13. Randomized Algorithms

- contention resolution
- global min cut
- linearity of expectation
- max 3-satisfiability
- universal hashing
- Chernoff bounds
- load balancing
Randomization

Algorithmic design patterns.
  • Greedy.
  • Divide-and-conquer.
  • Dynamic programming.
  • Network flow.
  • Randomization.

Randomization. Allow fair coin flip in unit time.

Why randomize? Can lead to simplest, fastest, or only known algorithm for a particular problem.

Ex. Symmetry-breaking protocols, graph algorithms, quicksort, hashing, load balancing, closest pair, Monte Carlo integration, cryptography, ....
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Contestion resolution in a distributed system

Contestion resolution. Given $n$ processes $P_1, \ldots, P_n$, each competing for access to a shared database. If two or more processes access the database simultaneously, all processes are locked out. Devise protocol to ensure all processes get through on a regular basis.

Restriction. Processes can’t communicate.

Challenge. Need symmetry-breaking paradigm.
Contestation resolution: randomized protocol

Protocol. Each process requests access to the database at time $t$ with probability $p = 1/n$.

Claim. Let $S[i, t] = \text{event that process } i \text{ succeeds in accessing the database at time } t$. Then $1 / (e \cdot n) \leq \Pr[S(i, t)] \leq 1/(2n)$.

Pf. By independence, $\Pr[S(i, t)] = p (1 - p)^{n-1}$.

- Setting $p = 1/n$, we have $\Pr[S(i, t)] = 1/n (1 - 1/n)^{n-1}$. □

Useful facts from calculus. As $n$ increases from 2, the function:
- $(1 - 1/n)^{n-1}$ converges monotonically from $1/4$ up to $1/e$.
- $(1 - 1/n)^{n-1}$ converges monotonically from $1/2$ down to $1/e$. 
Contention resolution: randomized protocol

Claim. The probability that process $i$ fails to access the database in $en$ rounds is at most $1/e$. After $e \cdot n (c \ln n)$ rounds, the probability $\leq n^{-c}$.

Pf. Let $F[i, t]$ = event that process $i$ fails to access database in rounds 1 through $t$. By independence and previous claim, we have

$$\Pr[F[i, t]] \leq (1 - 1/(en))^t.$$

- Choose $t = [e \cdot n]$:
  $$\Pr[F(i, t)] \leq \left(1 - \frac{1}{en}\right)^{en} \leq \left(1 - \frac{1}{en}\right)^{en} \leq \frac{1}{e}$$

- Choose $t = [e \cdot n]\lfloor c \ln n \rfloor$:
  $$\Pr[F(i, t)] \leq \left(\frac{1}{e}\right)^{c \ln n} = n^{-c}$$
Contestion resolution: randomized protocol

Claim. The probability that all processes succeed within $2e \cdot n \ln n$ rounds is $\geq 1 - 1/n$.

Pf. Let $F[t]$ = event that at least one of the $n$ processes fails to access database in any of the rounds 1 through $t$.

$$\Pr[ F[t] ] = \Pr \left[ \bigcup_{i=1}^{n} F[i,t] \right] \leq \sum_{i=1}^{n} \Pr[ F[i,t] ] \leq n \left( 1 - \frac{1}{en} \right)^t$$

- Choosing $t = 2 \lceil en \rceil \lceil c \ln n \rceil$ yields $\Pr[F[t]] \leq n \cdot n^2 = 1/n$. 

Union bound. Given events $E_1, \ldots, E_n$, 

$$\Pr \left[ \bigcup_{i=1}^{n} E_i \right] \leq \sum_{i=1}^{n} \Pr[ E_i ]$$
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Global minimum cut

**Global min cut.** Given a connected, undirected graph $G = (V, E)$, find a cut $(A, B)$ of minimum cardinality.

**Applications.** Partitioning items in a database, identify clusters of related documents, network reliability, network design, circuit design, TSP solvers.

