

Lecture 1: Introduction: Good vs Bad Algorithms; Cryptography

1 The Central Paradigm of Computer Science

- The central *paradigm* in computer science is that an algorithm \mathbf{A} is good if:
 - \mathbf{A} runs in *polynomial time* in the input size n .
 - That is, \mathbf{A} runs in time $T(n) = O(n^k)$ for some constant number k .
 - $T(n) = 100n + 55$
 - $T(n) = \frac{1}{2}n^2 + 999 \log n$
 - $T(n) = 6n^7 + 900000n^2 - \sqrt{n}$
 - An algorithm is *bad* if it runs in exponential time.
 - $T(n) = 2^n + 100n^5$
 - $T(n) = 1.000000001^n - n^3 - n$
 - An algorithm is *good* if it runs in *polynomial time* in the input size n .

		Input Size n		
		10	100	1000
n		10	100	1000
Runtime of Algorithm	n^2	100	10000	1000000
	2^n	10^3	10^{30}	10^{300}

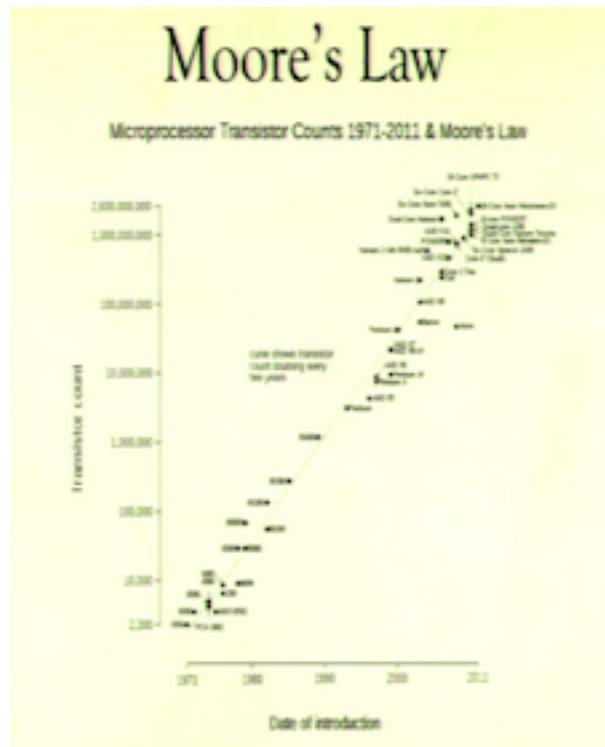
2 Good versus Bad (Algorithms)

- For example, consider the problem of sorting n numbers.
 - A Good Algorithm: **MergeSort** runs in time $O(n \cdot \log n)$
 - A Bad Algorithm: **BruteForce Search** runs in time $O(n \cdot n!) \gg 2^n$

3 An Equivalent Characterization

- This central **paradigm** has an equivalent formulation
 - A runs in **polynomial time** in the input size n .
 - The input sizes that A can solve, in a fixed amount T of time, **scales multiplicatively** with increasing computational power.

		Input Sizes solved in Time T	
		Power = 1	Power = 2
n		T	2T
Runtime of Algorithm	n^2	\sqrt{T}	$\sqrt{2} \cdot \sqrt{T}$
2^n		$\log T$	$1 + \log T$

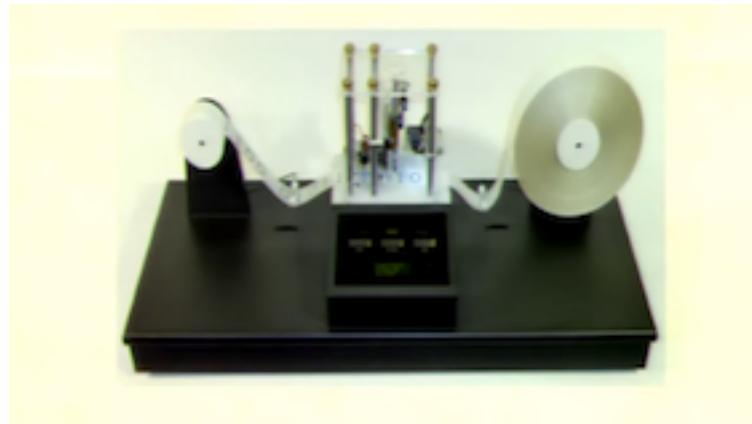


- Moore's Law: Computational power *doubles* roughly every two years.
→ Functional time algorithms will never be able to solve large problems.

- The practical implications are perhaps simpler to understand with this latter formulation.
- Thus, improvements in hardware will *never* overcome ***bad algorithm design***.
- Indeed, the current dramatic breakthroughs in computer science are based upon better (faster and higher performance) algorithmic techniques.

4 Robustness

- This measure of quality or "goodness" is ***robust***



- All reasonable models of algorithms are polynomial time equivalent.
 - Otherwise one model could perform, say, an exponential number of operations in the time another model took to perform just one.
- The standard formal model is the **Turing Machine**.

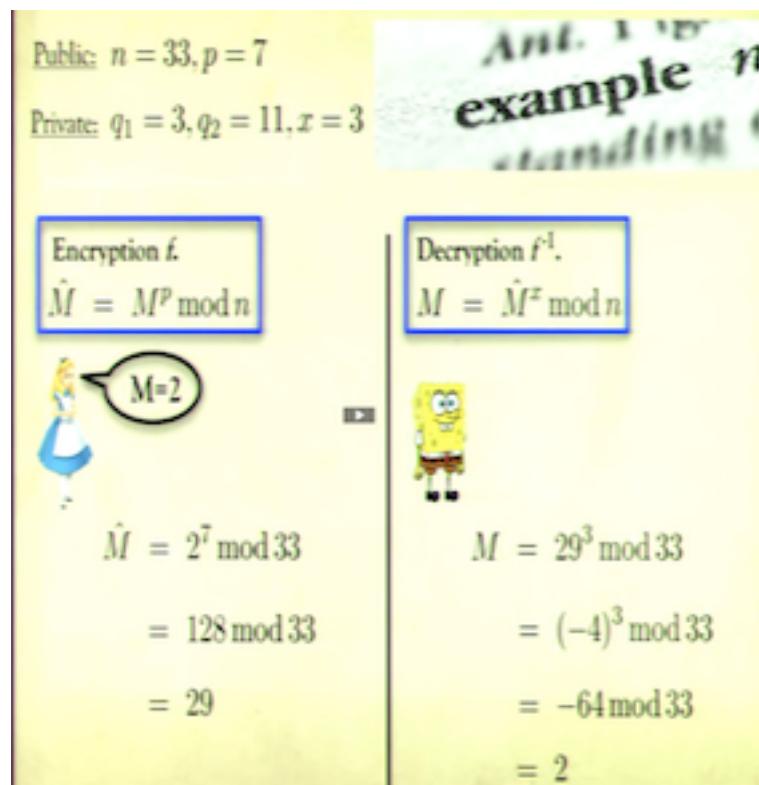
5 Cryptography (Just an example of the course)

- Alice wants to send Bob a message.
But she is worried that Eve might intercept the message.
She decides to encrypt the message M as $\hat{M} = f(M)$.
Bob can then decrypt the message via $f^{-1}(\hat{M}) = f^{-1}(f(M)) = M$
Eve cannot understand the message.

Alice encrypts the message M with the (encryption) lock f .
 Bob decrypts the message \hat{M} with the (decryption) key f^{-1} .
 Because Eve does not have the key she cannot decipher the message.
 Two major problems occur.

- Problem one:
 - Eve might be able to break the code.
 → That is, given \hat{M} she may be able to reconstruct f and then f^{-1} .
 - Standard techniques for code-breaking include:
 - Frequency Analysis
 - * e.g. "e" is the most common letter and "the" is the most common word in English.
 - Cribs.
 - * e.g. German weather reports were exploited to help decode messages from the "unbreakable" **Enigma Machine**.
- Problem Two:
 - Alice and Bob need to agree on what the encryption code (lock) f is.
 - But to do this they need to exchange a message discussing the code.
- Public-Key Cryptography
 - In fact, Bob gives absolutely everyone a copy of the (encryption) lock.
 - Everyone can send the message and Bob has the key.
 - The lock is made public. This is called public-key cryptography.
 - But this idea sounds completely crazy.
 - Doesn't this solution to Problem Two make Problem One inevitable?
 - That is, if Eve has the lock f then won't she use it to decode $f(M)$?
 - No, not if f is hard to invert for anybody except Bob himself.
 - But do such functions f that are hard to invert exist?
 - Yes!
- RSA Encryption

1. Bob chooses two large prime numbers q_1, q_2 and a large number p that is co-prime to $(q_1 - 1) \cdot (q_2 - 1)$
2. Bob's public key is (p, n) where $n = q_1 \cdot q_2$
Encryption: $\hat{M} = M^p \bmod n$
3. Bob's prime key is (q_1, q_2, x) where x is the inverse of p modular $(q_1 - 1) \cdot (q_2 - 1)$
Decryption: $M = \hat{M}^x \bmod n$
 - o $p = 2^{16} + 1$
 - o $q_1, q_2 = 2048$ bits
4. An Example
 - (a) Prime numbers q_1, q_2 and number p co-prime to $(q_1 - 1) \cdot (q_2 - 1)$
 $p = 7, q_1 = 3, q_2 = 11$ valid as $\gcd(7, 20) = 1$
 - (b) Public key (p, n) where $n = q_1 \cdot q_2$
 $n = 33$
 - (c) Private Key (q_1, q_2, x) where x is the inverse of p mod $(q_1 - 1) \cdot (q_2 - 1)$
 $x = 33$ as $3 \cdot 7 = 1 \bmod 20$
Public: $n = 33, p = 7$
Private $q_1 = 3, q_2 = 11, x = 3$



- Is RSA Encryption Safe?

- Public-key cryptography lies at the heart of the modern economy.
 - Financial Services
 - Online Shopping
 - Secure Messaging
- So it is extremely important that the method is safe.
- We claim it is because:
 - Bob has a good algorithm for decryption.
 - Eve only has a bad algorithm for decryption.

- Bob has a Good Decryption Algorithm

- Initially, Bob can do the following in polynomial time.
 - Choose the primes q_1, q_2
 - Choose a number p that is co-prime with $(q_1 - 1) \cdot (q_2 - 1)$

- Find x the inverse (Using Euclid's Algorithm) of p and $(q_1 - 1) \cdot (q_2 - 1)$
- Using fast exponentiation, encoding and decoding is polynomial time.
Encryption: $\hat{M} = M^p \bmod n$
- Eve has a Bad Decryption Algorithm
Decryption: $M = \hat{M}^x \bmod n$
- To decrypt Eve needs to find x the inverse of p and $(q_1 - 1) \cdot (q_2 - 1)$
 - She knows p , but does not know q_1, q_2
 - Instead, she only knows $n = q_1 \cdot q_2$
- So to find q_1, q_2 , she needs to find the prime factorization of n .
- But it is believed that finding the prime factors of a b -bit number is hard.
 - Instead, the obvious algorithm attempts to divide n by each integer in the range $\{2, 3, \dots, 2^b\} \leftarrow 4096$ bits
- In fact, if Eve could find x without q_1, q_2 , then she could use x to find q_1, q_2 . That is, she could factor n .
 - ⇒ Eve only has exponential time algorithm to decrypt.

Lecture 2: Recursive Algorithms: MergeSort; Binary Search; The Master Theorem; The Recursive Tree Method

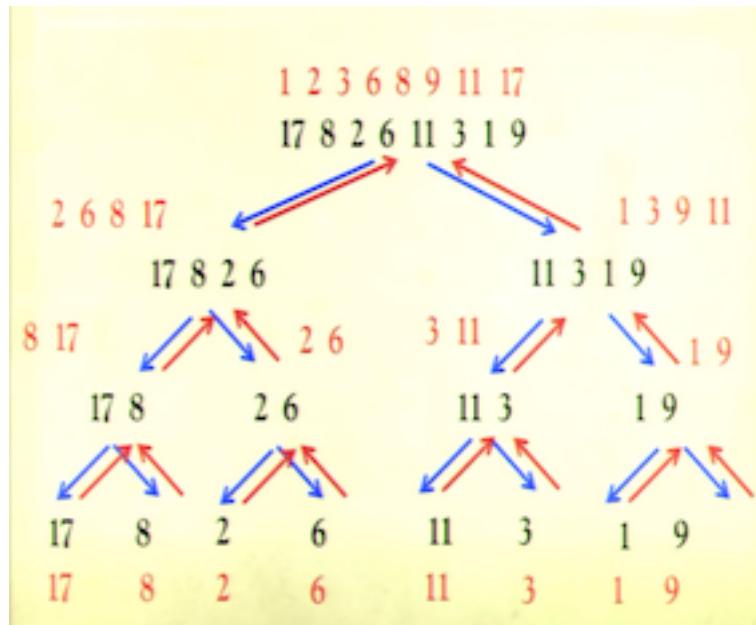
1 Reductions and Sub-Routines

- Solving a problem by **reducing** it (or a sub-problem of it) to another problem is the most fundamental technique in algorithm design.
- Specifically, algorithm A may use another algorithm B as a sub-routine.
- This has numerous advantages:
 - **Code Verification:** the correctness of A is independent of B .
 - **Code Reuse:** a great time-saver.
- A simple but very powerful special case of this paradigm is when the algorithm calls itself!
 - This method is called **recursion**.

