloor = (y-1)/2$. So now:

$$x^y = x^{(2\times(y-1)/2)+1} = x^{(2\times\lfloor y/2 \rfloor)+1} = x \cdot x^{2\times\lfloor y/2 \rfloor}$$

Proving Recurrence 2

Let $\langle q, r \rangle$ be the quotient and remainder of x/y and $\langle q', r' \rangle$ be the quotient and remainder of $(\lfloor x/2 \rfloor)/y$. Prove recurrence correctness.

$$\langle q, r \rangle = \begin{cases} \langle 0, 0 \rangle, if\ x = 0 \\ \langle 2q', 2r' \rangle, if\ x\ even\ and\ 2r' < y \\ \langle 2q', 2r' + 1 \rangle, if\ x\ odd\ and\ 2r' + 1 < y \\ \langle 2q' + 1, 2r' - y \rangle, if\ x\ even\ and\ 2r' \geq y \\ \langle 2q' + 1, 2r' + 1 - y \rangle, otherwise \end{cases}$$

To be absolutely clear, what are the quotient and remainder of x/y?

We call q the quotient, and r the remainder if and only if q and r are non-negative integers that satisfy:

$$x = q \cdot y + r, where\ r\ \epsilon\ \{0, 1, \ldots, y - 1\}$$

Practice

Proof by case analysis:

1. If $x = 0$, then $x = 0 = 0 \cdot y + 0$. So, recurrence is correct for this case.
2. If x is even and $2r' < y$: then $\lfloor x/2 \rfloor = x/2$. So:

$$\lfloor x/2 \rfloor = x/2 = q' \cdot y + r'$$

$$x = (2q') \cdot y + 2r'$$

$$q = 2q', r = 2r'$$

Where we infer the last line from the facts that: (i) equation is of the form from definition for quotient and remainder, (ii) $r' \geq 0 \rightarrow 2r' \geq 0$, and (iii) we are given $2r' \leq y - 1$.

3. If x is odd and $2r' + 1 < y$: $\lfloor x/2 \rfloor = (x-1)/2$

$$\lfloor x/2 \rfloor = (x-1)/2 = q' \cdot y + r'$$

$$x - 1 = (2q') \cdot y + 2r'$$

$$x = (2q') \cdot y + (2r' + 1)$$

4. x is even, $2r' \geq y$: $\lfloor x/2 \rfloor = x/2$. So:

$$\lfloor x/2 \rfloor = x/2 = q' \cdot y + r'$$

$$x = (2q') \cdot y + 2r'$$

This is of the form of the definition of quotient and remainder, except that we need to confirm that $2r'$ indeed lies between 0 and $y - 1$. Which it does not necessarily. Actually, we are given that $2r' \geq y$ and therefore not between 0 and $y - 1$. Now we observe:

$$x = (2q') \cdot y + 2r'$$

$$x = (2q' + 1) \cdot y + (2r' - y)$$

Now only question that remains: is it the case that $2r' - y \epsilon \{0, 1, \dots, y - 1\}$?

- Is $2r' - y \geq 0$? Yes, because $2r' \geq y$
- Is $2r' - y \leq y - 1$? Yes, because:

$$r' \leq y - 1$$

$$2r' \leq 2y - 2$$

$$2r' - y \leq y - 2 \leq y - 1$$

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5. x odd, $2r' + 1 \geq y$:

$$\lfloor x/2 \rfloor = (x-1)/2 = q' \cdot y + r'$$

$$x = (2q') \cdot y + (2r' + 1)$$
$$x = (2q' + 1) \cdot y + (2r' + 1 - y)$$

Now:

- $2r' + 1 - y \geq 0$ because $2r' + 1 \geq y$.
- $2r' + 1 - y \leq y - 1$ because:

$$r' \leq y - 1$$
$$2r' + 1 \leq 2y - 1$$
$$2r' + 1 - y \leq y - 1$$

Proving Recurrence 3

Prove that $BinSearch$ is correct.