**Network flow solution.**
- Replace every edge $(u, v)$ with two antiparallel edges $(u, v)$ and $(v, u)$.
- Pick some vertex $s$ and compute $s$–$v$ cut separating $s$ from each other node $v \in V$.

**False intuition.** Global min-cut is harder than min $s$-$t$ cut.
**Contraction algorithm.** [Karger 1995]

- Pick an edge $e = (u, v)$ uniformly at random.
- **Contract** edge $e$.
  - replace $u$ and $v$ by single new super-node $w$
  - preserve edges, updating endpoints of $u$ and $v$ to $w$
  - keep parallel edges, but delete self-loops
- Repeat until graph has just two nodes $u_1$ and $v_1$.
- Return the cut (all nodes that were contracted to form $v_1$).
**Contraction algorithm.** [Karger 1995]

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Reference: Thore Husfeldt
**Contraction algorithm**

**Claim.** The contraction algorithm returns a min cut with prob $\geq 2 / n^2$.

**Pf.** Consider a global min-cut $(A^*, B^*)$ of $G$.

- Let $F^*$ be edges with one endpoint in $A^*$ and the other in $B^*$.
- Let $k = |F^*| = $ size of min cut.
- In first step, algorithm contracts an edge in $F^*$ probability $k / |E|$.
- Every node has degree $\geq k$ since otherwise $(A^*, B^*)$ would not be a min-cut $\Rightarrow |E| \geq \frac{1}{2} k n \iff k / |E| \leq 2 / n$.
- Thus, algorithm contracts an edge in $F^*$ with probability $\leq 2 / n$. 

![Diagram of contraction algorithm](image)
Contraction algorithm

Claim. The contraction algorithm returns a min cut with prob $\geq \frac{2}{n^2}$.

Pf. Consider a global min-cut $(A^*, B^*)$ of $G$.

- Let $F^*$ be edges with one endpoint in $A^*$ and the other in $B^*$.
- Let $k = |F^*| = \text{size of min cut}.$
- Let $G'$ be graph after $j$ iterations. There are $n' = n - j$ supernodes.
- Suppose no edge in $F^*$ has been contracted. The min-cut in $G'$ is still $k$.
- Since value of min-cut is $k$, $|E'| \geq \frac{1}{2} k n' \iff k / |E'| \leq \frac{2}{n'}$.
- Thus, algorithm contracts an edge in $F^*$ with probability $\leq \frac{2}{n'}$.
- Let $E_j = \text{event that an edge in } F^* \text{ is not contracted in iteration } j.$

\[
\Pr[E_1 \cap E_2 \cdots \cap E_{n-2}] = \Pr[E_1] \times \Pr[E_2 | E_1] \times \cdots \times \Pr[E_{n-2} | E_1 \cap E_2 \cdots \cap E_{n-3}]
\geq \left(1 - \frac{2}{n}\right) \left(1 - \frac{2}{n-1}\right) \cdots \left(1 - \frac{2}{n-4}\right) \left(1 - \frac{2}{3}\right)
= \left(\frac{n-2}{n}\right) \left(\frac{n-3}{n-1}\right) \cdots \left(\frac{2}{4}\right) \left(\frac{1}{3}\right)
= \frac{2}{n(n-1)}
\geq \frac{2}{n^2}
\]
**Contraction algorithm**

**Amplification.** To amplify the probability of success, run the contraction algorithm many times.

**Claim.** If we repeat the contraction algorithm \( n^2 \ln n \) times, then the probability of failing to find the global min-cut is \( \leq 1 / n^2 \).

**Pf.** By independence, the probability of failure is at most

\[
\left(1 - \frac{2}{n^2}\right)^{n^2 \ln n} = \left[\left(1 - \frac{2}{n^2}\right)^{\frac{1}{2} n^2}\right]^{2 \ln n} \leq \left(e^{-1}\right)^{2 \ln n} = \frac{1}{n^2}
\]

\((1 - 1/x)^x \leq 1/e\)
Contraction algorithm: example execution

Reference: Thore Husfeldt
Global min cut: context

Remark. Overall running time is slow since we perform $\Theta(n^2 \log n)$ iterations and each takes $\Omega(m)$ time.

