12. Approximation

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Big Idea: Renegotiating Problems

Sometimes we want to solve a problem, but there is an obstacle

- computational complexity: problem is *NP*-hard or undecidable
- ill-posed: don't know how to phrase problem as precise input/output statement

These are insurmountable; progress not possible.

Sometimes we can *negotiate* on the definition of the problem

- adjust input/output def'n to correspond to an easier problem
- more specific input; or more general output
- ideally, computational problem still helps with the business problem
- combines CS hard skills with business soft skills

Approximation

Approximation: output is *nearly-optimal* but not necessarily truly optimal.

- proximity to optimality is quantified, **proven**
- "approximation", "approximate" are technical terms; use other words like "decent" for informal ideas about near-optimality
- suitable for business scenarios where approximate solutions are adequate
- need to rewrite problem definition
- every optimization problem has corresponding approximation problems; but these are distinct problems

Example: optimal vs. approximate graph coloring

graph coloring
input: connected graph $G = (V, E)$
output: coloring c using k colors, where each vertex $v \in V$ is assigned color $c(v) \in \{1, \ldots k\}$, no pair of adjacent vertices are assigned the same color, and the number of colors k is minimal

3-approximate graph coloring
input: connected graph $G = (V, E)$
output: coloring c using k colors, where each vertex $v \in V$ is assigned color $c(v) \in \{1, \ldots k\}$, no pair of adjacent vertices are assigned the same color, **and the number of colors k satisfied $k \leq 3k^\star$, where $k^\star$ is the fewest colors possible for G**

Approximation vs. Other Approaches

Other ways of dealing with unsolvable problems:

- say "no"
- when n is small enough, just use exponential-time algorithm
- no *proof* of solution quality, but nonetheless sometimes good enough:
  - machine learning algorithms (also, in M.L. humans don't need to precisely define what counts as "correct")
  - fast heuristic algorithms
  - Monte Carlo algorithms
  - other AI algorithms

Approximation

- pros: *provable* solution quality, often fast
- con: human needs to design and analyze algorithm for each specific problem

Performance Ratios

Approximation ratio $\rho(n)$: ratio between quality of algorithm's output and optimal output; smaller is better

- for **maximization** problem: if optimal quality is $C^\star$ and alg. produces quality C, by definition $C^\star \geq C$, and define

$$\rho(n) = \frac{C^\star}{C}$$

- for **minimization** problem: if optimal quality is $C^\star$ and alg. produces quality C, by definition $C \geq C^\star$ and define

$$\rho(n) = \frac{C}{C^\star}$$

Recall 3-approx. vertex cover: output $\#$ colors $\leq 3k^\star$

Fixed Approximation Ratios

Some approxmation algorithms have a fixed approximation ratio that is “baked in” to the design of the algorithm.

Ex.: algorithm that solves 3-approx. vertex cover would have fixed $\rho(n) = 3$

In general, better (smaller) ratios require slower algorithms.
(note 1-approximation algorithms produce optimal solutions.)

Deriving a different $\rho(n)$ vs. time trade-offs requires designing an entirely different algorithm.

Approximation Schemes

approximation scheme: family of related algorithms, such that, for any parameter $\epsilon > 0$, scheme defines a $(1+\epsilon)$-approximate algorithm

- think of ϵ as being a **const** variable
- time-performance trade-off is fully tuneable at compile time

Polynomial Time Approximation Scheme (PTAS): approx. scheme where runtime is polynomial in n; nothing said of relationship to ϵ
e.g. $O(2^{1/\epsilon} n \log n)$

Fully PTAS: runtime is polynomial in n and $1/\epsilon$
e.g. $O((1/\epsilon)^2 n^3)$

Vertex Cover Problem

vertex cover problem
input: undirected graph $G = (V, E)$
output: set of vertices $C \subseteq V$, of minimal size $|C|$, such that every edge in E is incident on at least one vertex in C

2-approximate vertex cover problem
input: undirected graph $G = (V, E)$
output: set of vertices $C \subseteq V$, such that every edge in E is incident on at least one vertex in C, and $|C| \leq 2|C^\star|$ where $C^\star$ is a minimal vertex cover for G

See Wiki page:
https://en.wikipedia.org/wiki/Vertex_cover

A Greedy Approximation Algorithm

Idea:

- every edge $e = (u, v)$ needs both $u \in C$ and $v \in C$
- so grab an edge $e = (u, v)$ and include u and v in C
- every other edge touching u or v is now covered, so eliminate them
- continue until every edge is either grabbed or eliminated
- good: definitely finds a correct cover C
- bad: depending on the order of the "grabs", heuristic can get tricked into picking sub-optimal vertices

2-Approximate Vertex Cover Pseudocode

```
function APPROX-VERTEX-COVER(G = (V, E))
    C = ∅
    T = E
    while T ≠ ∅ do
        Let e = (u, v) be an arbitrary edge in T
        C = C ∪ {u, v}
        Remove from T any edge f that is incident on u or v
    end while
    return C
end function
```

Efficiency Analysis: $O(m + n)$, using proper data structures

Vertex Cover Performance Ratio

Lemma: APPROX-VERTEX-COVER is a 2-approximation algorithm.
Need: for any $G, |C| \le 2|C^\star|$
Proof sketch:

- Let A be the set of edges chosen inside the **while** loop
- will bound $|C|, |C^\star|$ both in terms of $|A|$
- **(1)** $|C^\star|$ vs. $|A|$
- $C^\star$ is a vertex cover, so for every edge $(u, v) \in A$, we must have $u \in C^\star$ and/or $v \in C^\star$
- the "Remove from T" step guarantees that, after (u, v) is chosen, no other edge incident on u or v will be chosen and added to A
- $\Rightarrow$ each $x \in C^\star$ covers *exactly* one edge in A
- $\Rightarrow |C^\star| \ge |A|$

Vertex Cover Performance Ratio (cont'd)

- **(2)** $|C|$ vs. $|A|$
- the $C = C \cup \{u, v\}$ step inserts 2 vertices into C
- due to the same "Remove from T" logic, neither u nor v was already in C
- $\Rightarrow |C| = 2|A|$ (note exact equality)
- **combining (1) and (2)**

$$|C| = 2|A| \leq 2|C^\star|$$

- QED

Commentary on this Proof

- note that us analysts do not know concretely which vertices are in $C^\star$
- the algorithm certainly doesn't know what $C^\star$ is, either
- all we do know is that, due to the definition of vertex cover, and the logic of our algorithm,

  $$\#\text{ vertices in optimal cover} \geq \#\text{ iterations } \textbf{while} \text{ loop}$$

- and, due to algorithm logic,

  $$\#\text{ iterations } \textbf{while} \text{ loop} = \#\text{ vertices chosen for approx. cover}$$

- in general, to prove an approx. ratio, need
  1. to bound quality of arbitrary, opaque optimal solution; and
  2. bound quality of approx. solution the same way

TSP

traveling salesperon problem (TSP)
input: a complete undirected graph $G = (V, E)$ where each edge has weight $w(e) \geq 0$
output: a sequence of vertices H forming a Hamiltonian cycle, minimizing total edge weight

Recall:

- *Cycle:* path that starts and ends at same vertex
- *Hamiltonian:* visits each vertex exactly once
- every complete graph contains some Hamiltonian cycle

Bad news:

- TSP is NP-complete; if $P \neq NP$, no polynomial-time optimization algorithm
- TSP is also APX-complete; if $P \neq NP$, no PTAS

Triangle Inequality

Triangle inequality in general: for distance function d and sites a, b, c,

$$d(a, c) \leq d(a, b) + d(b, c)$$

$\Rightarrow$ direct path $a \to b$ always cheaper than two-step path $a \to b \to c$ (or tied)

Triangle inequality in a complete graph: for vertices x, y, z and edge weights w,

$$w(x, z) \leq w(x, y) + w(y, z)$$

$\Rightarrow$ same intuition; adding an intermediate step is never a shortcut
$\Rightarrow$ automatically holds for Euclidean graphs

TSP with Triangle Inequality (TSPTI)

input: a complete undirected graph $G = (V, E)$ where each edge has weight $w(e) \geq 0$; and for any $x, y, z \in V$,
$w(x, z) \leq w(x, y) + w(y, z)$
output: (same)

- **renegotiating** TSP
- different problem; *NP*-completeness and *APX*-completeness proofs may not apply
- less-general problem
- probably still relevant to practical applications of TSP

TSPTI Approximation Algorithm Idea

- need a structure that can lower-bound an optimal cycle $H^\star$ and upper-bound our approximate cycle H
- *minimum spanning tree* features
  - minimizes weight of chosen edges
  - connects all vertices
  - can be computed fast
- but an MST is not a Hamiltonian cycle; MST is acyclic, for one thing
- *Euler tour:* cycle around a tree; preorder, inorder, postorder
- build an MST; perform preorder traversal; treat that vertex order as Hamiltonian cycle

Approximate TSPTI Pseudocode

```
function APPROX-TSPTI(G = (V, E), w)
    T = PRIM − MST(G, w)
    H = empty sequence of vertices
    for vertex v in preorder traversal of tree T do
        H.ADDBACK(v)
    end for
    return H
end function
```

Analysis: Prim's algorithm takes $O(m + n \log n)$ (w/ Fibonacci heap), traversal takes $O(m + n)$, total $O(m + n \log n)$ time

TSPTI Performance

Lemma: APPROX-TSPTI is a 2-approximation algorithm

Proof Sketch:

- let $H^\star$ be an optimal Hamiltonian cycle for G
- (1) every spanning tree is one edge short of a cycle; and weights are nonnegative; so the weight of our tree T obeys $w(T) \leq w(H^\star)$
- (2) a *full tour* W is the sequence of vertices in both a preorder and postorder tour, and has weight $w(W) = 2w(T)$
- (3) combining (1) and (2), $w(W) \leq 2w(H^\star)$
- (4) our H is like W with some vertices removed, so $w(H) \leq w(W)$
- combining (3) and (4),

$$w(H) \leq w(W) \leq 2w(H^\star)$$

Summary

There is a 2-approximation algorithm for vertex cover that takes $O(m+n)$ time.

There is a 2-approximation algorithm for TSP with the triangle inequality that takes $O(m+n\log n)$ time.