$BinSearch(A[1, \dots, n], lo, hi, i)$

1. $\boldsymbol{if}\ lo \leq hi\ \boldsymbol{then}$
2. $\quad mid \leftarrow \lfloor (lo + hi)/2 \rfloor$
3. $\quad \boldsymbol{if}\ A[mid] = i\ \boldsymbol{then}\ \boldsymbol{return}\ true$
4. $\quad \boldsymbol{if}\ A[mid] < i\ \boldsymbol{then}\ \boldsymbol{return}\ BinSearch(A, mid + 1, hi, i)$
5. $\quad \boldsymbol{else}\ \boldsymbol{return}\ BinSearch(A, lo, mid - 1, i)$
6. $\boldsymbol{else}\ \boldsymbol{return}\ false$

Above is recursive version of binary search. Iterative version:

$BinSearch(A[1, \dots, n], lo, hi, i)$

1. $\boldsymbol{while}\ lo \leq hi\ \boldsymbol{do}$
2. $\quad mid \leftarrow \lfloor (lo + hi)/2 \rfloor$
3. $\quad \boldsymbol{if}\ A[mid] = i\ \boldsymbol{then}\ \boldsymbol{return}\ true$
4. $\quad \boldsymbol{if}\ A[mid] < i\ \boldsymbol{then}\ lo \leftarrow mid + 1$

Practice

5. $\boldsymbol{else}\ hi \leftarrow mid - 1$
6. $\boldsymbol{else\ return}\ false$

Typically, for iterative algorithms, towards correctness, we articulate a *loop invariant*:

Let $lo^{(in)}$ and $hi^{(in)}$ be the values of lo and hi respectively on input. Just before we successfully enter an iteration of the $\boldsymbol{while}$ loop of Line (1), it is true that:

$$i \epsilon A\left[lo^{(in)}, \dots, hi^{(in)}\right] \rightarrow i \epsilon A[lo, \dots, hi]$$

Going back to the recursive version, what is a correctness property?

Given $A[1, \dots, n]$ an array that is sorted, non-decreasing, lo, hi are each $\epsilon \{1, \dots, n\}$ on input, $BinSearch(A, lo, hi, i)$ returns:

- $True \rightarrow (lo \leq hi)\ and\ (i \epsilon A[lo, \dots, hi])$
- $False \rightarrow either\ (lo > hi)\ or\ (i\ is\ not\ \epsilon\ A[lo, \dots, hi])$

Proof by case analysis:

Case 1: $lo > hi$ on input: then $\boldsymbol{if}$ condition of Line (1) evaluates to **false**, and we correctly return **false** in Line (6). Then, this is either from (a) Line (6) without making any recursive calls, or (b) as the return value from a recursive call from one of Lines (4) or (5).

For (b), we first observe that $lo \leq hi$ because the only recursive calls are within the $\boldsymbol{if}$ block of Line (1). So, all that remains to be proven is that indeed: $i \notin A[lo, \dots, hi]$.

We prove that by induction on $hi - lo + 1$. Base case: $hi - lo + 1 = 1$. We claim we return $\boldsymbol{false}$ within the first recursive invocation. That is, we claim: (i) $mid + 1 > hi$ and $lo > mid - 1$, (ii) $mid = lo = hi$, and (iii) $i \neq A[mid]$.

(ii) easy to prove:

$$hi - lo + 1 = 1$$
$$\Rightarrow lo = hi$$
$$\Rightarrow mid = \left\lfloor \frac{(lo + hi)}{2} \right\rfloor = \left\lfloor \frac{(lo + lo)}{2} \right\rfloor = \left\lfloor \frac{(2 \cdot lo)}{2} \right\rfloor = \frac{2 \cdot lo}{2} = lo = hi$$

(iii) is $\boldsymbol{true}$, because then we would have returned $\boldsymbol{true}$ in Line (3).

To prove (i): we simply exploit: $mid = hi = lo$

$$mid = hi \Rightarrow mid + 1 > hi$$
$$mid = lo \Rightarrow mid - 1 < lo$$

So, the algorithm is correct if it returns $\boldsymbol{false}$, and $hi - lo + 1 = 1$.

For the step, we know that on input $lo < hi$. So, we returned $\boldsymbol{false}$ in some recursive call. So, all we have to prove to appeal to induction assumption: $hi - (mid + 1) < hi - lo$ and $(mid - 1) - lo < hi - lo$.

Proving Master Theorem

Give a closed form solution for the following recurrence. Assume: $f: \mathbb{R}^+ \rightarrow \mathbb{R}^+$ is non-decreasing, $a > 0, b > 1, d \geq 0$.

$$f(n) = \begin{cases} \Theta(1), if n \leq 1 \\ a \cdot f\left(\frac{n}{b}\right) + \Theta(n^d), otherwise \end{cases}$$