Improvement. [Karger–Stein 1996] $O(n^2 \log^3 n)$.

- Early iterations are less risky than later ones: probability of contracting an edge in min cut hits 50% when $n / \sqrt{2}$ nodes remain.
- Run contraction algorithm until $n / \sqrt{2}$ nodes remain.
- Run contraction algorithm twice on resulting graph and return best of two cuts.

Extensions. Naturally generalizes to handle positive weights.

Best known. [Karger 2000] $O(m \log^3 n)$.

faster than best known max flow algorithm or deterministic global min cut algorithm
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**Expectation**

**Expectation.** Given a discrete random variable $X$, its expectation $E[X]$ is defined by:

$$E[X] = \sum_{j=0}^{\infty} j \Pr[X = j]$$

**Waiting for a first success.** Coin is heads with probability $p$ and tails with probability $1-p$. How many independent flips $X$ until first heads?

$$E[X] = \sum_{j=0}^{\infty} j \cdot \Pr[X = j] = \sum_{j=0}^{\infty} j (1 - p)^{j-1} p = \frac{p}{1-p} \sum_{j=0}^{\infty} j (1 - p)^j = \frac{p}{1-p} \cdot \frac{1-p}{p^2} = \frac{1}{p}$$

$$j - 1 \text{ tails} \quad 1 \text{ head}$$
Expectation: two properties

**Useful property.** If $X$ is a 0/1 random variable, $E[X] = \Pr[X = 1]$.

**Pf.** \[
E[X] = \sum_{j=0}^{\infty} j \cdot \Pr[X = j] = \sum_{j=0}^{1} j \cdot \Pr[X = j] = \Pr[X = 1]
\]

(not necessarily independent)

**Linearity of expectation.** Given two random variables $X$ and $Y$ defined over the same probability space, $E[X + Y] = E[X] + E[Y]$.

**Benefit.** Decouples a complex calculation into simpler pieces.
Guessing cards

Game. Shuffle a deck of $n$ cards; turn them over one at a time; try to guess each card.

Memoryless guessing. No psychic abilities; can’t even remember what’s been turned over already. Guess a card from full deck uniformly at random.

Claim. The expected number of correct guesses is 1.

Pf. [surprisingly effortless using linearity of expectation]

- Let $X_i = 1$ if $i^{th}$ prediction is correct and 0 otherwise.
- Let $X = \text{number of correct guesses} = X_1 + \ldots + X_n$.
- $E[X_i] = \Pr[X_i = 1] = 1/n$.
- $E[X] = E[X_1] + \ldots + E[X_n] = 1/n + \ldots + 1/n = 1$. □

*linearity of expectation*
Guessing cards

**Game.** Shuffle a deck of $n$ cards; turn them over one at a time; try to guess each card.

**Guessing with memory.** Guess a card uniformly at random from cards not yet seen.

**Claim.** The expected number of correct guesses is $\Theta(\log n)$.

**Pf.**

- Let $X_i = 1$ if $i^{th}$ prediction is correct and 0 otherwise.
- Let $X = \text{number of correct guesses} = X_1 + \ldots + X_n$.
- $E[X_i] = \Pr[X_i = 1] = 1 / (n - (i - 1))$.
- $E[X] = E[X_1] + \ldots + E[X_n] = 1 / n + \ldots + 1 / 2 + 1 / 1 = H(n)$.

\[\ln(n+1) < H(n) < 1 + \ln n\]
Coupon collector

Coupon collector. Each box of cereal contains a coupon. There are $n$ different types of coupons. Assuming all boxes are equally likely to contain each coupon, how many boxes before you have $\geq 1$ coupon of each type?