2 MergeSort

- We can sort n numbers into non-decreasing order using the following algorithm:

```
MergeSort( $x_1, x_2, \dots, x_n$ )
If  $n = 1$  then output  $x_1$ 
Else output Merge{MergeSort( $x_1, \dots, x_{\frac{n}{2}}$ ), MergeSort( $x_{\frac{n}{2}+1}, \dots, x_n$ )}
```



- Two Problems:

- Does the algorithm work?

Yes!

→ The algorithm calls itself on smaller instances

★ The division process terminates with a set of base cases of size 1.

→ MergeSort trivially works on the base cases.

→ So, given the validity of the Merge Step, the correctness of the algorithm follows by **strong induction**.

★ As long as base case is correct and merge step works, everything will be fine.

- If so, is it efficient (polynomial time)?

Yes! Look at the recursive formula.

→ To analyze this we represent the running time $T(n)$ via a **recurrence**:

$$T(n) = 2 \cdot T\left(\frac{n}{2}\right) + c \cdot n$$

★ $2 \cdot T\left(\frac{n}{2}\right)$: Recurse on two problems with half the size.

★ $c \cdot n$: It takes linear time to merge two sorted lists.

Base Case: $T(1) = 1$

- ★ Or we can use $T(c) = O(1)$ for any constant c .

→ The Running Time of MergeSort

- ★ Theorem: MergeSort runs in time $O(n \cdot \log n)$

★ Proof:

1. By adding dummy numbers, we may assume n is a power of two: $n = 2^k$

2. We can unwind the recursive formula as follows:

$$T(n) = 2 \cdot T\left(\frac{n}{2}\right) + c \cdot n$$

Proof [cont.]

- But this unwinding operation can be repeated:

$$\begin{aligned} T(n) &= 2 \cdot T\left(\frac{n}{2}\right) + c \cdot n \\ &= 2 \cdot \left(2 \cdot T\left(\frac{n}{4}\right) + c \cdot \frac{n}{2}\right) + cn \\ &= 2^2 \cdot T\left(\frac{n}{4}\right) + 2 \cdot cn \\ &= 2^2 \cdot \left(2 \cdot T\left(\frac{n}{8}\right) + c \cdot \frac{n}{4}\right) + 2 \cdot cn \\ &= 2^3 \cdot T\left(\frac{n}{8}\right) + 3 \cdot cn \\ &= 2^3 \cdot \left(2 \cdot T\left(\frac{n}{16}\right) + c \cdot \frac{n}{8}\right) + 3 \cdot cn \\ &= 2^4 \cdot T\left(\frac{n}{16}\right) + 4 \cdot cn \\ &\quad \vdots \\ &= 2^k \cdot T(1) + k \cdot cn \end{aligned}$$

Since $n = 2^k$

Proof [cont.]

- Thus we have:

$$T(n) = 2^k \cdot T(1) + k \cdot cn$$

$$= n \cdot T(1) + k \cdot cn$$

$$= n \cdot (1 + k \cdot c)$$

- But c is a constant so:

$$T(n) = O(n \cdot k)$$

- Furthermore, $k = \log n$, so we get that:

$$T(n) = O(n \cdot \log n)$$



□

3 Binary Search

- We can search for a key k in a sorted array of cardinality n using the binary search algorithm:

BinarySearch($a_1, a_2, \dots, a_n : k$)

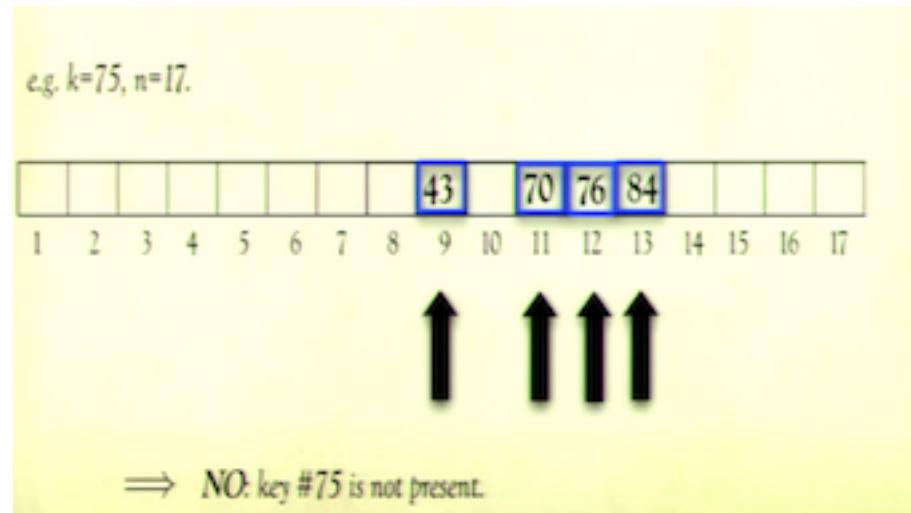
While $n > 0$ do:

If $a_{\frac{n}{2}} = k$ output YES

Else if $a_{\frac{n}{2}} > k$ output BinarySearch($a_1, a_2, \dots, a_{\frac{n}{2}-1} : k$)

Else if $a_{\frac{n}{2}} < k$ output BinarySearch($a_{\frac{n}{2}+1}, \dots, a_n : k$)

Output NO



- Does this work?
 - The validity of the binary search follows simply by strong induction.(The base case is trivially true.)
- Running Time?
 - Recurrence:

Recursive Formula: $T(n) = T\left(\frac{n}{2}\right) + c$
Base Case: $T(1) = 1$

 - Theorem: Binary Search runs in time $O(\log n)$
 1. By adding dummy numbers, we may assume n is a power of two: $n = 2^k$
 2. We can unwind the recursive formula as follows:

$$\begin{aligned}
 T(n) &= T\left(\frac{n}{2}\right) + c \\
 &= \left(T\left(\frac{n}{4}\right) + c\right) + c \\
 &= T\left(\frac{n}{4}\right) + 2 \cdot c
 \end{aligned}$$

- Again this unwinding operation can be repeated:

$$\begin{aligned}
 T(n) &= T\left(\frac{n}{4}\right) + 2 \cdot c \\
 &= \left(T\left(\frac{n}{8}\right) + c\right) + 2 \cdot c \\
 &= T\left(\frac{n}{8}\right) + 3 \cdot c \\
 &\quad \vdots \\
 &\quad \vdots \\
 &= T\left(\frac{n}{2^k}\right) + k \cdot c
 \end{aligned}$$

- Hence:

$$\begin{aligned}
 T(n) &= T(1) + k \cdot c \quad \text{--- Since } n = 2^k \\
 &= 1 + \log n \cdot c
 \end{aligned}$$

- This gives the claimed running time:

$$T(n) = O(\log n)$$

4 Divide and Conquer Algorithms

- A **divide and conquer** algorithm recursively breaks up a problem of size n in smaller sub-problems such that:
 - There are exactly a sub-problems.
 - Each sub-problem has size at most $\frac{1}{b} \cdot n$
 - Once solved, the solutions to the sub-problems can be combined to produce a solution to the original problem in time $O(n^d)$
- So the run-time of a divide and conquer algorithm satisfies the recurrence:

$$T(n) = a \cdot T\left(\frac{n}{b}\right) + O(n^d)$$

- MergeSort and Binary Search are indeed **divide and conquer** algorithms.

	Recursion Formula	a	b	d
MergeSort	$T(n) = 2 \cdot T\left(\frac{n}{2}\right) + O(n^1)$	2	2	1
Binary Search	$T(n) = 1 \cdot T\left(\frac{n}{2}\right) + O(n^0)$	1	2	0

5 Non-Military Applications of Divide and Conquer

- Divide and Conquer has many other non-military, practical applications:
 - Big Data
 - Distributed Algorithms
 - Clustering and Classification
 - MapReduce

6 Dummy Entries

- MergeSort actually has the recurrence:

$$\hat{T}(n) = \hat{T}(\lceil \frac{n}{2} \rceil) + \hat{T}(\lfloor \frac{n}{2} \rfloor) + c \cdot n$$

- Recall we got around this by adding dummy entries:

- We found \hat{n} the smallest power of 2 greater than n .
- For this case, MergeSort then does have recurrence:

$$T(n) = 2 \cdot T(\frac{n}{2}) + c \cdot n$$

- But we also have:

$$\hat{T}(n) \leq T(\bar{n}) = O(\bar{n} \cdot \log \bar{n}) = O(n \cdot \log n)$$

- Here is another way to solve the recurrence:

$$\hat{T}(n) = \hat{T}(\lceil \frac{n}{2} \rceil) + \hat{T}(\lfloor \frac{n}{2} \rfloor) + c \cdot n$$

- As we only want to upper bound the running time, we can use:

$$\hat{T}(n) \leq 2 \cdot \hat{T}(\frac{n}{2} + 1) + c \cdot n$$

Note: This $+1$ does not seem to fit with our methodology, but we can fix this by applying a **domain transformation**.

- Domain Transformation

- For the domain transformation, simply set: $T(n) = \hat{T}(n + 2)$
- Thus we have: $T(n) = 2 \cdot T(\frac{n}{2}) + \hat{c} \cdot n$

$$\begin{aligned}
 T(n) &= \hat{T}(n+2) \\
 &\leq \hat{T}\left(\frac{n+2}{2} + 1\right) + c \cdot (n+2) \\
 &\leq \hat{T}\left(\frac{n+2}{2} + 1\right) + \hat{c} \cdot n \\
 &= \hat{T}\left(\frac{n}{2} + 2\right) + \hat{c} \cdot n \\
 &= T\left(\frac{n}{2}\right) + \hat{c} \cdot n
 \end{aligned}$$

Note: missing a 2 in front of \hat{T}

- o Of course, we can solve this recurrence as: $T(n) = O(n \cdot \log n)$
- o Therefore, $\hat{T}(n) = T(n - 2) = O(n \cdot \log n)$
- o As well as ceilings and floors, domain transformations can be used to simplify many other recurrences; e.g. removing lower order terms.

Lecture 3: The Master Theorem; The Recursive Tree Method

1 Quick Review

- Recall, a divide-and-conquer algorithm recursively breaks up a problem of size n in smaller sub-problems such that:
 - There are exactly a sub-problems.
 - Each sub-problem has size at most $\frac{1}{b} \cdot n$
 - Once solved, the solutions to the sub-problems can be combined to produce a solution to the original problem in time $O(n^d)$
- So the run-time of a divide and conquer algorithm satisfies the recurrence:
$$T(n) = a \cdot T\left(\frac{n}{b}\right) + O(n^d)$$
- Examples: MergeSort and Binary Search

2 The Master Theorem

- **The Master Theorem:** If $T(n) = a \cdot T\left(\frac{n}{b}\right) + O(n^d)$ for constants $a > 0$, $b > 1$, and $d \geq 0$ then:

$$T(n) = \begin{cases} O(n^d) & \text{if } a < b^d \text{ [Case I]} \\ O(n^d \cdot \log n) & \text{if } a = b^d \text{ [Case II]} \\ O(n^{\log_b a}) & \text{if } a > b^d \text{ [Case III]} \end{cases}$$

- Sanity Check: What does this give for MergeSort and Binary Search?

	a	b	d	Case	Runtime
MergeSort	2	2	1	II	$O(n \cdot \log n)$
Binary Search	1	2	0	II	$O(\log n)$

3 Proof of The Master Theorem

- Fact One:

Fact 1. $\sum_{k=0}^{\ell} \tau^k = \frac{1 - \tau^{\ell+1}}{1 - \tau}$ for any $\tau \neq 1$.

Proof.

- We have:

$$\begin{aligned} (1 - \tau) \cdot \sum_{k=0}^{\ell} \tau^k &= \sum_{k=0}^{\ell} \tau^k - \sum_{k=1}^{\ell+1} \tau^k \\ &= \tau^0 - \tau^{\ell+1} \\ &= 1 - \tau^{\ell+1} \end{aligned}$$

- Dividing both sides by $(1 - \tau)$ gives the result.

- Fact Two:

Fact 2. $x^{\log_b y} = y^{\log_b x}$ for any base b .

Proof.

$$\log_b z^p = p \cdot \log_b z$$

- Observe that, by the *power rule of logarithms*, we have:

$$\log_b x \cdot \log_b y = \log_b(y^{\log_b x})$$

- Similarly:

$$\log_b x \cdot \log_b y = \log_b(x^{\log_b y})$$

- Putting this together gives

$$\log_b(y^{\log_b x}) = \log_b(x^{\log_b y})$$

$$\Rightarrow x^{\log_b y} = y^{\log_b x}$$

- Proof of the Master Theorem

Proof.