Proposed approach: Inductive "rewriting" of the function f. But first: adopt concrete functions wherever we have $\Theta(\cdot)$, $O(\cdot)$ or $\Omega(\cdot)$. In this case: adopt 1 for $\Theta(1)$, and n^d for $\Theta(n^d)$. Now onto the rewriting:

$$f(n) = a \cdot f\left(\frac{n}{b}\right) + n^d$$
$$= a \cdot \left(a \cdot f\left(\frac{n}{b^2}\right) + \left(\frac{n}{b}\right)^d\right) + n^d$$
$$= a^2 \cdot f\left(\frac{n}{b^2}\right) + a \cdot \left(\frac{n}{b}\right)^d + n^d$$
$$= a^2\left(a \cdot f\left(\frac{n}{b^3}\right) + \left(\frac{n}{b^2}\right)^d\right) + a \cdot \left(\frac{n}{b}\right)^d + n^d$$
$$= a^3 \cdot f\left(\frac{n}{b^3}\right) + a^2 \cdot \left(\frac{n}{b^2}\right)^d + a \cdot \left(\frac{n}{b}\right)^d + n^d$$
$$= a^3 \cdot f\left(\frac{n}{b^3}\right) + n^d\left(\left(\frac{a}{b^d}\right)^2 + \left(\frac{a}{b^d}\right)^1 + \left(\frac{a}{b^d}\right)^0\right)$$

$$= a^4 \cdot f\left(\frac{n}{b^4}\right) + n^d \left(\left(\frac{a}{b^d}\right)^3 + \left(\frac{a}{b^d}\right)^2 + \left(\frac{a}{b^d}\right)^1 + \left(\frac{a}{b^d}\right)^0\right)$$

...

$$= a^{\log_b n} \cdot f(1) + n^d \cdot \left(\left(\frac{a}{b^d}\right)^{(\log_b n)-1} + \left(\frac{a}{b^d}\right)^{(\log_b n)-2} + \cdots + \left(\frac{a}{b^d}\right)^0\right)$$

$$= a^{\log_b n} + n^d \cdot \left(\left(\frac{a}{b^d}\right)^{(\log_b n)-1} + \left(\frac{a}{b^d}\right)^{(\log_b n)-2} + \cdots + \left(\frac{a}{b^d}\right)^0\right)$$

To figure out the power of a in that last term:

> Power of a is the same as the power of b inside the $f\left(\frac{n}{b^x}\right)$. In other words, we're asking: what is the power of b, i.e., x for which $\frac{n}{b^x} = 1$? Answer: $\frac{n}{b^x} = 1 \Leftrightarrow n = b^x \Leftrightarrow x = \log_b n$.

Our next step: Simplify/figure out:

$$S = \left(\frac{a}{b^d}\right)^{(\log_b n)-1} + \left(\frac{a}{b^d}\right)^{(\log_b n)-2} + \cdots + \left(\frac{a}{b^d}\right)^0$$

Suppose:

$$T = r^{q-1} + r^{q-2} + \cdots + r^0$$
$$\Rightarrow r \cdot T = r^q + r^{q-1} + \cdots + r$$

Now subtract one from the other:

$$\Rightarrow T - r \cdot T = r^0 - r^q$$
$$\Rightarrow (1 - r) \cdot T = 1 - r^q$$
$$\Rightarrow T = \frac{1 - r^q}{1 - r}, provided\ r \neq 1$$

When $r = 1$, how do we figure out what T is? Answer: then, T is:

$$T = 1^{q-1} + 1^{q-2} + \cdots + 1^0$$
$$= 1 + 1 + \cdots + 1 \rightarrow q\ instances\ of\ 1$$
$$= q$$

So, going back to our S:

Practice

$$S = \left(\frac{a}{b^d}\right)^{(\log_b n)-1} + \left(\frac{a}{b^d}\right)^{(\log_b n)-2} + \cdots + \left(\frac{a}{b^d}\right)^{0}$$

$$\Longrightarrow S = \frac{1 - \left(\frac{a}{b^d}\right)^{\log_b n}}{1 - \left(\frac{a}{b^d}\right)}, provided \frac{a}{b^d} \neq 1$$

And:

$$S = \log_b n\,, when \frac{a}{b^d} = 1$$

When is $\frac{a}{b^d} = 1$? Answer: $d = \log_b a$.

So, going back to our $f(n)$: first, the case that $d = \log_b a$.