Claim. The expected number of steps is $\Theta(n \log n)$.

Pf.

- Phase $j = \text{time between } j \text{ and } j + 1 \text{ distinct coupons.}$
- Let $X_j = \text{number of steps you spend in phase } j$.
- Let $X = \text{number of steps in total} = X_0 + X_1 + \ldots + X_{n-1}$.

$$E[X] = \sum_{j=0}^{n-1} E[X_j] = \sum_{j=0}^{n-1} \frac{n}{n-j} = n \sum_{i=1}^{n} \frac{1}{i} = nH(n)$$

\[\begin{align*}
\text{prob of success} &= \frac{n-j}{n} \\
\Rightarrow \text{expected waiting time} &= \frac{n}{n-j}
\end{align*}\]
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Maximum 3-satisfiability

Maximum 3-satisfiability. Given a 3-SAT formula, find a truth assignment that satisfies as many clauses as possible.

\[
\begin{align*}
C_1 &= x_2 \lor \overline{x_3} \lor \overline{x_4} \\
C_2 &= x_2 \lor x_3 \lor x_4 \\
C_3 &= \overline{x_1} \lor x_2 \lor x_4 \\
C_4 &= \overline{x_1} \lor \overline{x_2} \lor x_3 \\
C_5 &= x_1 \lor x_2 \lor \overline{x_4}
\end{align*}
\]

Remark. \textbf{NP}-hard search problem.

Simple idea. Flip a coin, and set each variable true with probability \(\frac{1}{2}\), independently for each variable.
Maximum 3-satisfiability: analysis

Claim. Given a 3-SAT formula with \( k \) clauses, the expected number of clauses satisfied by a random assignment is \( 7k / 8 \).

Pf. Consider random variable
\[
Z_j = \begin{cases} 
1 & \text{if clause } C_j \text{ is satisfied} \\
0 & \text{otherwise.}
\end{cases}
\]

- Let \( Z = \) number of clauses satisfied by random assignment.

\[
E[Z] = \sum_{j=1}^{k} E[Z_j]
\]

(linearity of expectation)

\[
= \sum_{j=1}^{k} \Pr[\text{clause } C_j \text{ is satisfied}]
= \frac{7}{8} k
\]
The probabilistic method

Corollary. For any instance of 3-SAT, there exists a truth assignment that satisfies at least a 7/8 fraction of all clauses.

Pf. Random variable is at least its expectation some of the time. ■

Probabilistic method. [Paul Erdös] Prove the existence of a non-obvious property by showing that a random construction produces it with positive probability!
**Maximum 3-satisfiability: analysis**

**Q.** Can we turn this idea into a 7/8-approximation algorithm?

**A.** Yes (but a random variable can almost always be below its mean).

**Lemma.** The probability that a random assignment satisfies \( \geq 7k / 8 \) clauses is at least \( 1 / (8k) \).

**Pf.** Let \( p_j \) be probability that exactly \( j \) clauses are satisfied; let \( p \) be probability that \( \geq 7k / 8 \) clauses are satisfied.

\[
\frac{7}{8} k = E[Z] = \sum_{j \geq 0} j p_j
\]

\[
= \sum_{j < 7k/8} j p_j + \sum_{j \geq 7k/8} j p_j
\]

\[
\leq \left( \frac{7}{8} k - \frac{1}{8} \right) \sum_{j < 7k/8} p_j + k \sum_{j \geq 7k/8} p_j
\]

\[
\leq \left( \frac{7}{8} k - \frac{1}{8} \right) \cdot 1 + k p
\]

Rearranging terms yields \( p \geq 1 / (8k) \).
Maximum 3-satisfiability: analysis

Johnson’s algorithm. Repeatedly generate random truth assignments until one of them satisfies \( \geq \frac{7k}{8} \) clauses.

Theorem. Johnson’s algorithm is a 7/8-approximation algorithm.