- We may assume n is a power of b : $n = b^\ell$

- So we have:

$$\begin{aligned} T(n) &= n^d + a \cdot \left(\frac{n}{b}\right)^d + a^2 \cdot \left(\frac{n}{b^2}\right)^d + \dots + a^\ell \cdot \left(\frac{n}{b^\ell}\right)^d \\ &= n^d \cdot \left(1 + a \cdot \left(\frac{1}{b}\right)^d + a^2 \cdot \left(\frac{1}{b^2}\right)^d + \dots + a^\ell \cdot \left(\frac{1}{b^\ell}\right)^d\right) \\ &= n^d \cdot \left(1 + \frac{a}{b^d} + \left(\frac{a}{b^d}\right)^2 + \dots + \left(\frac{a}{b^d}\right)^\ell\right) \end{aligned}$$

Proof [cont.] $T(n) = n^d \cdot \left(1 + \frac{a}{b^d} + \left(\frac{a}{b^d}\right)^2 + \cdots + \left(\frac{a}{b^d}\right)^\ell \right)$

Case 1: $\frac{a}{b^d} < 1$

- Set $\tau = \frac{a}{b^d}$
- Then: $T(n) = n^d \cdot \sum_{k=0}^{\ell} \tau^k$
- Applying Fact 1, we know that:

$$\sum_{k=0}^{\ell} \tau^k = \frac{1 - \tau^{\ell+1}}{1 - \tau} \leq \frac{1}{1 - \tau} = O(1)$$

As a , b , and d are constants so is $1 - \tau$.

- Therefore:

$$T(n) \leq n^d \cdot \frac{1}{1 - \frac{a}{b^d}} = n^d \cdot \frac{b^d}{b^d - a} = O(n^d)$$

Proof [cont.] $T(n) = n^d \cdot \left(1 + \frac{a}{b^d} + \left(\frac{a}{b^d}\right)^2 + \cdots + \left(\frac{a}{b^d}\right)^\ell \right)$

Case II: $\frac{a}{b^d} = 1$

- Then: $T(n) = n^d \cdot (\ell + 1)$
- But $n = b^\ell$ so $\ell = \log_b n$.
- As b is a constant greater than one, this gives: $T(n) = O(n^d \cdot \log n)$

Case III: $\frac{a}{b^d} > 1$

- Again set $\tau = \frac{a}{b^d}$
 - Then: $T(n) = n^d \cdot \sum_{k=0}^{\ell} \tau^k$
 - Applying Fact 1 gives:
- $$\sum_{k=0}^{\ell} \tau^k = \frac{\tau^{\ell+1} - 1}{\tau - 1} \leq \frac{\tau^{\ell+1}}{\tau - 1} = O(\tau^{\ell+1}) = O(\tau^\ell)$$
- As a , b , and d are constants so is τ .

Proof [cont.]

- Thus: $T(n) = O(n^d \cdot \tau^\ell)$

- Observe that:

$$n^d \cdot \tau^\ell = n^d \cdot \left(\frac{a}{b^d}\right)^\ell$$

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

$$= 1 \cdot a^\ell \quad \longrightarrow \boxed{\text{As } n=b^\ell.}$$

$$= a^{\log_b n}$$

$$= n^{\log_b a} \quad \longrightarrow \boxed{\text{By Fact 2.}}$$

- This gives the final case:

$$T(n) = O(n^{\log_b a})$$

- Specifically, what matters is **not** the statement of the Master Theorem but the **ideas** underlying its proof.
 - First, if we understand the proof then we can easily reconstruct the theorem.
 - Second, if we understand the proof then we can easily apply the method to a much broader class of problems. For example:
 - The sub-problems have different sizes.

e.g. The Deterministic Selection Algorithm.

$$T(n) \leq T\left(\frac{7n}{10}\right) + T\left(\frac{n}{5}\right) + O(n)$$

→ The combination function is not of the form $f(n) = n^d$.

e.g. Euclid's Greatest Common Divisor Algorithm.

$$T(n) = T\left(\frac{n}{2}\right) + O(\log n)$$

→ The parameters a , b , and d are not constants.

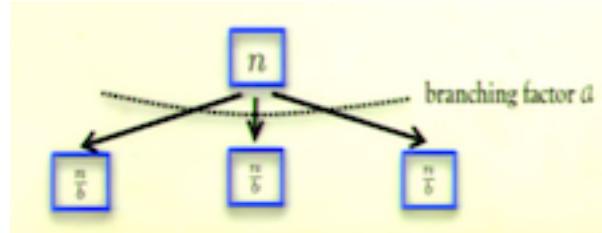
$$\text{e.g. } T(n) = \sqrt{n} \cdot T(\sqrt{n}) + O(n^{\frac{1}{\log \log n}})$$

4 The Recursion Tree Method

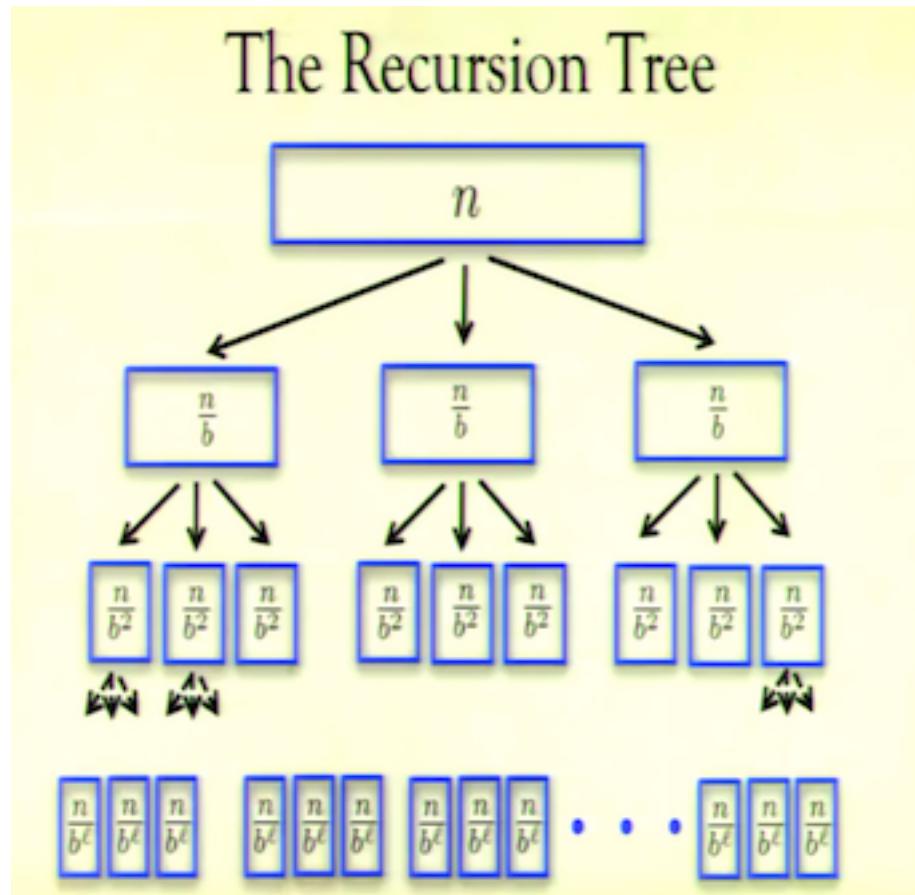
- The Master Theorem is a special case of the **recursion tree method**.
- Specifically, we model the *divide and conquer* recursive formula by a tree:

$$T(n) = a \cdot T\left(\frac{n}{b}\right) + O(n^d)$$

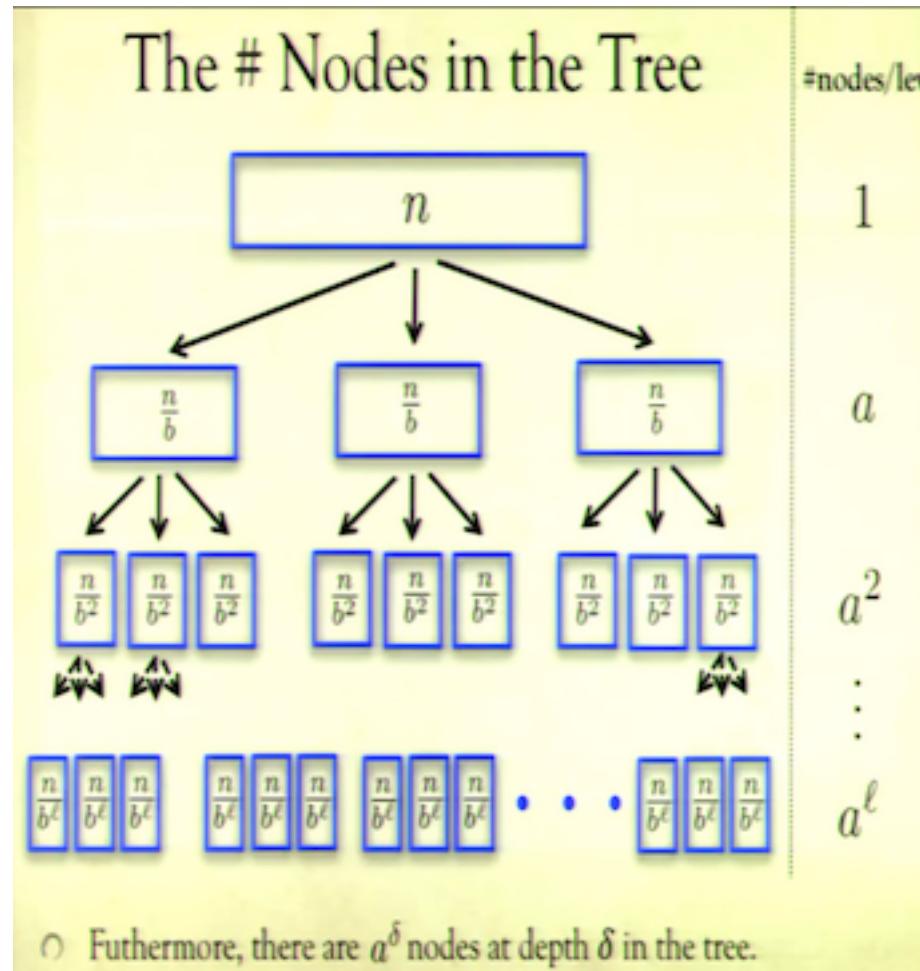
- The **root** node of the tree has a label n .
- The root has a children each with label $\frac{n}{b}$.



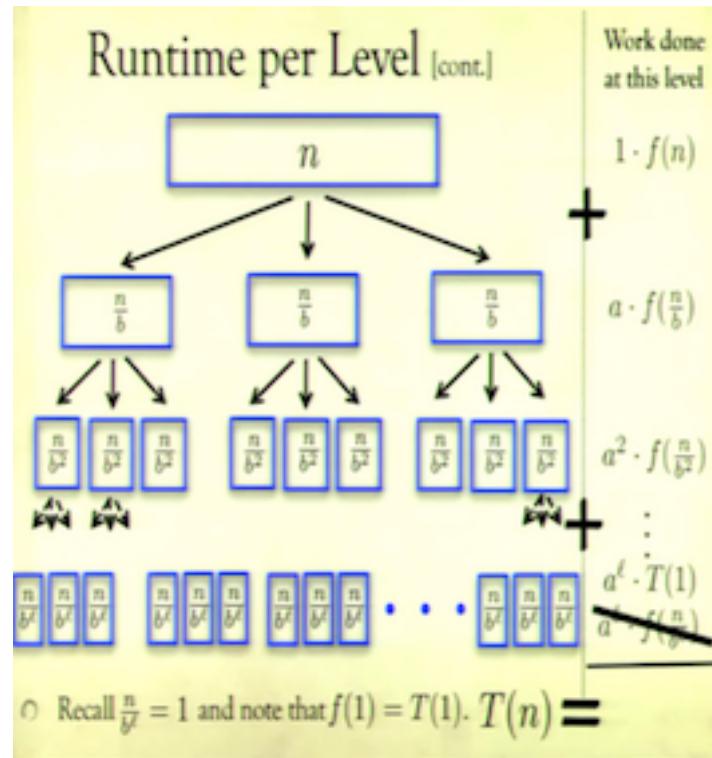
- This pattern then repeats at the children, then grandchildren, etc.