But even before that: rewrite $a^{\log_b n} = n^{\log_b a}$. Because:

$$x = a^{\log_b n} \Leftrightarrow \log_b x = \log_b a \cdot \log_b n \Leftrightarrow x = n^{\log_b a}$$

$$f(n) = n^{\log_b a} + n^d \cdot S$$

So, when $d = \log_b a\,, S = \log_b n$. So, in this case:

$$f(n) = n^d + n^d \cdot \log_b n$$
$$= \Theta(n^d \cdot \log n)$$

Onto the other two cases: $d \neq \log_b a$.

$$f(n) = n^{\log_b a} + \cdots + n^d \cdot S$$

Before we continue: a closer look at $\left(\frac{a}{b^d}\right)^{\log_b n}$:

$$\left(\frac{a}{b^d}\right)^{\log_b n} = \frac{a^{\log_b n}}{(b^d)^{\log_b n}}$$
$$= \frac{n^{\log_b a}}{(b^{\log_b n})^d}$$
$$= \frac{n^{\log_b a}}{n^d}$$

So: when $d \neq \log_b a$

Practice

$$S = \frac{1 - \frac{n^{\log_b a}}{n^d}}{1 - \left(\frac{a}{b^d}\right)}$$

So, going back to $f(n)$:

$$\begin{aligned} f(n) &= n^{\log_b a} + n^d \cdot S \\ &= n^{\log_b a} + \frac{1}{1 - \left(\frac{a}{b^d}\right)} \cdot \left(n^d - n^{\log_b a}\right) \\ &= c \cdot n^{\log_b a} + c' \cdot n^d, for\ positive\ constants\ c, c' \end{aligned}$$

So, if $d > \log_b a$: $f(n) = \Theta(n^d)$

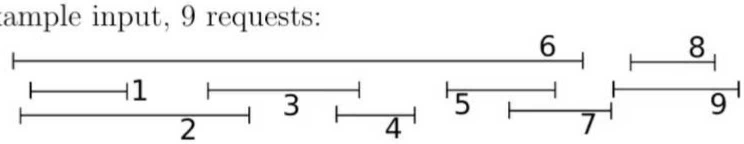
And if $d < \log_b a$: $f(n) = \Theta\left(n^{\log_b a}\right)$

Proving Greediness

Given as input n meeting requests, $\langle s_1, f_1\rangle, \langle s_2, f_2\rangle, \ldots, \langle s_n, f_n\rangle$, where each $s_i, f_i \in \mathbb{Z}^+$ is a start- and finish-time and $s_i < f_i$. We want a subset of those requests that is of maximum size that are pairwise conflict-free.

Two requests $\langle s_i, f_i\rangle$, $\langle s_j, f_j\rangle$ are in conflict if $s_i \leq f_j$, and $s_j \leq f_i$, or vice versa.

Example input, 9 requests:



Request 5 is in conflict with each Request 6 and 7. But is conflict-free with Request 2.

An optimal (maximum-sized) conflict-free set: $\{1, 3, 5, 9\}$. Another: $\{1, 4, 7, 8\}$.

Prove: this problem possesses a greedy choice.

Candidate greedy choice: request with earliest finish time.

Proof strategy: "cut and paste."

For this problem, we prove two claims in order:

Claim 1: *Suppose for some input of n requests, $O = \{o_1, \ldots, o_k\}$ is an optimal (maximum-sized) set of requests which are pairwise conflict-free ordered in increasing finish time. Suppose our* *greedy algorithm outputs $G = \{g_1, \ldots, g_l\}$, ordered in increasing finish time. Then, it is true that: for every $i = 1, 2, \ldots, l, f(g_i) \leq f(o_i)$.*

Proof. Note: it must be the case that $l \leq k$. And therefore, $k = l$, i.e., greedy is optimal.

Proof by induction on i. Base case: $i = 1$. In our greedy algorithm, we first pick exactly a meeting that finishes earliest amongst all requests. Therefore, immaterial of what o_1 is, $f(g_1) \leq f(o_1)$.

Induction assumption: for $i = j - 1$, it is true that $f(g_i) \leq f(o_i)$.

Step: to prove that $f(g_j) \leq f(o_j)$. We observe:

- $f(o_{j-1}) \leq s(o_j)$ – because the set O is conflict-free requests, ordered in increasing finish, and therefore, start times.
- $f(g_{j-1}) \leq f(o_{j-1})$ – induction assumption.
- Therefore, $f(g_{j-1}) \leq s(o_j)$. Therefore $f(g_j) \leq f(o_j)$ – because after we greedily choose g_{j-1} and eliminate all requests that are in conflict, o_j still remains. And our greedy choice is exactly to pick a request that remains that finishes earliest, and we happened to pick g_j.

Claim 2: *Given sets O, G as in Claim 1, o_{l+1} cannot exist in O.*

Proof. By Claim 1, $f(g_l) \leq f(o_l)$. And because the O set is all conflict-free, $f(o_l) \leq s(o_{l+1})$. Therefore, $f(g_l) \leq s(o_{l+1})$. So, o_{l+1} not in conflict with g_l, and so was available to be chosen after g_l was chosen and all conflicts were eliminated.

Contradiction to the assumption that greedy algorithm terminates only when no more requests available to choose from.