Pf. By previous lemma, each iteration succeeds with probability \( \geq \frac{1}{(8k)} \). By the waiting-time bound, the expected number of trials to find the satisfying assignment is at most \( 8k \). □
Maximum satisfiability

Extensions.
- Allow one, two, or more literals per clause.
- Find max weighted set of satisfied clauses.

Theorem. [Asano–Williamson 2000] There exists a 0.784-approximation algorithm for \( \text{MAX-SAT} \).

Theorem. [Karloff–Zwick 1997, Zwick+computer 2002] There exists a 7/8-approximation algorithm for version of \( \text{MAX-3-SAT} \) in which each clause has at most 3 literals.

Theorem. [Håstad 1997] Unless \( P = \text{NP} \), no \( \rho \)-approximation algorithm for \( \text{MAX-3-SAT} \) (and hence \( \text{MAX-SAT} \)) for any \( \rho > 7/8 \).

very unlikely to improve over simple randomized algorithm for \( \text{MAX-3-SAT} \)
Monte Carlo vs. Las Vegas algorithms

**Monte Carlo.** Guaranteed to run in poly-time, likely to find correct answer. 
*Ex:* Contraction algorithm for global min cut.

**Las Vegas.** Guaranteed to find correct answer, likely to run in poly-time. 
*Ex:* Randomized quicksort, Johnson’s MAX-3-SAT algorithm.

**Remark.** Can always convert a Las Vegas algorithm into Monte Carlo, but no known method (in general) to convert the other way.
RP and ZPP

**RP.** [Monte Carlo] Decision problems solvable with one-sided error in poly-time.

One-sided error.
- If the correct answer is *no*, always return *no*.
- If the correct answer is *yes*, return *yes* with probability $\geq \frac{1}{2}$.

**ZPP.** [Las Vegas] Decision problems solvable in expected poly-time.

Theorem. $P \subseteq ZPP \subseteq RP \subseteq NP$.

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Dictionary data type

**Dictionary.** Given a universe $U$ of possible elements, maintain a subset $S \subseteq U$ so that inserting, deleting, and searching in $S$ is efficient.

**Dictionary interface.**

- $create()$: initialize a dictionary with $S = \emptyset$.
- $insert(u)$: add element $u \in U$ to $S$.
- $delete(u)$: delete $u$ from $S$ (if $u$ is currently in $S$).
- $lookup(u)$: is $u$ in $S$?

**Challenge.** Universe $U$ can be extremely large so defining an array of size $|U|$ is infeasible.

**Applications.** File systems, databases, Google, compilers, checksums P2P networks, associative arrays, cryptography, web caching, etc.
Hashing

Hash function. \( h : U \rightarrow \{ 0, 1, \ldots, n - 1 \} \).

Hashing. Create an array \( a \) of length \( n \). When processing element \( u \), access array element \( a[h(u)] \).

Collision. When \( h(u) = h(v) \) but \( u \neq v \).
- A collision is expected after \( \Theta(\sqrt{n}) \) random insertions.
- Separate chaining: \( a[i] \) stores linked list of elements \( u \) with \( h(u) = i \).
Ad-hoc hash function

Ad-hoc hash function.

```java
int hash(String s, int n) {
    int hash = 0;
    for (int i = 0; i < s.length(); i++)
        hash = (31 * hash) + s[i];
    return hash % n;
}
```

hash function à la Java string library

**Deterministic hashing.** If $|U| \geq n^2$, then for any fixed hash function $h$, there is a subset $S \subseteq U$ of $n$ elements that all hash to same slot. Thus, $\Theta(n)$ time per lookup in worst-case.

**Q.** But isn’t ad-hoc hash function good enough in practice?
Algorithmic complexity attacks

When can’t we live with ad-hoc hash function?

- Obvious situations: aircraft control, nuclear reactor, pace maker, ....
- Surprising situations: denial-of-service attacks.