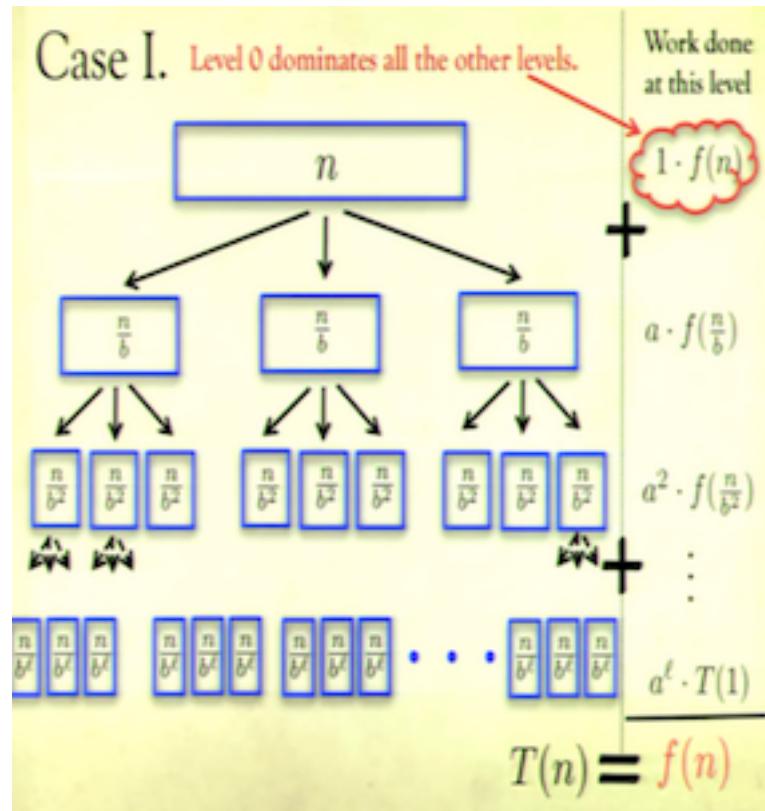


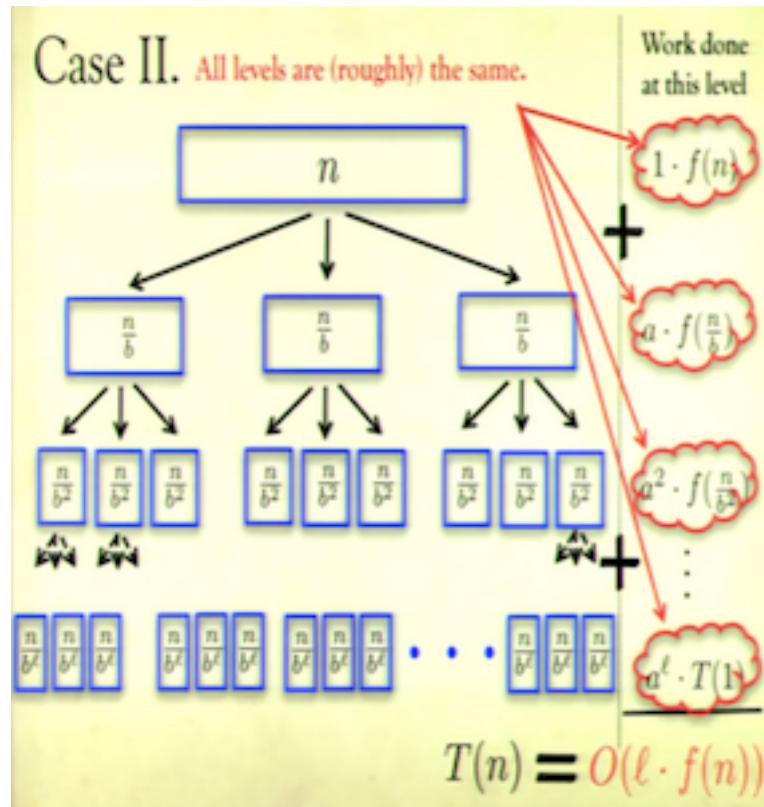
- This process stops at the **leaves** (base cases) which have label $\frac{n}{b^l} = 1$. (As $n = b^l$)

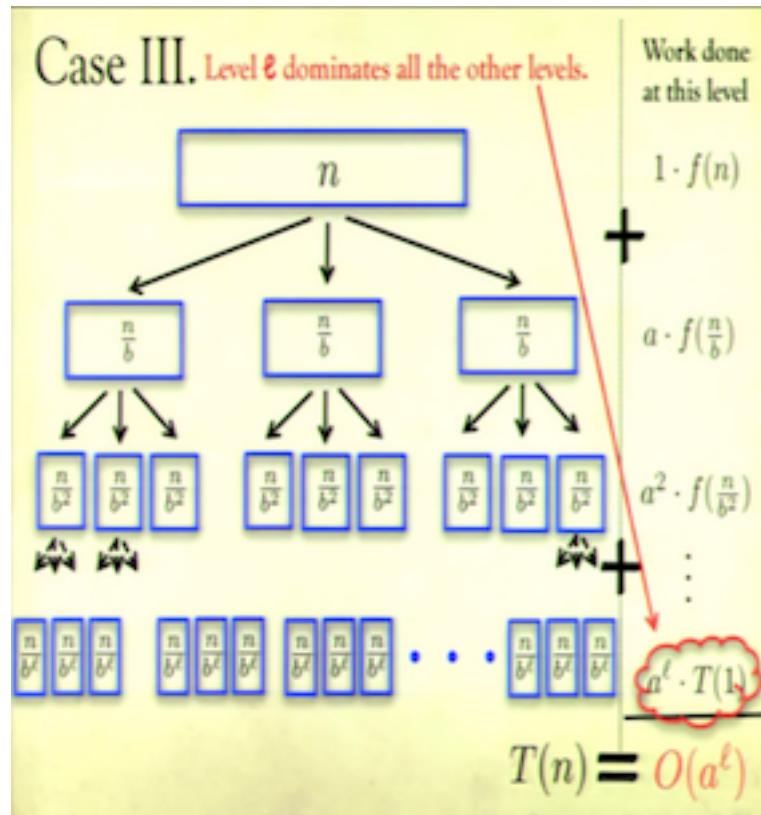


- How much time do we spend at each level?

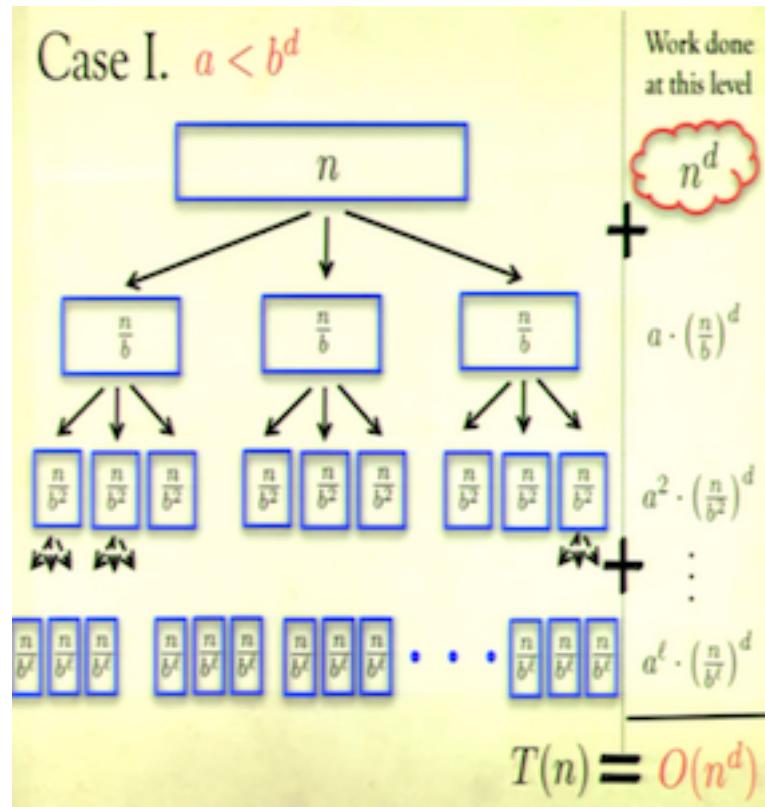


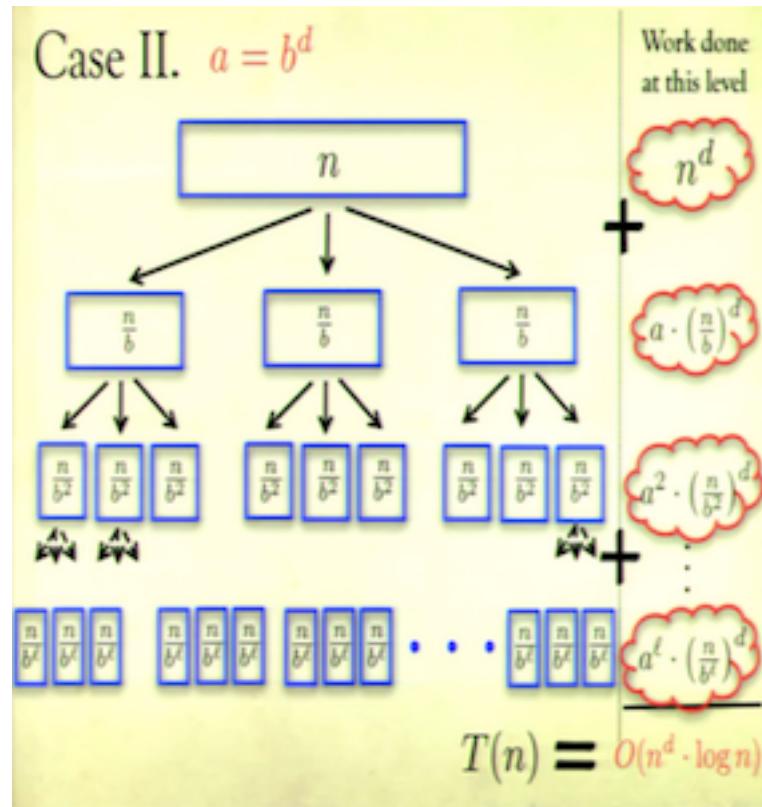


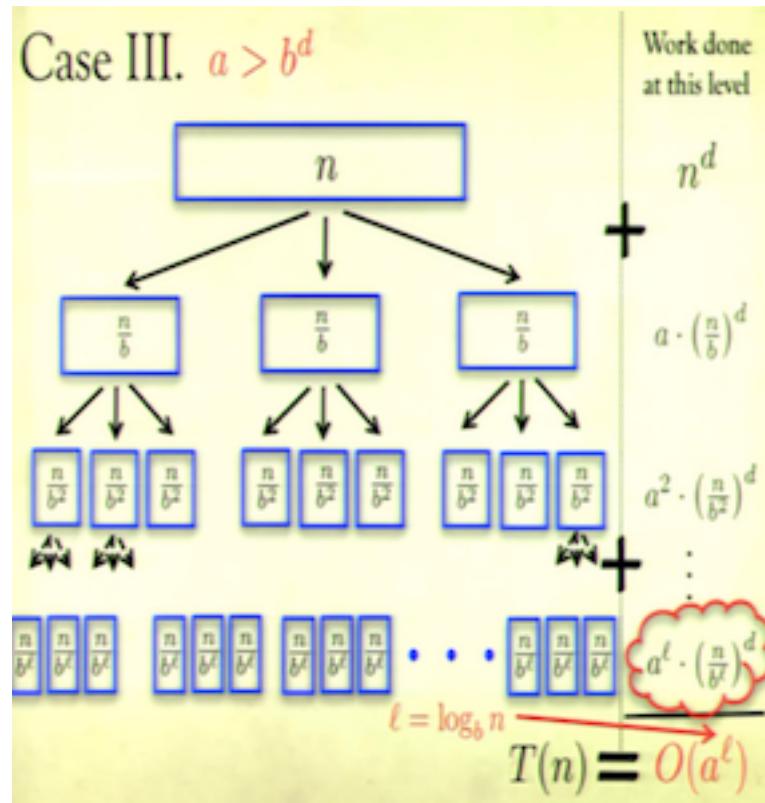




- This gives us the proof of the Master Theorem:







Lecture 4: Recursive Algorithms: Fast Multiplication; Fast Matrix Multiplication

1 Multiplication

- Grade School Multiplication
 - Perform n^2 multiplications to multiply two n-digit numbers
→ Long multiplication has running time $\Omega(n^2)$.
- Russian Peasant Multiplication
 $\text{Mult}(x, y)$
If $x = 1$ **then** output y
If x is odd **then** output $y + \text{Mult}([\frac{x}{2}], 2y)$
If x is even **then** output $\text{Mult}(\frac{x}{2}, 2y)$

Binary Representation	x	y
1 0	46	324
2 1	23	648
4 1	11	1296
8 1	5	2592
16 0	2	5184
32 1	1	10368
46		14904

- This method does work. Base case ($x = 1$) is verified. The step when x is even also works. The only tricky step is that x is odd when you divide x by 2, you actually divide x by 2 and minus a half, so that you need to add one y back.
- How long does it take?