Real world exploits. [Crosby–Wallach 2003]

- Linux 2.4.20 kernel: save files with carefully chosen names.
- Perl 5.8.0: insert carefully chosen strings into associative array.
- Bro server: send carefully chosen packets to DOS the server, using less bandwidth than a dial-up modem.
Hashing performance

**Ideal hash function.** Maps $m$ elements uniformly at random to $n$ hash slots.
- Running time depends on length of chains.
- Average length of chain = $\alpha = m / n$.
- Choose $n \approx m \Rightarrow$ expect $O(1)$ per insert, lookup, or delete.

**Challenge.** Explicit hash function $h$ that achieves $O(1)$ per operation.

**Approach.** Use randomization for the choice of $h$.

adversary knows the randomized algorithm you’re using, but doesn’t know random choice that the algorithm makes
Universal hashing (Carter–Wegman 1980s)

A universal family of hash functions is a set of hash functions $H$ mapping a universe $U$ to the set $\{0, 1, \ldots, n - 1\}$ such that

- For any pair of elements $u \neq v$: $\Pr_{h \in H} [h(u) = h(v)] \leq 1/n$
- Can select random $h$ efficiently.
- Can compute $h(u)$ efficiently.

**Ex.** $U = \{a, b, c, d, e, f\}$, $n = 2$.

$$H = \{h_1, h_2\}$$

$\Pr_{h \in H} [h(a) = h(b)] = 1/2$

$\Pr_{h \in H} [h(a) = h(c)] = 1$

$\Pr_{h \in H} [h(a) = h(d)] = 0$

$\ldots$

$H = \{h_1, h_2, h_3, h_4\}$

$\Pr_{h \in H} [h(a) = h(b)] = 1/2$

$\Pr_{h \in H} [h(a) = h(c)] = 1/2$

$\Pr_{h \in H} [h(a) = h(d)] = 1/2$

$\Pr_{h \in H} [h(a) = h(e)] = 1/2$

$\Pr_{h \in H} [h(a) = h(f)] = 0$

$\ldots$
Universal hashing: analysis

**Proposition.** Let $H$ be a universal family of hash functions mapping a universe $U$ to the set $\{0, 1, \ldots, n-1\}$; let $h \in H$ be chosen uniformly at random from $H$; let $S \subseteq U$ be a subset of size at most $n$; and let $u \not\in S$. Then, the expected number of items in $S$ that collide with $u$ is at most 1.

**Pf.** For any $s \in S$, define random variable $X_s = 1$ if $h(s) = h(u)$, and 0 otherwise. Let $X$ be a random variable counting the total number of collisions with $u$.

$$E_{h \in H}[X] = E[\sum_{s \in S} X_s] = \sum_{s \in S} E[X_s] = \sum_{s \in S} \Pr[X_s = 1] \leq \sum_{s \in S} \frac{1}{n} = \frac{|S|}{n} \leq 1$$

- linearity of expectation
- $X_s$ is a 0–1 random variable
- universal

**Q.** OK, but how do we design a universal class of hash functions?
Designing a universal family of hash functions

**Modulus.** We will use a prime number $p$ for the size of the hash table.

**Integer encoding.** Identify each element $u \in U$ with a base-$p$ integer of $r$ digits: $x = (x_1, x_2, \ldots, x_r)$.

**Hash function.** Let $A$ = set of all $r$-digit, base-$p$ integers. For each $a = (a_1, a_2, \ldots, a_r)$ where $0 \leq a_i < p$, define

$$h_a(x) = \left( \sum_{i=1}^{r} a_i x_i \right) \mod p$$

maps universe $U$ to set $\{ 0, 1, \ldots, p - 1 \}$

**Hash function family.** $H = \{ h_a : a \in A \}$. 
Designing a universal family of hash functions

**Theorem.** $H = \{ h_a : a \in A \}$ is a universal family of hash functions.

**Pf.** Let $x = (x_1, x_2, \ldots, x_r)$ and $y = (y_1, y_2, \ldots, y_r)$ be two distinct elements of $U$.

We need to show that $\Pr[h_a(x) = h_a(y)] \leq 1 / p$.

- Since $x \neq y$, there exists an integer $j$ such that $x_j \neq y_j$.
- We have $h_a(x) = h_a(y)$ iff

$$a_j (y_j - x_j) \equiv \sum_{i \neq j} a_i (x_i - y_i) \mod p$$

- Can assume $a$ was chosen uniformly at random by first selecting all coordinates $a_i$ where $i \neq j$, then selecting $a_j$ at random. Thus, we can assume $a_i$ is fixed for all coordinates $i \neq j$.
- Since $p$ is prime, $a_j z \equiv m \mod p$ has at most one solution among $p$ possibilities.  

Thus $\Pr[h_a(x) = h_a(y)] \leq 1 / p$.  

Number theory fact

**Fact.** Let \( p \) be prime, and let \( z \not\equiv 0 \mod p \). Then \( \alpha z \equiv m \mod p \) has at most one solution \( 0 \leq \alpha < p \).

**Pf.**

- Suppose \( 0 \leq \alpha_1 < p \) and \( 0 \leq \alpha_2 < p \) are two different solutions.
- Then \( (\alpha_1 - \alpha_2) z \equiv 0 \mod p \); hence \( (\alpha_1 - \alpha_2) z \) is divisible by \( p \).
- Since \( z \not\equiv 0 \mod p \), we know that \( z \) is not divisible by \( p \).
- It follows that \( (\alpha_1 - \alpha_2) \) is divisible by \( p \).
- This implies \( \alpha_1 = \alpha_2 \).  

**Bonus fact.** Can replace “at most one” with “exactly one” in above fact.

**Pf idea.** Euclid’s algorithm.

**Here’s where we use that \( p \) is prime**
Universal hashing: summary

Goal. Given a universe $U$, maintain a subset $S \subseteq U$ so that insert, delete, and lookup are efficient.

Universal hash function family. $H = \{ h_a : a \in A \}$.

\[ h_a(x) = \left( \sum_{i=1}^{r} a_i x_i \right) \mod p \]

- Choose $p$ prime so that $m \leq p \leq 2m$, where $m = |S|$.
- Fact: there exists a prime between $m$ and $2m$. can find such a prime using another randomized algorithm (!)

Consequence.
- Space used $= \Theta(m)$.
- Expected number of collisions per operation is $\leq 1$
  $\Rightarrow$ $O(1)$ time per insert, delete, or lookup.
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**Theorem.** Suppose $X_1, \ldots, X_n$ are independent 0-1 random variables. Let $X = X_1 + \ldots + X_n$. Then for any $\mu \geq E[X]$ and for any $\delta > 0$, we have

$$\Pr[X > (1 + \delta)\mu] < \left[\frac{e^\delta}{(1 + \delta)^{1+\delta}}\right]^\mu$$

**Pf.** We apply a number of simple transformations.

- For any $t > 0$,

  $$\Pr[X > (1 + \delta)\mu] = \Pr\left[e^{tx} > e^{t(1+\delta)\mu}\right] \leq e^{-t(1+\delta)\mu} \cdot E[e^{tX}]$$

  - $f(x) = e^{tx}$ is monotone in $x$  
  - Markov’s inequality: $\Pr[X > a] \leq E[X] / a$

- Now

  $$E[e^{tX}] = E[e^{t \sum_i X_i}] = \prod_i E[e^{tX_i}]$$

  - definition of $X$
  - independence
Chernoff Bounds (above mean)

**Pf.** [ continued ]

- Let \( p_i = \Pr [X_i = 1] \). Then,

\[
E[e^{tX_i}] = p_ie^t + (1 - p_i)e^0 = 1 + p_i(e^t - 1) \leq e^{p_i(e^t - 1)}
\]

for any \( \alpha \geq 0, 1 + \alpha \leq e^\alpha \)