$\# \text{Steps} = \# \text{Bits}(x)$

Binary Representation	x	y
1 0	46	324 ← $\# \text{Bits}(y)$
2 1	23	648
4 1	11	1296
8 1	5	2592
16 0	2	5184
32 1	1	10366 ← $\# \text{Bits}(y) + \# \text{Bits}(x)$
		14904

- There are n iterations, so we add up to n numbers with at most $2n$ digits each.
 → Runtime = $O(n^2)$

2 Divide and Conquer Multiplication

○ Let:

$$\begin{aligned} \mathbf{x} &= x_n x_{n-1} \cdots x_{\frac{n}{2}+1} \boxed{x_{\frac{n}{2}} \cdots x_2 x_1} & \text{e.g. } \mathbf{x} &= 4132 \\ \mathbf{y} &= y_n y_{n-1} \cdots y_{\frac{n}{2}+1} \boxed{y_{\frac{n}{2}} \cdots y_2 y_1} & \mathbf{y} &= 6703 \\ && \mathbf{y}_L && \mathbf{y}_R \end{aligned}$$

$$\begin{aligned} \text{○ Then: } \mathbf{x} &= 10^{\frac{n}{2}} \cdot \mathbf{x}_L + \mathbf{x}_R & \mathbf{x} &= 41 \cdot 10^2 + 32 \\ \mathbf{y} &= 10^{\frac{n}{2}} \cdot \mathbf{y}_L + \mathbf{y}_R & \mathbf{y} &= 67 \cdot 10^2 + 03 \end{aligned}$$

○ Thus:

$$\begin{aligned} \mathbf{x} \cdot \mathbf{y} &= (10^{\frac{n}{2}} \cdot \mathbf{x}_L + \mathbf{x}_R) \cdot (10^{\frac{n}{2}} \cdot \mathbf{y}_L + \mathbf{y}_R) \\ &= 10^n \cdot \mathbf{x}_L \mathbf{y}_R + 10^{\frac{n}{2}} \cdot (\mathbf{x}_L \mathbf{y}_R + \mathbf{x}_R \mathbf{y}_L) + \mathbf{x}_R \mathbf{y}_R \\ \implies \mathbf{x} \cdot \mathbf{y} &= 10^n \cdot \mathbf{x}_L \mathbf{y}_R + 10^{\frac{n}{2}} \cdot (\mathbf{x}_L \mathbf{y}_R + \mathbf{x}_R \mathbf{y}_L) + \mathbf{x}_R \mathbf{y}_R \end{aligned}$$

Note: There is one typo in Prof's lecture slide. It should be $10^n \cdot x_L y_L$.

- How long does this algorithm take?

- The recursive formula for the running time is:

$$T(n) = 4 \cdot T\left(\frac{n}{2}\right) + O(n)$$

Note: by padding some zeroes in front of the number, we can assume n is a power of 2.

- Thus we have a = 4, b = 2, and d = 1.
- This is Case 3 of the Master Theorem.
- Runtime = $O(n^{\log_2 4}) = O(n^2)$

3 Gauss's Complex Number Multiplication

- Gauss considered the product of complex numbers:

$$(a + b \cdot i) \cdot (c + d \cdot i) = ac - bd + (bc + ad) \cdot i$$

- This seems to require taking **four products**, but he observed that:

$$(bc + ad) = (a + b) \cdot (c + d) - ac - bd$$

- So we can calculate ac and bd then we only need to perform only one more product to find $(bc + ad)$, namely $(a + b) \cdot (c + d)$.
 - Multiplying two complex numbers involves only **three products!** We got extra addition here, but addition is much cheaper than multiplication when we think about the running time.

4 Application of Gauss's Complex Number Multiplication

- We can simply replace i by $10^{\frac{n}{2}}$

$$(a + b \cdot i) \cdot (c + d \cdot i) = ac - bd + (bc + ad) \cdot i$$

$$(x_R + x_L \cdot 10^{\frac{n}{2}}) \cdot (y_R + y_L \cdot 10^{\frac{n}{2}}) = x_R y_R + x_L y_L \cdot 10^n + (x_L y_R + x_R y_L) \cdot 10^{\frac{n}{2}}$$

◦ But now we have:

$$\cancel{(x_L y_R + x_R y_L)} = (x_R + x_L) \cdot (y_R + y_L) - \cancel{x_R y_R} - \cancel{x_L y_L}$$

$$(x_L y_R + x_R y_L) = x_R y_R + x_L y_L - (x_R - x_L) \cdot (y_R - y_L)$$

- Multiplying involves only **three products** with $\frac{n}{2}$ digit numbers!
 - The recursive formula for the running time is:
- $$T(n) = 3 \cdot T\left(\frac{n}{2}\right) + O(n)$$
- Thus we have $a = 3$, $b = 2$, and $d = 1$.

- This is Case 3 of the Master Theorem.
- Runtime = $O(n^{\log_2 3}) = O(n^{1.59})$
- So we can multiply two n-bit numbers using less than n^2 operations!

5 Fast Fourier Transforms

- We can multiply two n-bit numbers in time $O(n \cdot \log n)$ using a **Fast Fourier Transform**.
- More generally, FFTs can be used to multiply two **polynomial functions**.
 - This has a vast number of applications, for example in *image compression* and *signal processing*.
 - Indeed, it has been described by Strang as "the most important numerical algorithm of our lifetime."
- We might study FFTs later in the course.

6 Matrix Multiplication

- High School Matrix Multiplication

○ Let:

$$X = \begin{pmatrix} x_{11} & x_{12} & \dots & x_{1n} \\ x_{21} & x_{22} & \dots & x_{2n} \\ \vdots & \ddots & & \vdots \\ x_{n1} & x_{n2} & \dots & x_{nn} \end{pmatrix} Y = \begin{pmatrix} y_{11} & y_{12} & \dots & y_{1n} \\ y_{21} & y_{22} & \dots & y_{2n} \\ \vdots & \ddots & & \vdots \\ y_{n1} & y_{n2} & \dots & y_{nn} \end{pmatrix}$$

○ Then:

$$Z = X \cdot Y = \begin{pmatrix} z_{11} & z_{12} & \dots & z_{1n} \\ z_{21} & z_{22} & \dots & z_{2n} \\ \vdots & \ddots & & \vdots \\ z_{n1} & z_{n2} & \dots & z_{nn} \end{pmatrix} \quad \text{where } z_{ij} = \sum_{k=1}^n x_{ik} \cdot y_{kj}$$

Then we perform n multiplications to calculate each entry of Z .

- Runtime = $\Omega(n^3)$

7 Divide and Conquer Matrix Multiplication

- We can also multiply matrices using **divide and conquer**.

$$\begin{aligned} \circ & \text{ Let: } X = \begin{pmatrix} A & B \\ C & D \end{pmatrix} \quad Y = \begin{pmatrix} E & F \\ G & H \end{pmatrix} \\ \circ & \text{ Then: } Z = \begin{pmatrix} A & B \\ C & D \end{pmatrix} \cdot \begin{pmatrix} E & F \\ G & H \end{pmatrix} = \begin{pmatrix} AE + BG & AF + BH \\ CE + DG & CF + DH \end{pmatrix} \end{aligned}$$

- Multiplying involves **eight products** with $\frac{n}{2} \times \frac{n}{2}$ matrices!

- The recursive formula for the running time is:

$$T(n) = 8 \cdot T\left(\frac{n}{2}\right) + O(n^2)$$

- Thus we have $a = 8$, $b = 2$, and $d = 2$.

- This is Case 3 of the Master Theorem.

- Runtime = $O(n^{\log_2 8}) = O(n^3)$

- No improvement!

8 An Algebraic Trick for Matrix Multiplication

○ Let:

$$\begin{aligned}
 Z &= \begin{pmatrix} A & B \\ C & D \end{pmatrix} \cdot \begin{pmatrix} E & F \\ G & H \end{pmatrix} && \text{where } S_1 = (B - D) \cdot (G + H) \\
 &= \begin{pmatrix} AE + BG & AF + BH \\ CE + DG & CF + DH \end{pmatrix} && S_2 = (A + D) \cdot (E + H) \\
 &= \begin{pmatrix} S_1 + S_2 - S_4 + S_6 & S_4 + S_5 \\ S_6 + S_7 & S_2 - S_3 + S_5 - S_7 \end{pmatrix} && S_3 = (A - C) \cdot (E + F) \\
 &&& S_4 = (A + B) \cdot H \\
 &&& S_5 = A \cdot (F - H) \\
 &&& S_6 = D \cdot (G - E) \\
 &&& S_7 = (C + D) \cdot E
 \end{aligned}$$

- Multiplying involves **seven products** with $\frac{n}{2} \times \frac{n}{2}$ matrices!
Note: Even though we now have 18 additions, the runtime is still dominated by the number of multiplications. Having additions would not hurt.
- The recursive formula for the running time is:

$$T(n) = 7 \cdot T\left(\frac{n}{2}\right) + O(n^2)$$
- Thus we have $a = 7$, $b = 2$, and $d = 2$.
- This is Case 3 of the Master Theorem.
- Runtime = $O(n^{\log_2 7}) = O(n^{2.81})$
- This divide and conquer matrix multiplication algorithm was designed by Strassen (1969). Since then faster algorithms (in theory) have been developed:
 - Theorem:** There is a matrix multiplication algorithm that runs in time $O(n^{2.37})$
 - Open Problem:** Is matrix multiplication $O(n^{2+\epsilon})$ time for any constant $\epsilon > 0$? Note: There is no way you can beat this because just to read the matrix x you have to do x^2 amount of work to read entries in x simply for y . Getting running time like this seems crazy but possible.

9 Fast Exponentiation

- Suppose we want to compute x^n
 - The slow way $x \cdot x \cdot \dots \cdot x$ requires $n-1$ multiplications.
 - A faster way is to use the fact that $x^n = x^{[\frac{n}{2}]} \cdot x^{[\frac{n}{2}]}$
- The latter method gives the following **recursive algorithm**.

Fast Exponentiation

FastExp(x,n)

If $n = 1$ **then** output x

Else

If n is even **then** output **FastExp**($x, [\frac{n}{2}]$) 2

If n is odd **then** output $x \cdot \text{FastExp}(x, [\frac{n}{2}])^2$

- The number of multiplications used in Fast Exponentiation satisfies:

$$T(n) \leq T([\frac{n}{2}]) + 2 \Rightarrow T(n) = T(\frac{n}{2}) + O(1)$$

- Thus we have $a = 1$, $b = 2$, and $d = 0$.
- This is Case 2 of the Master Theorem.
- Runtime = $O(n^0 \cdot \log n) = O(\log n)$

Lecture 5: Recursive Algorithms: The Median; Randomized Selection; Deterministic Selection

1 The Median Problem

- How do we find the **median** of a set $S = \{x_1, x_2, \dots, x_n\}$?
 - We could sort the list and then output the $\lceil \frac{n}{2} \rceil$ th number.
 - Using **Mergesort**, or otherwise, this will take time $O(n \cdot \log n)$.
- Is there a **faster** way to find the median?
 - We only need the median number. Sorting all numbers seems overkill. We cannot do better than $O(n)$ since we need to read n numbers in linear time.

2 The Selection Problem

- How do we find the k^{th} smallest number in $S = \{x_1, x_2, \dots, x_n\}$?
 - Again, we could sort the list and then output the k^{th} number.
 - Using **Mergesort**, this takes time $O(n \cdot \log n)$.
- We can do this much faster using recursion...

The Selection Algorithm

Select(S , k)

If $|S| = 1$ then output x_1

Else

Set $S_L = \{x_i \in S : x_i < x_1\}$

Set $S_R = \{x_i \in S \setminus x_1 : x_i \geq x_1\}$

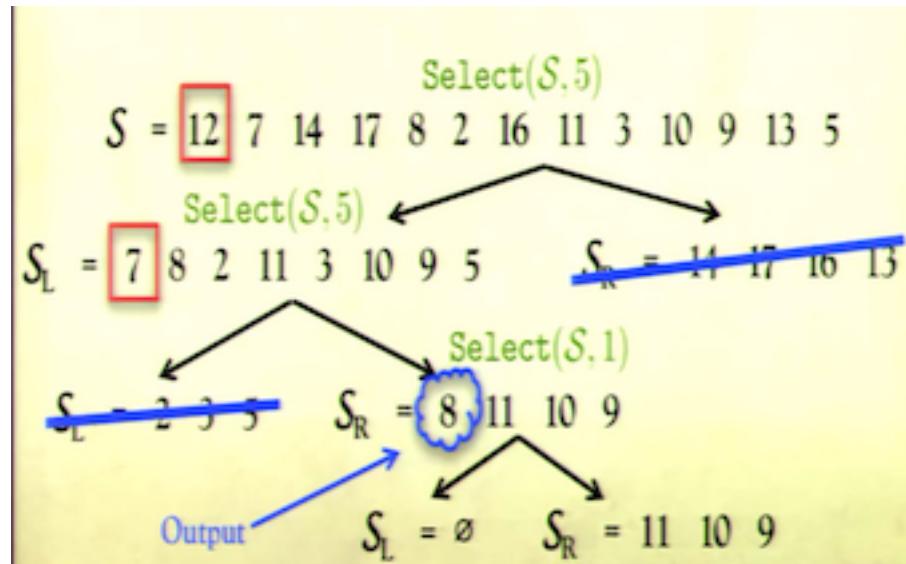
If $|S_L| = k - 1$ then output x_1

If $|S_L| > k - 1$ then output **Select(S_L , k)**

If $|S_L| < k - 1$ then output **Select(S_R , $k - 1 - |S_L|$)**

Note:

1. If $|S_L| = k - 1$, then x_1 is the k^{th} smallest number.
2. On the other hand, suppose smallest S_L contains at least $(k - 1)$ elements, which means it gets k elements, in other words, k^{th} smallest element is actually in S_L . The k^{th} smallest element of S must also be the k^{th} smallest element of S_L .
3. S_L union x_1 has the most $(k - 1)$ elements, which means the k^{th} smallest element must be in the set S_R . Since everything in S_R is at least bigger than in x_1 , so that it is also bigger than in S_L , to find the k^{th} smallest element, we need $(k - 1 - |S_L|)$.



- How many comparisons $T(n)$ does this recursive selection algorithm make?

```

The Selection Algorithm       $T(n) = n - 1 + T(\max\{|\mathcal{S}_L|, |\mathcal{S}_R|\})$ 
Select( $\mathcal{S}, k$ )            $= n - 1 + T(n - 1)$ 

If  $|\mathcal{S}| = 1$  then output  $x_1$ 

Else
  Set  $\mathcal{S}_L = \{x_i \in \mathcal{S} : x_i < x_1\}$ 
  Set  $\mathcal{S}_R = \{x_i \in \mathcal{S} \setminus x_1 : x_i \geq x_1\}$ 
  If  $|\mathcal{S}_L| = k - 1$  then output  $x_1$             $n-1$  comparisons
  If  $|\mathcal{S}_L| > k - 1$  then output Select( $\mathcal{S}_L, k$ )
  If  $|\mathcal{S}_L| < k - 1$  then output Select( $\mathcal{S}_R, k - 1 - |\mathcal{S}_L|$ )

```

- So in the **worst case** we have is:

$$\begin{aligned}
T(n) &= (n - 1) + T(n - 1) \\
&= (n - 1) + (n - 2) + T(n - 2) \\
&\quad \cdot \\
&\quad \cdot \\
&\quad \cdot \\
&= (n - 1) + (n - 2) + (n - 3) + \cdots + 2 + 1 \\
&= \frac{1}{2}n(n - 1) \\
&= \Omega(n^2)
\end{aligned}$$

- This is terrible - sorting would have been faster!
- How to fix it? Use Balanced Pivots!
 - The problem is the algorithm repeatedly **pivots** on the first number in the current list.
→ But if we are unlucky this pivot could be very **unbalanced**. That is:
 $\max\{|\mathcal{S}_L|, |\mathcal{S}_R|\} \approx n$
 $\min\{|\mathcal{S}_L|, |\mathcal{S}_R|\} \approx 0$

- What we would like is to select a pivot that is **balanced**. That is:

$$\max\{|S_L|, |S_R|\} \approx \frac{n}{2}$$

$$\min\{|S_L|, |S_R|\} \approx \frac{n}{2}$$

- Randomization

- The current algorithm is **deterministic** in the choice of the pivot.
- To fix the problem we consider a **randomized** implementation.
 - Do not pivot deterministically on x_1 .
 - Instead choose the pivot at random from $S = \{x_1, x_2, \dots, x_n\}$

Randomized Selection

RandSelect(S, k)

If $|S| = 1$ then output x_1
 Else

Pick a random pivot $x_\tau \in \{x_1, x_2, \dots, x_n\}$

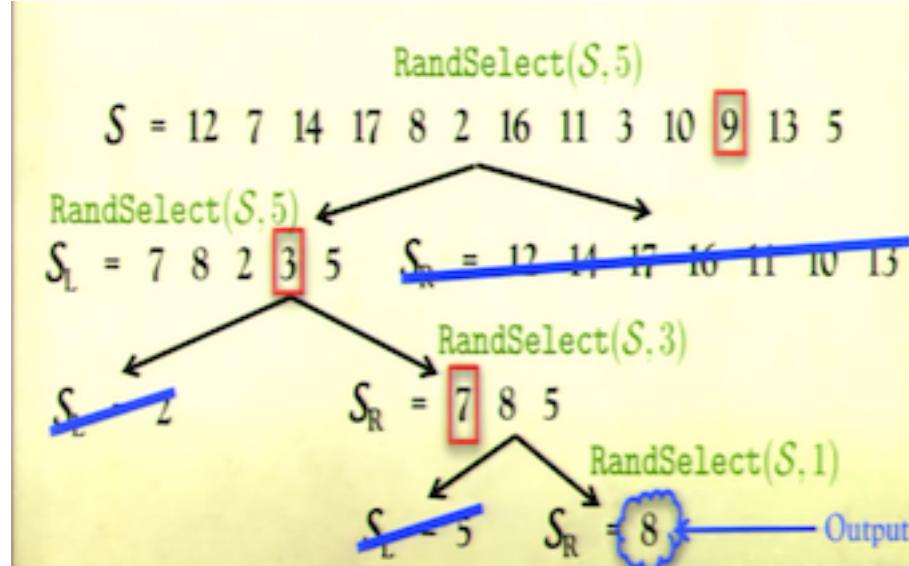
Set $S_L = \{x_i \in S : x_i < x_\tau\}$

Set $S_R = \{x_i \in S \setminus x_\tau : x_i \geq x_\tau\}$

If $|S_L| = k - 1$ then output x_τ

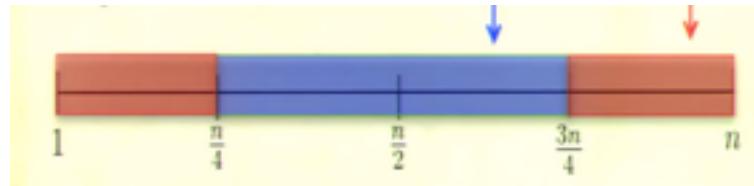
If $|S_L| > k - 1$ then output **RandSelect**(S_L, k)

If $|S_L| < k - 1$ then output **RandSelect**($S_R, k - 1 - |S_L|$)



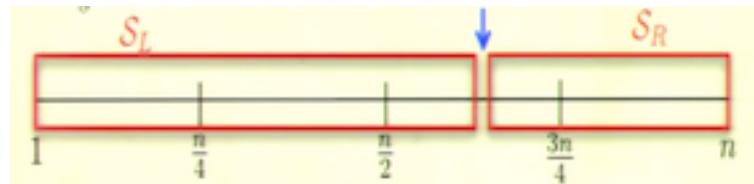
- Good vs Bad Pivots

- Imagine all the numbers in the **sorted** order.



- With probability $\frac{1}{2}$ the pivot x_τ lies between the 1st and 3rd quartiles.
→ We say that such a pivot is **good**.
→ We say that such a pivot is **bad**.
- **Key Observation:** If the pivot is good then

$$\max\{|S_L|, |S_R|\} \leq \frac{3}{4} \cdot n$$



- Expected Runtime

- In the worst case, the *randomized algorithm* will pick the worst pivot! The probability of its happening is very small.
- So, for randomized algorithms, we are always interested in the **expected runtime** $\bar{T}(n) = E(T(n))$, not the worst case run time.
- Using our observation, we then have that:

$$\bar{T}(n) \leq \frac{1}{2} \cdot \bar{T}\left(\frac{3n}{4}\right) + \frac{1}{2} \cdot \bar{T}(n) + O(n)$$

Note:

1. The first term: the probability of $\frac{1}{2}$ comes from I make a good pivot.
If I make a good pivot, I know both of my subsets will have at most the size of $\frac{3n}{4}$.

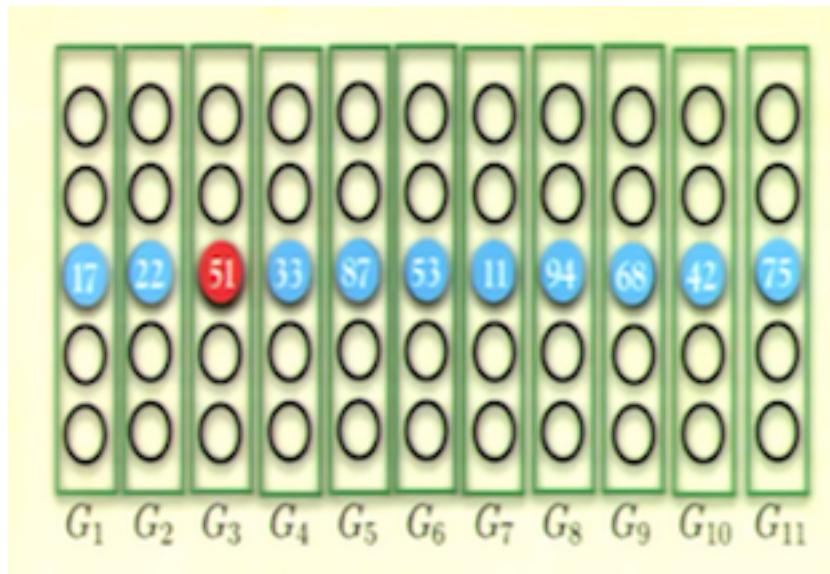
2. The second term: If I get a bad pivot, the size of the next problem might be $(n - 1)$, which is $\bar{T}(n)$ in the worst case.

$$\begin{aligned}\bar{T}(n) &\leq \frac{1}{2} \cdot \bar{T}\left(\frac{3n}{4}\right) + \frac{1}{2} \cdot \bar{T}(n) + O(n) \\ \Rightarrow \frac{1}{2} \cdot \bar{T}(n) &\leq \frac{1}{2} \cdot \bar{T}\left(\frac{3n}{4}\right) + O(n) \\ \Rightarrow \bar{T}(n) &\leq \bar{T}\left(\frac{3n}{4}\right) + O(n)\end{aligned}$$

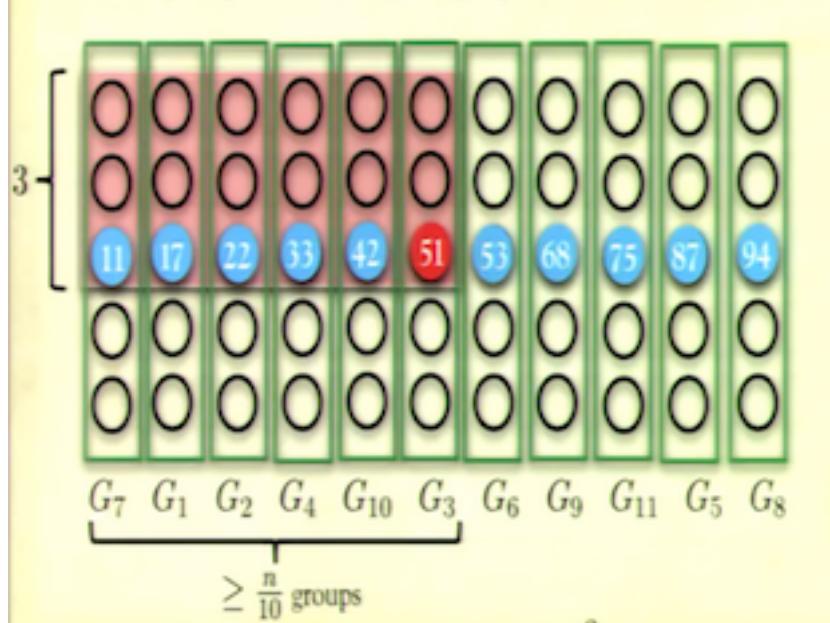
- Apply the Master Theorem
 - $a = 1$, $b = \frac{4}{3}$, and $d = 1$
 - This is Case 1 of the Master Theorem.
 - Runtime = $O(n^d) = O(n)$

3 Deterministic Selection

- So we have a **linear time randomized** algorithm for the selection problem.
- Is there a linear time **deterministic** algorithm?
- To do this, what we need is a deterministic method to find a **good pivot**.
- The idea is to find "the median of the medians."
- Divide $S = \{x_1, x_2, \dots, x_n\}$ into groups of cardinality five:
 $G_1 = \{x_1, \dots, x_5\}, G_2 = \{x_6, \dots, x_{10}\}, \dots, G_{\frac{n}{5}} = \{x_{n-4}, \dots, x_n\}$
- Now sort each group and let z_i be the median of the group G_i



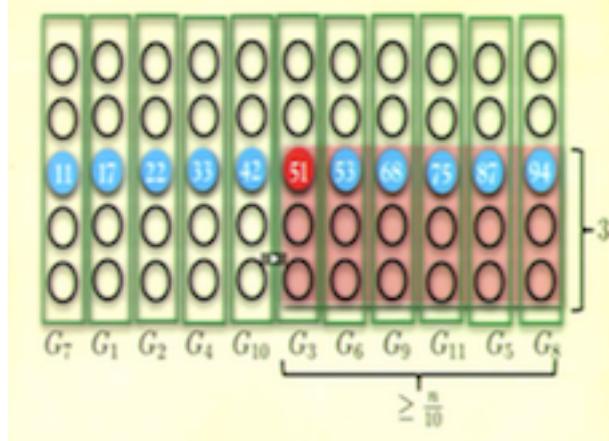
- Let m be the **median** of $Z = \{z_1, z_2, \dots, z_{\frac{n}{5}}\}$
- As a thought experiment, reorder the groups their median values:



- The **median of the medians** is greater than at least $\frac{3n}{10} - 1$ numbers in S
- $$\Rightarrow |S_R| = |\{x_i \in S \setminus m : x_i \geq m\}| \leq \frac{7}{10} \cdot n$$

- There are at least $\frac{3n}{10} - 1$ numbers in S that are at least as big as m .

$$\Rightarrow |S_L| = |\{x_i \in S : x_i < m\}| \leq \frac{7}{10} \cdot n$$



- Thus, using m as a **pivot** we have:

$$\Rightarrow \max\{|S_L|, |S_R|\} \leq \frac{7}{10} \cdot n$$

- Thus the median of the medians is a **good** pivot.
- But how actually do we find the median of the medians?
 - Using the same **deterministic recursive algorithm!**

;))

Deterministic Selection**DetSelect(S, k)****If** $|S| = 1$ **then** output x_1 **Else**Partition S into $\lceil \frac{n}{5} \rceil$ groups of 5.**For** $j = \{1, 2, \dots, \lceil \frac{n}{5} \rceil\}$ Let z_j be the median of group G_j **Let** $Z = \{z_1, z_2, \dots, z_{\lceil \frac{n}{5} \rceil}\}$ Set $m \leftarrow \text{DetSelect}(Z, \lceil \frac{n}{10} \rceil)$ Set $S_L = \{x_i \in S : x_i < m\}$ Set $S_R = \{x_i \in S \setminus \{m\} : x_i \geq m\}$ **If** $|S_L| = k - 1$ **then** output m **If** $|S_L| > k - 1$ **then** output $\text{DetSelect}(S_L, k)$ **If** $|S_L| < k - 1$ **then** output $\text{DetSelect}(S_R, k - 1 - |S_L|)$

- Thus, using m as a **pivot** we have:

$$\Rightarrow \max\{|S_L|, |S_R|\} \leq \frac{7}{10} \cdot n$$

- The recursive formula for the running time is then:

$$\Rightarrow T(n) \leq T\left(\frac{7n}{10}\right) + T\left(\frac{n}{5}\right) + O(n)$$

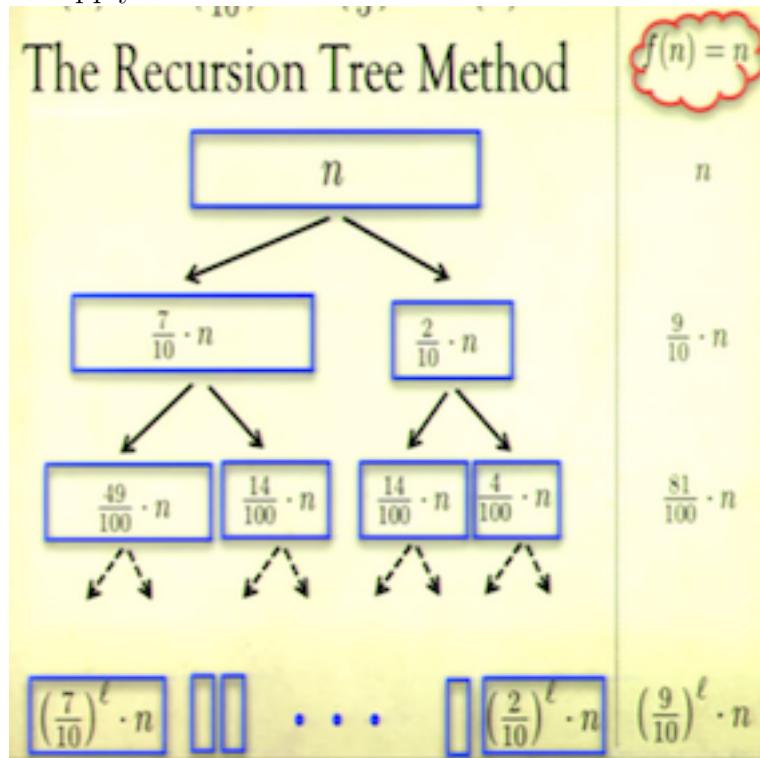
Note:

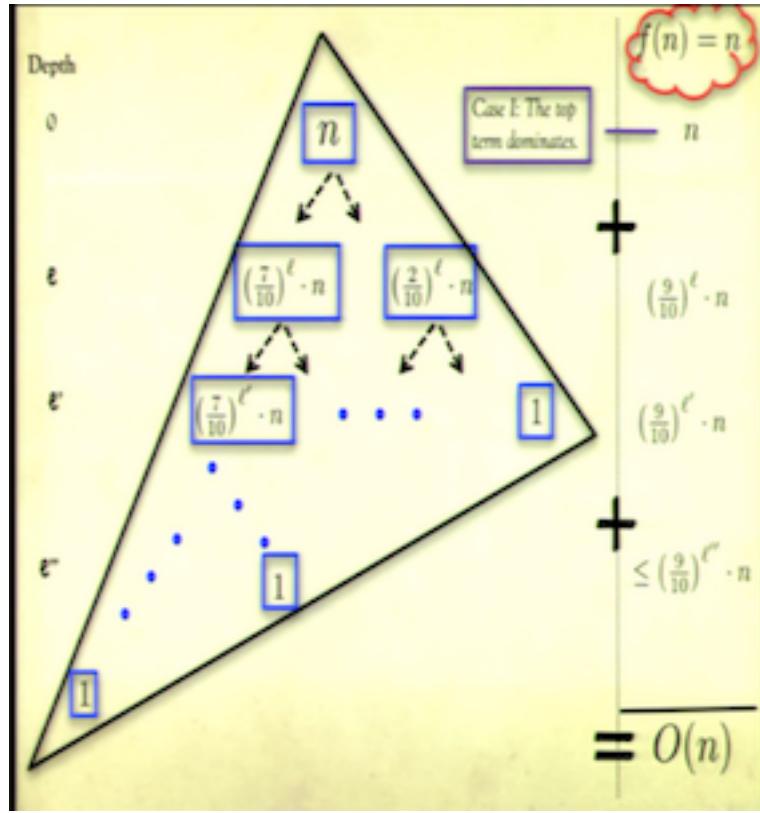
1. $T\left(\frac{n}{5}\right)$ term comes from finding the median of the medians.
2. $T\left(\frac{7n}{10}\right)$ comes from that pivoting on the median of the medians gives a significantly smaller sub-problem.
3. $O(n)$ comes from breaking in groups of size 5, finding the median of each group, and pivoting on the median of medians.

- But this does **not** fit with the Master Theorem!

- The problem is not broken into the same sized sub-problems. One is $\frac{7n}{10}$, the other is $\frac{n}{5}$.

- This does not matter as we understand the proof of the Master Theorem,
 ⇒ Apply the Recursion Tree Method!





- Runtime: $T(n) = O(n)$
- Thus, we have a deterministic **linear time** algorithm to solve the selection problem (and, specifically, to find the median).
 - The reason why finding the selection problem is useful is that we try to find the median, but when we break it into two sub-problems, we are not finding the medians in the sub-problems since things would be shifted, and we might be finding the n over 10th problem.
 - This works a lot in induction. When you do induction proof, here we are using a stronger algorithm as our sub-routines, using k th selection problem, rather than median problem, which assumes that you have a stronger induction hypothesis. Often we use induction, it is hard to prove an easy result, since we can prove a general result, using induction.

Lecture 6: Recursive Algorithms: Finding the Closest Pair of Points

1 Finding the Closest Pair of Points in the Plane

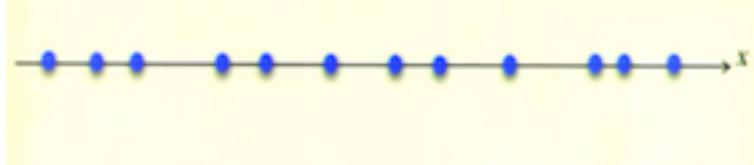
- Given n points $P = \{p_1, p_2, \dots, p_n\}$. find the **closest** pair of points.
- We will investigate the case where the points are in the plane.
 - The point set is 2-dimensional.
 - Specifically, for each $1 \leq i \leq n$, let $P_i = (x_i, y_i)$.
- How quickly can we do?

2 Exhaustive Search

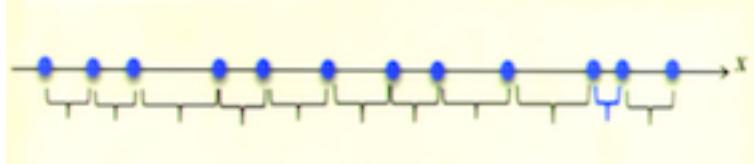
- The simplest algorithm to try is **exhaustive search**:
 - Calculate the distance between every pair of points.
 - Output the pair with the *shortest pairwise distance*.
- What is the running time?
 - There are n points so there are $\binom{n}{2} = \frac{n \cdot (n-1)}{2}$ pairs of points.
 - Runtime = $O(n^2)$
- So we have an efficient (dumb) algorithm! Is there a faster algorithm?

3 The 1-Dimensional Case

- It will be informative to first study the problem in 1-D.



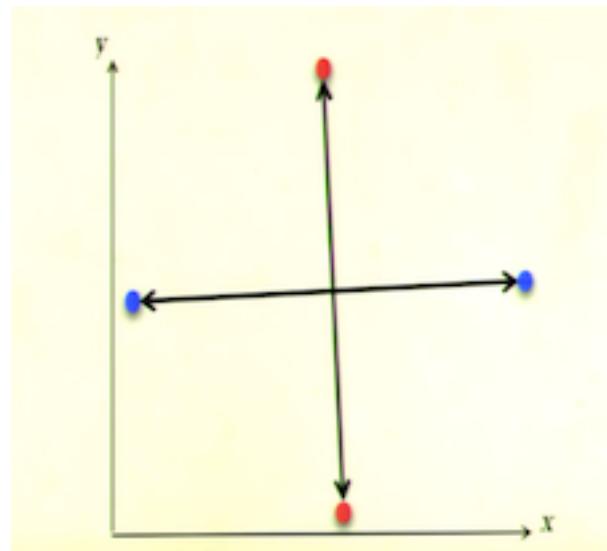
- In 1-D the closest pair of points must be adjacent in the x-ordering.



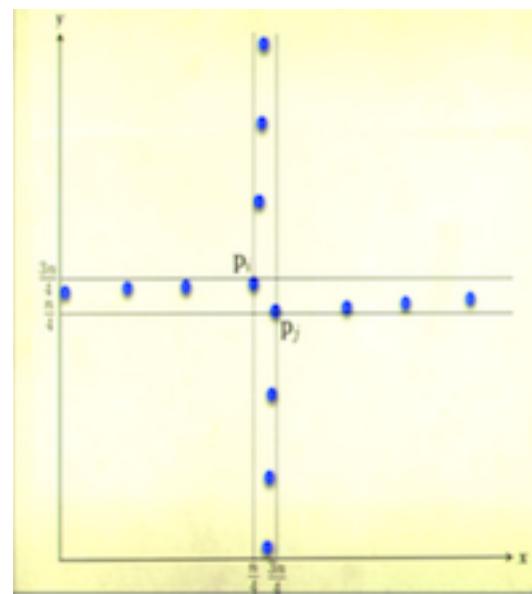
- Thus we need to calculate only $n-1$ pairwise distances.
- Runtime (If the points are already sorted in x-order) = $O(n^2)$
- Runtime (If we have to sort the points first) = $O(n \cdot \log n)$

4 The 2-Dimensional Case

- Let mimic this idea in 2-D on the point set $P = \{p_1, p_2, \dots, p_n\}$
 - Order the points by their x-coordinate; we call this x-ordering.
 - Order the points by their y-coordinate; we call this y-ordering.
 - Find the *pair of points closest* in their x-coordinate.
 - Find the *pair of points closest* in their y-coordinate.
- Does this algorithm work?

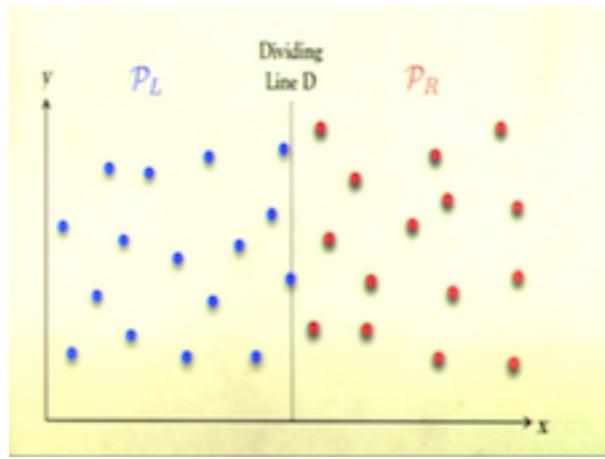


- NO!
- This algorithm does not work: points that are close in their x-coordinate (or y-coordinate) may be very far apart.
- In fact, the closest point P_i and P_j can be separated by many places in both the x-ordering and the y-ordering.