- Combining everything:

\[
\Pr[X > (1 + \delta)\mu] \leq e^{-t(1+\delta)\mu} \prod_i E[e^{tX_i}] \leq e^{-t(1+\delta)\mu} \prod_i e^{p_i(e^t - 1)} \leq e^{-t(1+\delta)\mu} e^{\mu(e^t - 1)}
\]

previous slide inequality above \( \sum_i p_i = \mathbb{E}[X] \leq \mu \)

- Finally, choose \( t = \ln(1 + \delta) \). ■
Chernoff Bounds (below mean)

**Theorem.** Suppose $X_1, \ldots, X_n$ are independent 0-1 random variables. Let $X = X_1 + \ldots + X_n$. Then for any $\mu \leq E[X]$ and for any $0 < \delta < 1$, we have

$$\Pr[X < (1 - \delta)\mu] < e^{-\delta^2 \mu / 2}$$

**Pf idea.** Similar.

**Remark.** Not quite symmetric since only makes sense to consider $\delta < 1$. 
13. RANDOMIZED ALGORITHMS

- contention resolution
- global min cut
- linearity of expectation
- max 3-satisfiability
- universal hashing
- Chernoff bounds
- load balancing
Load balancing

Load balancing. System in which $m$ jobs arrive in a stream and need to be processed immediately on $m$ identical processors. Find an assignment that balances the workload across processors.

Centralized controller. Assign jobs in round-robin manner. Each processor receives at most $\lfloor m/n \rfloor$ jobs.

Decentralized controller. Assign jobs to processors uniformly at random. How likely is it that some processor is assigned “too many” jobs?
Load balancing

Analysis.

• Let $X_i$ = number of jobs assigned to processor $i$.
• Let $Y_{ij} = 1$ if job $j$ assigned to processor $i$, and 0 otherwise.
• We have $E[Y_{ij}] = 1/n$.
• Thus, $X_i = \sum_j Y_{ij}$, and $\mu = E[X_i] = 1$.
• Applying Chernoff bounds with $\delta = c - 1$ yields $\Pr[X_i > c] < \frac{e^{c-1}}{c^c}$

• Let $\gamma(n)$ be number $x$ such that $x^x = n$, and choose $c = e \gamma(n)$.

\[
\Pr[X_i > c] < \frac{e^{c-1}}{c^c} < \left(\frac{e}{c}\right)^c = \left(\frac{1}{\gamma(n)}\right)^{e\gamma(n)} < \left(\frac{1}{\gamma(n)}\right)^{2\gamma(n)} = \frac{1}{n^2}
\]

• Union bound $\Rightarrow$ with probability $\geq 1 - 1/n$ no processor receives more than $e \gamma(n) = \Theta(\log n / \log \log n)$ jobs.

Bonus fact: with high probability, some processor receives $\Theta(\log n / \log \log n)$ jobs
Load balancing: many jobs

**Theorem.** Suppose the number of jobs $m = 16n \ln n$. Then on average, each of the $n$ processors handles $\mu = 16 \ln n$ jobs. With high probability, every processor will have between half and twice the average load.

**Pf.**
- Let $X_i, Y_{ij}$ be as before.
- Applying Chernoff bounds with $\delta = 1$ yields

\[
\Pr[X_i > 2\mu] < \left(\frac{e}{4}\right)^{16n \ln n} < \left(\frac{1}{e}\right)^{\ln n} = \frac{1}{n^2}
\]

\[
\Pr[X_i < \frac{1}{2}\mu] < e^{-\frac{1}{2} \left(\frac{1}{2}\right)^2 16n \ln n} = \frac{1}{n^2}
\]

- Union bound $\Rightarrow$ every processor has load between half and twice the average with probability $\geq 1 - 2/n$.  