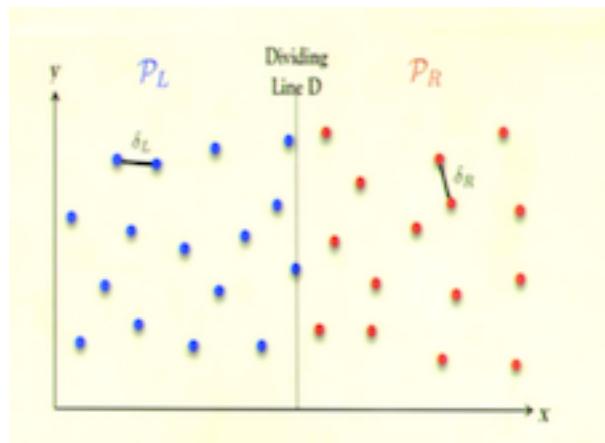


5 A Second Attempt

- Instead let's try a **divide and conquer** approach.
- Partition the points into *two groups* of cardinality $\frac{n}{2}$.
 - One way to do this is via the x-ordering using a **dividing line D**.
 - We select D to pass through the point with the median x-coordinate.
→ We have a linear time algorithm to find the median.



- Three Possibilities:
 - We can now recursively search for the closest pairs in P_L and P_R .



- But there is a third possibility:
 - The closest pair could have one point in P_L and the other in P_R .
 - So after calculating δ_L and δ_R , we must check there is no better solution between a point in P_L and a point in P_R .

6 The Recursive Algorithm

- We can now recursively search for the closest pair in P_L and P_R .
 1. Find the point q with the **median** x-coordinate.
 2. Partition P into P_L and P_R using point q .
 3. Recursively find the *closest pairs* of points in P_L and P_R .
 4. Find the *closest pair* with one point in P_L and one point in P_R .
 5. Amongst the three pairs found, output the **closest** pair.
- How long does this algorithm take?
 - The recursive formula for the running time is:
$$T(n) = 2 \cdot T\left(\frac{n}{2}\right) + O(n^2)$$

→ Thus we have $a = 2$, $b = 2$, and $d = 2$.

→ This is Case 1 of the Master Theorem.

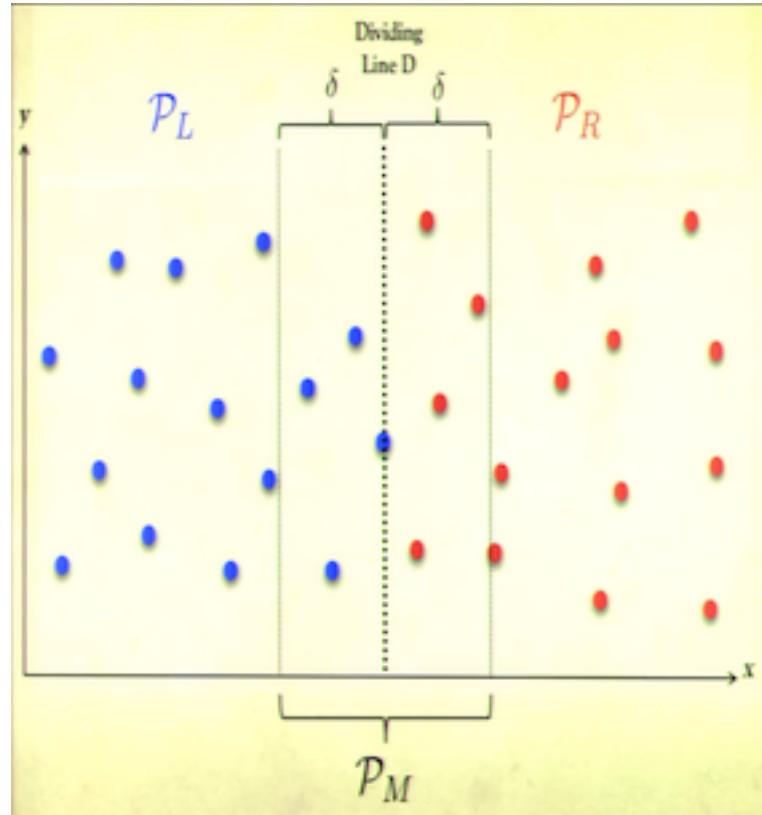
→ Runtime = $O(n^d) = O(n^2)$

 - The bottleneck operation is Step 4.
 - So it takes far too long to measure the distance between every point in P_L and every point in P_R .
 - But, by solving the subproblems of P_L and P_R , we know the **minimum pairwise distance** is at most:

$$\delta = \min\{\delta_L, \delta_R\}$$

- The trick is to use the value of δ (which we have calculated by recursion) to reduce the number of distances we measure between P_L and P_R .

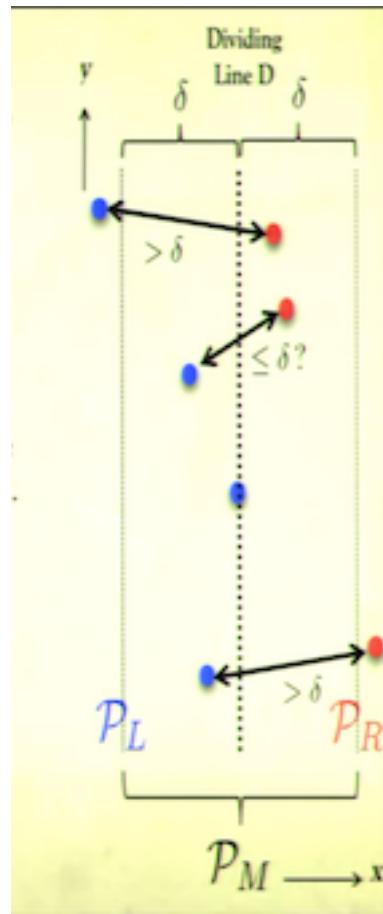
→ The key observation is that to find a better solution than δ the two points in P_L and P_R must be very close to the **dividing line** D...



- **Lemma.** Take $p_i \in P_L$ and $p_j \in P_R$. If $d(p_i, p_j) \leq \delta$ then $\{p_i, p_j\} \subseteq P_M$

Proof.

- Without loss of generality, assume $p_i \notin P_M$
- Then $d(p_i, p_j) > \delta$
- So if the closest-pair is not in one of the two sub-problems then it is within P_M .



7 A modified Recursive Algorithm

- This induces the following fine-tuned divide and conquer algorithm.
 1. Find the point q with the **median** x-coordinate.
 2. Partition P into P_L and P_R using point q .
 3. Recursively find the *closest pairs* of points in P_L and P_R .
 4. Find the *closest pair* of points in P_M . (Modified step!)
 5. Amongst the three pairs found, output the **closest** pair.
- Problem?
 - But P_M may contain **all** (or most) of the points!

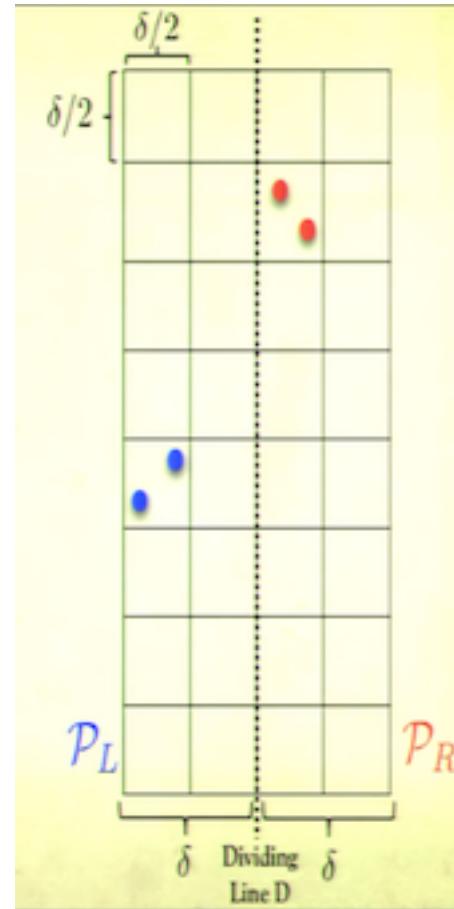
→ It may still take time $\Omega(n^2)$ to measure the distances between the points in $P_L \cap P_M$ and the points in $P_R \cap P_M$

- But the points in P_M cover a narrow band of the x-axis. As we have a **target** distance of δ , this helps a lot...
 - Consider the area of the plane within distance of δ of the line D.
 - Divide this area up into small squares of width $\frac{\delta}{2}$ as shown.

Claim. No two points of P lie in the same square.

- To prove this claim we will use the following observation.

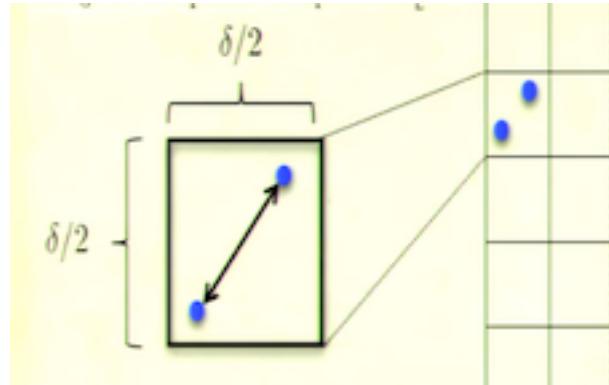
Observation. Each square lies entirely on one side of D.



Claim. No two points of P lie in the same square.

Proof.

→ *wlog* Take two points in a square in P_L .



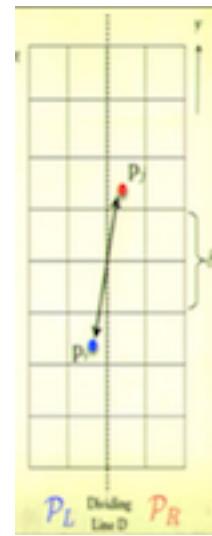
→ Their pairwise distance is: $\leq \sqrt{(\frac{\delta}{2})^2 + (\frac{\delta}{2})^2} = \frac{\delta}{\sqrt{2}} < \delta$

→ This contradicts the fact that the smallest pairwise distance in P_L is at least δ .

- **Theorem.** Take $p_i \in P_L$ and $p_j \in P_R$. If $d(p_i, p_j) \leq \delta$ then p_i and p_j have at most 10 points between them in the y-order of P_M .

Proof.

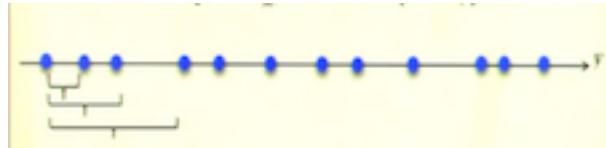
- *wlog* p_i is below p_j in the y-order.
- Then p_j is in the same row of squares as p_i or in the next two higher rows.
- If not, $d(p_i, p_j) > \delta$, a contradiction.



- But, by the claim, there is at most one point in each square.
→ There are at most 10 points between p_i and p_j in the y-order of P_M .

8 Back to the 1-D Case

- This means the basic idea for the 1-D algorithm will work here!
- To find the **closest pair** in P_M , we first sort the points by y-coordinate.



- Then, rather than finding the distances between points that are 1-apart in the y-order, we find the distances for all pairs up to 11 places apart.
- Thus we calculate less than $11n$ pairwise distances. After doing so:
 - Either we find a pair of points at distance less than δ .
 - Or we conclude that **no** such pair of points exists.
- So give $\delta = \min\{\delta_L, \delta_R\}$, we can check if there is a pair of points between P_L and P_R that are closer than δ in time $O(n)$.

9 An Enhanced Modified Recursive Algorithm

- Finally, this will give us very fast divide and conquer algorithm.
 1. Find the point q with the **median** x-coordinate.
 2. Partition P into P_L and P_R using point q .
 3. Recursively find the *closest pairs* of points in P_L and P_R .
 4. Find the *closest pair* of points in P_M using the **enhanced** 1-D algorithm.
(Modified step!)
 5. Amongst the three pairs found, output the **closest** pair.
- The recursive formula for the running time is:

$$T(n) = 2 \cdot T\left(\frac{n}{2}\right) + O(n)$$

- Find median wrt x-coordinates.
- Partition into P_L and P_R .
- Find P_M .
- Apply 1-D algorithm on P_M .
- Thus we have $a = 2$, $b = 2$, and $d = 1$.
- This is Case 2 of the Master Theorem.
- Runtime = $O(n^d \cdot \log n) = O(n \cdot \log n)$
- Validity of the algorithm:
 - As usual, the correctness of the algorithm follows by *strong induction*.
 - For the **bases cases**, we can find the closest pair in constant time by exhaustive search when there are a constant number of points left.
 - That the **inductive step** is correct follows by our discussion in the lecture.
→ **Theorem.** The recursive algorithm finds the minimum distance pair of points.

10 Computational Geometry

- The closest pair of points problem was a foundation question in the field of **computational geometry**.
- Computational has numerous applications:
 - Computer Graphics.
 - Computer Vision.
 - Geometric Information Systems.
 - ...